

Fuzzy Password-Authenticated Key Exchange

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Abstract. Consider key agreement by two parties who start out knowing a common secret (which we refer to as “pass-string”, a generalization of “password”), but face two complications: (1) the pass-string may come from a low-entropy distribution, and (2) the two parties’ copies of the pass-string may have some noise, and thus not match exactly. We provide the first efficient and general solutions to this problem that enable, for example, key agreement based on commonly used biometrics such as iris scans.

The problem of key agreement with each of these complications individually has been well studied in literature. Key agreement from low-entropy shared pass-strings is achieved by *password-authenticated key exchange* (PAKE), and key agreement from noisy but high-entropy shared pass-strings is achieved by information-reconciliation protocols as long as the two secrets are “close enough.” However, the problem of key agreement from noisy low-entropy pass-strings has never been studied.

We introduce (universally composable) *fuzzy password-authenticated key exchange* (fPAKE), which solves exactly this problem. fPAKE does not have any entropy requirements for the pass-strings, and enables secure key agreement as long as the two pass-strings are “close” for some notion of closeness. We also give two constructions. The first construction achieves our fPAKE definition for any (efficiently computable) notion of closeness, including those that could not be handled before even in the high-entropy setting. It uses Yao’s garbled circuits in a way that is only two times more costly than their use against semi-honest adversaries, but that guarantees security against malicious adversaries. The second construction is more efficient, but achieves our fPAKE definition only for pass-strings with low Hamming distance. It builds on very simple primitives: robust secret sharing and PAKE.

Keywords: Authenticated Key Exchange, PAKE, Hamming Distance, Error Correcting Codes, Yao’s Garbled Circuits

1 Introduction

Consider key agreement by two parties who start out knowing a common secret (which we refer to as “pass-string”, a generalization of “password”). These parties may face several complications: (1) the pass-string may come from a non-uniform, low-entropy distribution, and (2) the two parties’ copies of the pass-string may have some noise, and thus not match exactly. The use of such pass-strings for security has been extensively studied; examples include biometrics and other human-generated data [Dau04, ZH93, BS00, EHMS00, MG09, MRW02, KR08], physically unclonable functions (PUFs) [PRTG02, GCvD02, TSS⁺06, SD07, YD10], noisy channels [Wyn75], quantum information [BBR88], and sensor readings of a common environment [HHW⁺17, HCS⁺18].

The Noiseless Case. When the starting secret is not noisy (i.e., the same for both parties), existing approaches work quite well. The case of low-entropy secrets is covered by *password-authenticated key exchange* (PAKE), in a long line of work the first formal models for which were introduced by Bellare *et al.* [BPR00] and Boyko *et al.* [BMP00]. A PAKE protocol allows two parties to agree on a shared high-entropy key if and only if they hold the same short password. Even though the password may have low entropy, PAKE ensures that off-line dictionary attacks are impossible. Roughly speaking, an adversary has to participate in one on-line interaction for every attempted guess at the password. Because key agreement is not usually the final goal, PAKE protocols need to be securely composable with whatever protocols (such as authenticated encryption) use the output key. This composability has been achieved by universally composable (UC) PAKE defined by Canetti *et al.* [CHK⁺05] and implemented in several follow-up works.

In the case of high-entropy secrets, off-line dictionary attacks are not a concern, which enables more efficient protocols. If the adversary is passive, randomness extractors [NZ93] do the job. The case of active adversaries is covered by the literature on so-called robust extractors defined by Boyen *et al.* [BDK⁺05] and, more generally, by many papers on privacy amplification protocols secure against active adversaries, starting with the work of Maurer [Mau97]. Composability for these protocols is less studied; in particular, most protocols leak information about the pass-string itself, in which case reusing the pass-string over multiple protocol executions may present problems [Boy04] (with the exception of [CFP⁺16]).

The Noisy Case. When the pass-string is noisy (i.e., the two parties have slightly different versions of it), this problem has been studied only for the case of high-entropy pass-strings. A long series of works on information-reconciliation protocols (started by Bennett *et al.* [BBR88]) and their one-message variants called fuzzy extractors (defined by Dodis *et al.* [DORS08], further enhanced for active security starting by Renner *et al.* [RW04]) achieves key agreement when the pass-string has a lot of entropy and not too much noise. Unfortunately, these approaches do not extend to the low-entropy setting and are not designed to prevent off-line dictionary attacks.

Constructions for the noisy case depend on the specific noise model. The case of binary Hamming distance — when the n pass-string bits held by the two parties are the

same at all but δ locations — is the best studied. Most existing constructions require, at a minimum, that the pass-string should have at least δ bits of entropy. This requirement rules out using most kinds of biometric data as the pass-string— for example, estimates of entropy for iris scans (transformed into binary strings via wavelet transforms and projections) are considerably lower than the amount of errors that need to be tolerated [BH09, Section 5]. Even the PAKE-based construction of Boyen *et al.* [BDK⁺05] suffers from the same problem.

One notable exception is the construction of Canetti *et al.* [CFP⁺16], which does not have such a requirement, but places other stringent limitations on the probability distribution of pass-strings. In particular, because it is a one-message protocol, it cannot be secure against off-line dictionary attacks.

1.1 Our Contributions

We provide definitions and constant-round protocols for key agreement from noisy pass-strings that:

- Resist off-line dictionary attacks and thus can handle low-entropy pass-strings,
- Can handle a variety of noise types and have high error-tolerance, and
- Have well specified composition properties via the UC framework [Can01].

Instead of imposing entropy requirements or other requirements on the distribution of pass-strings, our protocols are secure as long as the adversary cannot guess a pass-string value that is sufficiently close. There is no requirement, for example, that the amount of pass-string entropy is greater than the number of errors; in fact, one of our protocols is suitable for iris scans. Moreover, our protocols prevent off-line attacks, so each adversarial attempt to get close to the correct pass-string requires an on-line interaction by the adversary. Thus, for example, our protocols can be meaningfully run with pass-strings whose entropy is only 30 bits—something not possible with any prior protocols for the noisy case.

New Models. Our security model is in the Universal Composability (UC) Framework of Canetti [Can01]. The advantage of this framework is that it comes with a composability theorem that ensures that the protocol stays secure even when run in arbitrary environments, including arbitrary parallel executions. Composability is particularly important for key agreement protocols, because key agreement is rarely the ultimate goal. The agreed-upon key is typically used for some subsequent protocol—for example, to instantiate a secure channel. Further, this framework allows us to give a definition that is indifferent to how the initial pass-strings are generated. We have no entropy requirements or constraints on the pass-string distribution; rather, security is guaranteed as long as the adversary’s input to the protocol is far enough from the correct pass-string.

As a starting point, we use the definition of UC security for PAKE from Canetti *et al.* [CHK⁺05]. The PAKE ideal functionality is defined as follows: the secret pass-strings (called “passwords” in PAKE) of the two parties are the inputs to the functionality, and two random keys, which are equal if and only if the two inputs are equal, are the outputs. The

main change we make to PAKE is enhancing the functionality to give equal keys even if the two inputs are not equal, as long as they are close enough. We also relax the security requirement to allow one party to find out some information about the other party’s input—perhaps even the entire input—if the two inputs are close. This relaxation makes sense in our application: if the two parties are honest, then the differences between their inputs are a problem rather than a feature, and we would not mind if the inputs were in fact the same. The benefit of this relaxation is that it permits us to construct more efficient protocols. (We also make a few other minor changes which will be described in Section 2.) We call our new UC functionality “Fuzzy Password-Authenticated Key Exchange” or fPAKE.

New Protocols. The only prior PAKE-based protocol for the noisy setting by Boyen *et al.* [BDK⁺05], although more efficient than ours, does not satisfy our goal. In particular, it is not composable, because it reveals information about the secret pass-strings (we demonstrate this formally in Appendix H). Because some information about the pass-strings is unconditionally revealed, high-entropy pass-strings are required. Thus, in order to realize our definition for arbitrary low-entropy pass-strings, we need to construct new protocols.

Realizing our fPAKE definition is easy using general two-party computation techniques for protocols with malicious adversaries and without authenticated channels [BCL⁺05]. However, we develop protocols that are considerably more efficient: our definitional relaxation allows us to build protocols that achieve security against malicious adversaries but cost just a little more than the generic two-party computation protocols that achieve security only against honest-but-curious adversaries (i.e., adversaries who do not deviate from the protocol, but merely try to infer information they are not supposed to know).

Our first construction uses Yao’s garbled circuits [Yao86, BHR12] and oblivious transfer (see Chou and Orlandi [CO15] and references therein). The use of these techniques is standard in two-party computation. However, by themselves they give protocols secure only against honest-but-curious adversaries. In order to prevent malicious behavior of the players, one usually applies the cut-and-choose technique [LP11], which is quite costly: to achieve an error probability of $2^{-\lambda}$, the number of circuits that need to be garbled increases by a factor of λ , and the number of oblivious transfers that need to be performed increases by a factor of $\lambda/2$. We show that for our special case, to achieve malicious security, it suffices to repeat the honest-but-curious protocol twice (once in each direction), incurring only a factor of 2 overhead over the semi-honest case.¹ Mohassel *et al.* [MF06] and Huang *et al.* [HKE12] suggest a similar technique (known as “dual execution”), but at the cost of leaking a bit of the adversary’s choice to the adversary. In contrast, our construction leaks nothing to the adversary at all (as long as the pass-strings are not close). This construction works regardless of what it means for the two inputs to be “close,” as long as the question of closeness can be evaluated by an efficient circuit.

¹ Gasti *et al.* [GSY⁺16] similarly use Yao’s garbled circuits for continuous biometric user authentication on a smartphone. Our approach can eliminate the third party in their application, at the cost of requiring two garbled circuits instead of one. As far as we know, ours is the first use of garbled circuits in the two-party fully malicious setting without calling on an expensive transformation.

Our second construction is for the Hamming case: the two n -character pass-strings have low Hamming distance if not too many characters of one party’s pass-string are different from the corresponding characters of the other’s pass-string. The two parties execute a PAKE protocol for each position in the string, obtaining n values each that agree or disagree depending on whether the characters of the pass-string agree or disagree in the corresponding positions. It is important that at this stage, agreement or disagreement at individual positions remains unknown to everyone; we therefore make use of a special variant of PAKE which we call *implicit-only PAKE* (we give a formal UC security definition of implicit-only PAKE and show that it is realized by the PAKE protocol of Bellare and Merritt [BM92] and Abdalla *et al.* [ACCP08]). This first step upgrades Hamming distance over a potentially small alphabet to Hamming distance over an exponentially large alphabet. We then secret-share the ultimate output key into n shares using a robust secret sharing scheme, and encrypt each share using the output of the corresponding PAKE protocol.

The second construction is more efficient than the first in the number of rounds, communication, and computation. However, it works only for Hamming distance. Moreover, it has an intrinsic gap between functionality and security: if the honest parties need to be within distance δ to agree, then the adversary may break security by guessing a secret within distance 2δ . See Figure 11 for a comparison between the two constructions.

The advantages of our protocols are similar to the advantages of universally composable PAKE: They provide composability, protection against off-line attacks, the ability to use low-entropy secret inputs, and handle any distribution of those inputs. And, of course, because we construct *fuzzy* PAKE, our protocols can handle noisy inputs—including many types of noisy inputs that could not be handled before. Our first protocol can handle any type of noise as long as the notion of “closeness” can be efficiently computed, whereas most prior work was for Hamming distance only. However, these advantages come at the price of efficiency. Our protocols require 2–5 rounds of interaction, as opposed to many single-message protocols in the literature [DKK⁺12, CFP⁺16, WCD⁺17]. They are also more computationally demanding than most existing protocols for the noisy case, requiring one public-key operation per input character. We emphasize, however, that our protocols are much less computationally demanding than the protocols based on general two-party computation, as already discussed above, or general-purpose obfuscation, as discussed in [BCKP14, Section 4.3.4].

2 Security Model

We now present a security definition for fuzzy password-authenticated key exchange (fPAKE). We adapt the definition of PAKE from Canetti *et al.* [CHK⁺05] to work for pass-strings (a generalization of “passwords”) that are similar, but not necessarily equal. Our definition uses measures of the distance $d(\text{pw}, \text{pw}')$ between pass-strings $\text{pw}, \text{pw}' \in \mathbb{F}_p^n$. In Section 3.3 and Section 4, Hamming distance is used, but in the generic construction of Section 3, any (efficiently computable) other notion of distance can be used instead. We say that pw and pw' are “similar enough” if $d(\text{pw}, \text{pw}') \leq \delta$ for a distance notion d and a threshold δ that is hard-coded into the functionality.

Parties first engage our functionality (described in Figure 1) by making `NewSession` queries, which include their pass-strings. Once both parties have made `NewSession` queries, the simulator can make `NewKey` queries on behalf of the parties, prompting the functionality to release an appropriate session key to the party in question. In an execution in which the adversary does not meddle, both session keys will be random: they will match if the pass-strings are “similar enough”, and be independent otherwise.

Modeling Adversarial Capabilities To model the possibility of dictionary attacks, the functionality allows the adversary to make one pass-string guess against each player (\mathcal{P}_0 and \mathcal{P}_1). In the real world, if the adversary succeeds in guessing (a pass-string similar enough to) party \mathcal{P}_i ’s pass-string, it can often choose (or at least bias) the session key computed by \mathcal{P}_i . To model this, the functionality then allows the adversary to set the session key for \mathcal{P}_i .

As usual in security notions for key exchange, the adversary also sets the session keys for corrupted players. In the definition of Canetti *et al.* [CHK⁺05], the adversary additionally sets \mathcal{P}_i ’s key if \mathcal{P}_{1-i} is corrupted. However, contrary to the original definition, we do not allow the adversary to set \mathcal{P}_i ’s key if \mathcal{P}_{1-i} is corrupted but did not guess \mathcal{P}_i ’s pass-string. We make this change in order to protect an honest \mathcal{P}_i from, for instance, revealing sensitive information to an adversary who did not successfully guess her pass-string, but did corrupt her partner.

Roles There are two categories of fPAKE protocols: *symmetric protocols* in which the two parties execute the same code, and *asymmetric protocols* in which the two parties execute different code. Frequently in asymmetric protocols, one party can be seen as the “sender” who initiates the protocol, and the other can be seen as the “receiver” who responds.²

In our ideal functionality, each party includes a `role` tag in her `NewSession` query; one party should identify herself as the sender (denoted as `role = sender`), while the other should identify herself as the receiver (`role = receiver`). The functionality simply forwards these role tags to the simulator; the roles do not affect any of the functionality’s decisions.

In the case of symmetric protocols, the `role` tags are unnecessary, since a sender and a receiver execute the same code. In the case of asymmetric protocols, the simulator needs the `role` tags in order to determine which code to execute. It might look strange that the functionality ignores these `role` tags once it forwards them to the simulator; it might seem that, in the case of an asymmetric protocol, the functionality should only proceed if one of the roles provided is `sender` and the other `receiver`. However, in such a situation, the simulator can trigger the desired behavior — an abort — simply by never issuing a `NewKey` query.³

² To reflect the fact that, even in symmetric protocols, one party likely requests that the other engage in key exchange with her, such a request message can be pre-pended to any symmetric protocol.

³ An asymmetric protocol where the parties do *not* abort when both are executing the same role’s code (but the resulting keys are not distributed as they should be) cannot securely instantiate our functionality.

The functionality fPAKE is parameterized by a security parameter λ and tolerances $\delta \leq \gamma$. It interacts with an adversary \mathcal{S} and two parties \mathcal{P}_0 and \mathcal{P}_1 via the following queries:

- **Upon receiving a query ($\text{NewSession}, \text{sid}, \text{pw}_i, \text{role}$) from party \mathcal{P}_i ,** where pw_i is a password and $\text{role} = \text{sender}$ implies that \mathcal{P}_i wishes to initiate a key exchange, while $\text{role} = \text{receiver}$ implies that \mathcal{P}_i wishes to respond:
 - Send ($\text{NewSession}, \text{sid}, \mathcal{P}_i, \text{role}$) to \mathcal{S} ;
 - If one of the following is true, record $(\mathcal{P}_i, \text{pw}_i)$ and mark this record **fresh**:
 - * This is the first **NewSession** query
 - * This is the second **NewSession** query and there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$
- **Upon receiving a query ($\text{TestPwd}, \text{sid}, \mathcal{P}_i, \text{pw}'_i$) from the adversary \mathcal{S} :**
 If there is a **fresh** record $(\mathcal{P}_i, \text{pw}_i)$, then set $d \leftarrow d(\text{pw}_i, \text{pw}'_i)$ and do:
 - If $d \leq \delta$, mark the record **compromised** and reply to \mathcal{S} with “correct guess”;
 - If $d > \delta$, mark the record **interrupted** and reply to \mathcal{S} with “wrong guess”.
- **Upon receiving a query ($\text{NewKey}, \text{sid}, \mathcal{P}_i, \text{sk}$) from the adversary \mathcal{S} :**
 If there is no record of the form $(\mathcal{P}_i, \text{pw}_i)$, or if this is not the first **NewKey** query for \mathcal{P}_i , then ignore this query. Otherwise:
 - If at least one of the following is true, then output (sid, sk) to player \mathcal{P}_i :
 - * The record is **compromised**
 - * \mathcal{P}_i is corrupted
 - * The record is **fresh**, \mathcal{P}_{1-i} is corrupted, and there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ with $d(\text{pw}_i, \text{pw}_{1-i}) \leq \delta$
 - If this record is **fresh**, both parties are honest, there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ with $d(\text{pw}_i, \text{pw}_{1-i}) \leq \delta$, a key sk' was sent to \mathcal{P}_{1-i} , and $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ was **fresh** at the time, then output (sid, sk') to \mathcal{P}_i ;
 - In any other case, pick a new random key sk' of length λ and send (sid, sk') to \mathcal{P}_i .
 - Mark the record $(\mathcal{P}_i, \text{pw}_i)$ as **completed**.

Fig. 1. Ideal Functionality fPAKE

Notes Another minor change we make is considering only two parties — \mathcal{P}_0 and \mathcal{P}_1 — in the functionality, instead of considering arbitrarily many parties and enforcing that only two of them engage the functionality. This is because universal composability takes care of ensuring that a two-party functionality remains secure in a multi-party world.

As in the definition of Canetti *et al.* [CHK⁺05], we consider only static corruptions in the standard corruption model of Canetti [Can01]. Also as in their definition, we chose not to provide the players with confirmation that key agreement was successful. The players might obtain such confirmation from subsequent use of the key.

Leakage By default, in the fPAKE functionality the **TestPwd** interface provides the adversary with one bit of information — whether the pass-string guess was correct or not. This definition can be strengthened by providing the adversary with no information at all, as in implicit-only PAKE ($\mathcal{F}_{\text{iPAKE}}$, Figure 8), or weakened by providing the adversary with extra information when the adversary’s guess is close enough.

To capture the diversity of possibilities, we introduce a more general **TestPwd** interface, described in Figure 2. It includes three leakage functions that we will instantiate in different ways below— L_c if the guess is close enough to succeed, L_f if it is too far. Moreover, a third leakage function— L_m for medium distance—allows the adversary to get some information even if the adversary’s guess is only somewhat close (closer than

- **Upon receiving a query** ($\text{TestPw}, \text{sid}, \mathcal{P}_i, \text{pw}'_i$) **from the adversary** \mathcal{S} :
 If there is a **fresh** record $(\mathcal{P}_i, \text{pw}_i)$, then set $d \leftarrow d(\text{pw}_i, \text{pw}'_i)$ and do:
 - If $d \leq \delta$, mark the record **compromised** and reply to \mathcal{S} with $L_c(\text{pw}_i, \text{pw}'_i)$;
 - If $\delta < d \leq \gamma$, mark the record **compromised** and reply to \mathcal{S} with $L_m(\text{pw}_i, \text{pw}'_i)$;
 - If $\gamma < d$, mark the record **interrupted** and reply to \mathcal{S} with $L_f(\text{pw}_i, \text{pw}'_i)$.

Fig. 2. A Modified TestPw Interface to Allow for Different Leakage

some parameter $\gamma \geq \delta$), but not close enough for successful key agreement. We thus decouple the distance needed for functionality from the (possibly larger) distance needed to guarantee security; the smaller the gap between these two distances, the better, of course.

Below, we list the specific leakage functions L_c , L_m and L_f that we consider in this work, in order of decreasing strength (or increasing leakage):

1. The strongest option is to provide no feedback at all to the adversary. We define fPAKE^N to be the functionality described in Figure 1, except that TestPw is from Figure 2 with

$$L_c^N(\text{pw}_i, \text{pw}'_i) = L_m^N(\text{pw}_i, \text{pw}'_i) = L_f^N(\text{pw}_i, \text{pw}'_i) = \perp.$$

2. The basic functionality fPAKE , described in Figure 1, leaks the correctness of the adversary's guess. That is, in the language of Figure 2,

$$\begin{aligned} L_c(\text{pw}_i, \text{pw}'_i) &= \text{"correct guess"}, \\ \text{and } L_m(\text{pw}_i, \text{pw}'_i) &= L_f(\text{pw}_i, \text{pw}'_i) = \text{"wrong guess"}. \end{aligned}$$

The classical PAKE functionality from [CHK⁺05] has such a leakage.

3. Assume the two pass-strings are strings of length n over some finite alphabet, with the j th character of the string pw denoted by $\text{pw}[j]$. We define fPAKE^M to be the functionality described in Figure 1, except that TestPw is from Figure 2, with L_c and L_m that leak the indices at which the guessed pass-string differs from the actual one when the guess is close enough (we will call this leakage the *mask* of the pass-strings). That is,

$$\begin{aligned} L_c^M(\text{pw}_i, \text{pw}'_i) &= (\{j \text{ s.t. } \text{pw}_i[j] = \text{pw}'_i[j]\}, \text{"correct guess"}), \\ L_m^M(\text{pw}_i, \text{pw}'_i) &= (\{j \text{ s.t. } \text{pw}_i[j] \neq \text{pw}'_i[j]\}, \text{"wrong guess"}) \\ \text{and } L_f^M(\text{pw}_i, \text{pw}'_i) &= \text{"wrong guess"}. \end{aligned}$$

4. The weakest definition — or the strongest leakage — reveals the entire actual pass-string to the adversary if the pass-string guess is close enough. We define fPAKE^P to be the functionality described in Figure 1, except that TestPw is from Figure 2, with

$$L_c^P(\text{pw}_i, \text{pw}'_i) = L_m^P(\text{pw}_i, \text{pw}'_i) = \text{pw}_i \quad \text{and} \quad L_f^P(\text{pw}_i, \text{pw}'_i) = \text{"wrong guess"}.$$

Here, L_c^P and L_m^P do not need to include “correct guess” and “wrong guess”, respectively, because this is information that can be easily derived from pw_i itself.

The first two functionalities are the strongest, but there are no known constructions that realize them, other than through generic two-party computation secure against malicious adversaries, which is an inefficient solution. The last two functionalities, though weaker, still provide meaningful security, especially when $\gamma = \delta$. Intuitively, this is because strong leakage only occurs when an adversary guesses a “close” pass-string, which enables him to authenticate as though he knows the real pass-string anyway.

In Section 3, we present a construction satisfying fPAKE^P for any efficiently computable notion of distance, with $\gamma = \delta$ (which is the best possible). We present a construction for Hamming distance satisfying fPAKE^M in Section 4, with $\gamma = 2\delta$.

3 General Construction Using Garbled Circuits

In this section, we describe a protocol realizing fPAKE^P that uses Yao’s garbled circuits [Yao86]. We briefly introduce this primitive in Sec. 3.1 and refer to Yakoubov [Yak17] for a more thorough introduction.

The Yao’s garbled circuit-based fPAKE construction has two advantages:

1. It is more flexible than other approaches; any notion of distance that can be efficiently computed by a circuit can be used. In Section 3.3, we describe a suitable circuit for Hamming distance. The total size of this circuit is $O(n)$, where n is the length of the pass-strings used. Edit distance is slightly less efficient, and uses a circuit whose total size is $O(n^2)$.
2. There is no gap between the distances required for functionality and security — that is, there is no leakage about the pass-strings used unless they are similar enough to agree on a key. In other words, $\delta = \gamma$.

Informally, the construction involves the garbled evaluation of a circuit that takes in two pass-strings as input, and computes whether their distance is less than δ . Because Yao’s garbled circuits are only secure against semi-honest garblers, we cannot simply have one party do the garbling and the other party do the evaluation. A malicious garbler could provide a garbling of the wrong function — maybe even a constant function — which would result in successful key agreement even if the two pass-strings are very different. However, as suggested by Mohassel *et al.* [MF06] and Huang *et al.* [HKE12], since a malicious evaluator (unlike a malicious garbler) cannot compromise the computation, by performing the protocol twice with each party playing each role once, we can protect against malicious behavior. They call this the *dual execution* protocol.

The dual execution protocol has the downside of allowing the adversary to specify and receive a single additional bit of leakage. It is important to note that because of this, dual execution cannot directly be used to instantiate fPAKE , because a single bit of leakage can be too much when the entropy of the pass-strings is low to begin with — a few adversarial attempts will uncover the entire pass-string. Our construction is as efficient as that of Mohassel *et al.* and Huang *et al.*, while guaranteeing no leakage to a malicious adversary in the case that the pass-strings used are not close. We describe how we achieve this in Section 3.1.3.

Due to the symmetric layout of our construction, we skip all `role` tags in this section.

3.1 Building Blocks

In Section 3.1.1, we briefly review oblivious transfer. In Section 3.1.2, we review Yao’s Garbled Circuits. In Section 3.1.3, we describe in more detail our take on the dual execution protocol, and how we avoid leakage to the adversary when the pass-strings used are dissimilar.

3.1.1 Oblivious Transfer (OT) Informally, 1-out-of-2 Oblivious Transfer (see Chou and Orlandi [CO15] and citations therein) enables one party (the sender) to transfer exactly one of two secrets to another party (the receiver). The receiver chooses (by index 0 or 1) which secret she wants. The security of the OT protocol guarantees that the sender does not learn this choice bit, and the receiver does not learn anything about the other secret.

3.1.2 Yao’s Garbled Circuits (YGC) In this section, we give a brief introduction to Yao’s garbled circuits [Yao86]. We refer to Yakoubov [Yak17] for a more detailed description, as well as a summary of some of the Yao’s garbled circuits optimizations [BMR90, KS08, PSSW09, KMR14, ZRE15, BMR16]. Informally, Yao’s garbled circuits are an asymmetric secure two-party computation scheme. They enable two parties with sensitive inputs (in our case, pass-strings) to compute a joint function of their inputs (in our case, an augmented version of similarity) without revealing any additional information about their inputs. One party “garbles” the function they wish to evaluate, and the other evaluates it in its garbled form.

Below, we summarize the garbling scheme formalization of Bellare *et al.* [BHR12], which is a generalization of YGC.

Functionality. A garbling scheme \mathcal{G} consists of a tuple of four polynomial-time algorithms (Gb, En, Ev, De):

1. $\text{Gb}(1^\lambda, f) \rightarrow (F, e, d)$. The garbling algorithm **Gb** takes in the security parameter λ and a circuit f , and returns a garbled circuit F , encoding information e , and decoding information d .
2. $\text{En}(e, x) \rightarrow X$. The encoding algorithm **En** takes in the encoding information e and an input x , and returns a garbled input X .
3. $\text{Ev}(F, X) \rightarrow Y$. The evaluation algorithm **Ev** takes in the garbled circuit F and the garbled input X , and returns a garbled output Y .
4. $\text{De}(d, Y) \rightarrow y$. The decoding algorithm **De** takes in the decoding information d and the garbled output Y , and returns the plaintext output y .

A garbling scheme $\mathcal{G} = (\text{Gb}, \text{En}, \text{Ev}, \text{De})$ is *projective* if the encoding information e consists of $2n$ wire labels (each of which is essentially a random string), where n is the number of input bits. Two wire labels are associated with each bit of the input; one wire label corresponds to the event of that bit being 0, and the other corresponds to the event of that bit being 1. The garbled input includes only the wire labels corresponding to the

actual values of the input bits. In projective schemes, in order to give the evaluator the garbled input she needs for evaluation, the garbler can send her all of the wire labels corresponding to the garbler’s input. The evaluator can then use OT to retrieve the wire labels corresponding to her own input.

Similarly, we call a garbling scheme *output-projective* if decoding information d consists of two labels for each output bit, one corresponding to each possible value of that bit. The garbling schemes used in this paper are both projective and output-projective.

Correctness. Informally, a garbling scheme $(\text{Gb}, \text{En}, \text{Ev}, \text{De})$ is *correct* if it always holds that $\text{De}(d, \text{Ev}(F, \text{En}(e, x))) = f(x)$.

Security. Bellare *et al.* [BHR12] describe three security notions for garbling schemes: *obliviousness*, *privacy* and *authenticity*. Informally, a garbling scheme $\mathcal{G} = (\text{Gb}, \text{En}, \text{Ev}, \text{De})$ is *oblivious* if a garbled function F and a garbled input X do not reveal anything about the input x . It is *private* if additionally knowing the decoding information d reveals the output y , but does not reveal anything more about the input x . It is *authentic* if an adversary, given F and X , cannot find a garbled output $Y' \neq \text{Ev}(F, X)$ which decodes without error.

In Appendix B, we define a new property of output-projective garbling schemes called *garbled output randomness*. Informally, it states that even given one of the output labels, the other should be indistinguishable from random.

3.1.3 Malicious Security: A New Take on Dual Execution with Privacy-Correctness Tradeoffs While Yao’s garbled circuits are naturally secure against a malicious evaluator, they have the drawback of being insecure against a malicious garbler. A garbler can “mis-garble” the function, either replacing it with a different function entirely or causing an error to occur in an informative way (this is known as “selective failure”).

Typically, malicious security is introduced to Yao’s garbled circuits by using the cut-and-choose transformation [LP15, Lin13, HKE13]. To achieve a $2^{-\lambda}$ probability of cheating without detection, the parties need to exchange λ garbled circuits [Lin13].⁴ Some of the garbled circuits are “checked”, and the rest of them are evaluated, their outputs checked against one another for consistency. Because of the factor of λ computational overhead, though, cut-and-choose is expensive, and too heavy a tool for fPAKE. Other, more efficient transformations such as LEGO [NO09] and authenticated garbling [WRK17] exist as well, but those rely heavily on pre-processing, which cannot be used in fPAKE since it requires advance interaction between the parties.

Mohassel *et al.* [MF06] and Huang *et al.* [HKE12] suggest an efficient transformation known as “dual execution”: each party plays each role (garbler and evaluator) once, and then the two perform a comparison step on their outputs in a secure fashion. Dual execution incurs only a factor of 2 overhead over semi-honest garbled circuits. However,

⁴ There are techniques [LR14] that improve this number in the amortized case when many computations are done — however, this does not fit our setting.

it does not achieve fully malicious security. It guarantees correctness, but reduces the privacy guarantee by allowing a malicious garbler to learn one bit of information of her choice. Specifically, if a malicious garbler garbles a wrong circuit, she can use the comparison step to learn one bit about the output of this wrong circuit on the other party’s input. This one extra bit of information could be crucially important, violating the privacy of the evaluator’s input in a significant way.

We introduce a tradeoff between correctness and privacy for boolean functions. For one of the two possible outputs (without loss of generality, ‘0’), we restore full privacy at the cost of correctness. The new privacy guarantee is that if the correct output is ‘0’, then a malicious adversary cannot learn anything beyond this output, but if the correct output is ‘1’, then she can learn a single bit of her choice. The new correctness guarantee is that a malicious adversary can cause the computation that should output ‘1’ to output ‘0’ instead, but not the other way around. Our privacy–correctness tradeoff is summarized in Figure 3.

[MF06] [HKE12]	Correct Output	Computed Output	Privacy
	1	1 OR ‘cheating’	1-bit leakage
	0	0 OR ‘cheating’	1-bit leakage
Our Protocol	Correct Output	Computed Output	Privacy
	1	1 OR 0	1-bit leakage
	0	0	full privacy

Fig. 3. The Privacy-Correctness Tradeoff of Dual Execution Protocols for Boolean Functions

The main idea of dual execution is to have the two parties independently evaluate one another’s circuits, learn the output values, and compare the output labels using a secure comparison protocol. This comparison step is simply a check for malicious behavior; if comparison fails, then honest party \mathcal{P}_i learns that \mathcal{P}_{1-i} cheated. If the comparison step succeeded, \mathcal{P}_{1-i} might still have cheated — and gleaned an extra bit of information — but \mathcal{P}_i is assured that she has the correct output.

In our construction, however, the parties need not learn the output values before the comparison. Instead, the parties can compare output labels *assuming* an output of ‘1’, and if the comparison fails, the output is determined to be ‘0’. More formally, let $d_0[0]$, $d_0[1]$ be the two output labels corresponding to \mathcal{P}_0 ’s garbled circuit, and $d_1[0]$, $d_1[1]$ be the two output labels corresponding to \mathcal{P}_1 ’s circuit. Let $Y_1 \in [d_1[0], d_1[1]]$ be the output label learned by \mathcal{P}_0 as a result of evaluation, and $Y_0 \in [d_0[0], d_0[1]]$ be the label learned by \mathcal{P}_1 . The two parties securely compare $(d_0[1], Y_1)$ to $(Y_0, d_1[1])$; if the comparison succeeds, the output is “1”.

Whereas in dual execution the comparison step is just a sanity check, here it determines the actual computation output. If the correct output is ‘1’, a cheating \mathcal{P}_{1-i} can still learn one bit of information by mis-garbling her circuit; depending on the output of the mis-garbled circuit, the comparison step will either succeed or fail. If the comparison fails, \mathcal{P}_i will accept an incorrect output of ‘0’, and never be aware that \mathcal{P}_{1-i} cheated. If

the correct output is ‘0’, however, there is nothing \mathcal{P}_{1-i} can do to cause the comparison step to succeed, since in order to do this, she would need to use the second output label $d_i[1]$ as an input. Since the true output was ‘0’, and thus $Y_i = d_i[0]$, by the garbled output randomness property of the garbling scheme, \mathcal{P}_{1-i} can’t even distinguish $d_i[1]$ from random.

Our privacy–correctness tradeoff is perfect for **fPAKE**. If the parties’ inputs are similar, learning a bit of information about each other’s inputs is not problematic, since arguably the small amount of noise in the inputs is a bug, not a feature. If the parties’ inputs are not similar, however, we are guaranteed to have no leakage at all. We pay for the lack of leakage by allowing a malicious party to force an authentication failure even when authentication should succeed, which either party can do anyway simply by providing an incorrect input.

In Section 3.2.2, we describe our Yao’s garbled circuit-based **fPAKE** protocol in detail. Note that in this protocol, we omit the final comparison step; instead, we use the output labels $((d_0[1], Y_1)$ and $(Y_0, d_1[1]))$ to compute the agreed-upon key directly (via XOR).

3.2 Construction

Building a **fPAKE** from YGC and OT is not straightforward, since all constructions of OT assume authenticated channels, and **fPAKE** (or **PAKE**) is designed with unauthenticated channels in mind. We therefore follow the framework of Canetti *et al.* [CDVW12], who build a UC secure **PAKE** protocol using OT. We first build our protocol assuming authenticated channels, and then apply the generic transformation of Barak *et al.* [BCL⁺05] to adapt it to the unauthenticated channel setting. More formally, we proceed in three steps:

1. First, in Section 3.2.1, we define a randomized fuzzy equality-testing functionality \mathcal{F}_{RFE} , which is analogous to the randomized equality-testing functionality of Canetti *et al.*
2. In Section 3.2.2, we build a protocol that securely realizes \mathcal{F}_{RFE} in the OT-hybrid model, assuming authenticated channels.
3. In Section 3.2.3, we apply the transformation of Barak *et al.* to our protocol. This results in a protocol that realizes the “split” version of functionality $\mathcal{F}_{\text{RFE}}^P$, which we show to be enough to implement to **fPAKE**^P. Split functionalities, which were introduced by Barak *et al.*, adapt functionalities which assume authenticated channels to an unauthenticated channels setting. The only additional ability an adversary has in a split functionality is the ability to execute the protocol separately with the participating parties.

3.2.1 The Randomized Fuzzy Equality Functionality Figure 4 shows the randomized fuzzy equality functionality $\mathcal{F}_{\text{RFE}}^P$, which is essentially what $\mathcal{F}_{\text{fPAKE}}^P$ would look like assuming authenticated channels. The primary difference between $\mathcal{F}_{\text{RFE}}^P$ and $\mathcal{F}_{\text{fPAKE}}^P$ is that the only pass-string guesses allowed by $\mathcal{F}_{\text{RFE}}^P$ are the ones actually used as protocol inputs; this limits the adversary to guessing by corrupting one of the participating

The functionality \mathcal{F}_{RFE} is parameterized by a security parameter λ and a tolerance δ . It interacts with an adversary \mathcal{S} and two parties \mathcal{P}_0 and \mathcal{P}_1 via the following queries:

- **Upon receiving a query ($\text{NewSession}, \text{sid}, \mathcal{P}_i$) from party $\mathcal{P}_i \in \{\mathcal{P}_0, \mathcal{P}_1\}$:**
 - Send ($\text{NewSession}, \text{sid}, \mathcal{P}_i$) to \mathcal{S} ;
 - If this is the first NewSession query, or if this is the second NewSession query and there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$, then record $(\mathcal{P}_i, \text{pw}_i)$.
- **Upon receiving a query ($\text{TestPwd}, \text{sid}, \mathcal{P}_i$) from the adversary \mathcal{S} , $\mathcal{P}_i \in \{\mathcal{P}_0, \mathcal{P}_1\}$:**
 If records of the form $(\mathcal{P}_0, \text{pw}_0)$ and $(\mathcal{P}_1, \text{pw}_1)$ do not exist, if \mathcal{P}_{1-i} is not corrupted, or this is not the first TestPwd query for \mathcal{P}_i , ignore this query. Otherwise, if $d(\text{pw}_0, \text{pw}_1) \leq \delta$, send pw_i to the adversary \mathcal{S} .
- **Upon receiving a query ($\text{NewKey}, \text{sid}, \mathcal{P}_i, \text{sk}$) from the adversary \mathcal{S} , $\mathcal{P}_i \in \{\mathcal{P}_0, \mathcal{P}_1\}$:**
 If there are no records of the form $(\mathcal{P}_i, \text{pw}_i)$ and $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$, or if this is not the first NewKey query for \mathcal{P}_i , then ignore this query. Otherwise:
 - If at least one of the following is true, then output (sid, sk) to party \mathcal{P}_i .
 - * \mathcal{P}_i is corrupted
 - * \mathcal{P}_{1-i} is corrupted and $d(\text{pw}_0, \text{pw}_1) \leq \delta$
 - If both parties are honest, $d(\text{pw}_0, \text{pw}_1) \leq \delta$, and a key k_{1-i} was sent to \mathcal{P}_{1-i} , then output (sid, k_{1-i}) to \mathcal{P}_i .
 - In any other case, pick a new random key k_i of length λ and send (sid, k_i) to \mathcal{P}_i .

Fig. 4. Ideal Functionality $\mathcal{F}_{\text{RFE}}^P$ for Randomized Fuzzy Equality

parties, not through man in the middle attacks. Like in $\mathcal{F}_{\text{fPAKE}}^P$, if a pass-string guess is “similar enough”, the entire pass-string is leaked. This leakage could be replaced with any other leakage from Section 2; \mathcal{F}_{RFE} would leak the correctness of the guess, $\mathcal{F}_{\text{RFE}}^M$ would leak which characters are the same between the two pass-strings, etc.

Note that, unlike the randomized equality functionality in the work of Canetti *et al.* [CDVW12], $\mathcal{F}_{\text{fPAKE}}^P$ has a TestPwd interface. This is because NewKey does not return the necessary leakage to an honest user. So, an interface enabling the adversary to retrieve additional information is necessary.

3.2.2 A Randomized Fuzzy Equality Protocol In Figure 5 we introduce a protocol Π_{RFE} that securely realizes $\mathcal{F}_{\text{RFE}}^P$ using Yao’s garbled circuits. Garbled circuits are secure against a malicious evaluator, but only a semi-honest garbler; however, we obtain security against malicious adversaries by having each party play each role once, as describe in Section 3.1.3. In more detail, both parties $\mathcal{P}_i \in \{\mathcal{P}_0, \mathcal{P}_1\}$ proceed as follows:

1. \mathcal{P}_i garbles the circuit f that takes in two pass-strings pw_0 and pw_1 , and returns ‘1’ if $d(\text{pw}_0, \text{pw}_1) \leq \delta$ and ‘0’ otherwise. Section 3.3 describes how f can be designed efficiently for Hamming distance. Instead of using the output of f (‘0’ or ‘1’), we will use the garbled output, also referred to as an *output label* in an output-projective garbling scheme. The possible output labels are two random strings — one corresponding to a ‘1’ output (we call this label $k_{i,\text{correct}} = d_i[1]$), and one corresponding to a ‘0’ output (we call this label $k_{i,\text{wrong}} = d_i[0]$).
2. \mathcal{P}_i uses OT to retrieve the input labels from \mathcal{P}_{1-i} ’s garbling that correspond to \mathcal{P}_i ’s pass-string. (Similarly, \mathcal{P}_i uses OT to send \mathcal{P}_{1-i} the input labels from her own garbling that correspond to \mathcal{P}_{1-i} ’s pass-string.)

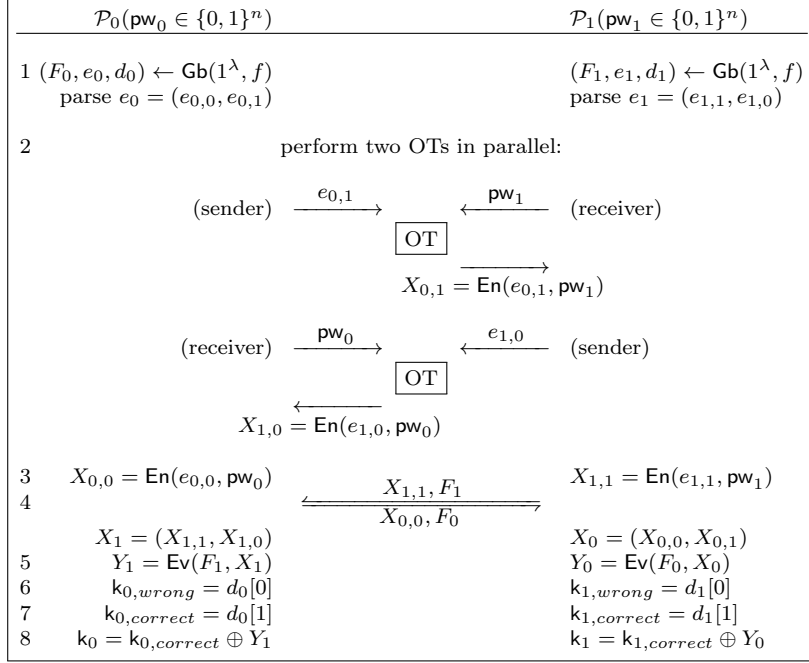


Fig. 5. A Protocol Π_{RFE} Realizing $\mathcal{F}_{\text{RFE}}^P$ using Yao's garbled circuits and an Ideal OT Functionality. If at any point an expected message fails to arrive (or arrives malformed), the parties output a random key. Subscripts are used to indicate who produced the object in question. If a double subscript is present, the second subscript indicates whose data the object is meant for use with. For instance, a double subscript 0, 1 denotes that the object was produced by party \mathcal{P}_0 for use with \mathcal{P}_1 's data; $e_{0,1}$ is encoding information produced by \mathcal{P}_0 to encode \mathcal{P}_1 's pass-string. Note that we abuse notation by encoding inputs to a single circuit separately; the input to \mathcal{P}_0 's circuit corresponding to \mathbf{pw}_0 is encoded by \mathcal{P}_0 locally, and the input corresponding to \mathbf{pw}_1 is encoded via OT. For any projective garbling scheme, this is not a problem.

3. \mathcal{P}_i sends \mathcal{P}_{1-i} her garbled circuit, together with the input labels from her garbling that correspond to her own pass-string. After this step, \mathcal{P}_i should have \mathcal{P}_{1-i} 's garbled circuit and a garbled input consisting of input labels corresponding to the bits of the two pass-strings.
4. \mathcal{P}_i evaluates \mathcal{P}_{1-i} 's garbled circuit, and obtains an output label Y_{1-i} (where $Y_{1-i} \in \{k_{1-i, \text{correct}}, k_{1-i, \text{wrong}}\}$).
5. \mathcal{P}_i outputs $k_i = k_{i, \text{correct}} \oplus Y_{1-i}$.

The natural question to ask is why Π_{RFE} only realizes $\mathcal{F}_{\text{RFE}}^P$, and not a stronger functionality with less leakage. We argue this assuming (without loss of generality) that \mathcal{P}_1 is corrupted. Π_{RFE} cannot realize a functionality that leaks less than the full pass-string \mathbf{pw}_0 to \mathcal{P}_1 if $d(\mathbf{pw}_0, \mathbf{pw}_1) \leq \delta$; intuitively, this is because if \mathcal{P}_1 knows a pass-string \mathbf{pw}_1 such that $d(\mathbf{pw}_0, \mathbf{pw}_1) \leq \delta$, \mathcal{P}_1 can extract the actual pass-string \mathbf{pw}_0 , as follows. If \mathcal{P}_1 plays the role of OT receiver and garbled circuit evaluator honestly, \mathcal{P}_0 and \mathcal{P}_1 will agree on $k_{0, \text{correct}}$. \mathcal{P}_1 can then mis-garble a circuit that returns $k_{1, \text{correct}}$ if the first bit of \mathbf{pw}_0 is 0, and $k_{1, \text{wrong}}$ if the first bit of \mathbf{pw}_0 is 1. By testing whether the resulting keys k_0 and k_1 match (which \mathcal{P}_1 can do in subsequent protocols where the key is used), \mathcal{P}_1 will

be able to determine the actual first bit of pw_0 . \mathcal{P}_1 can then repeat this for the second bit, and so on, extracting the entire pass-string pw_0 . Of course, if \mathcal{P}_1 does *not* know a sufficiently close pw_1 , \mathcal{P}_1 will not be able to perform these tests, because the keys will not match no matter what circuit \mathcal{P}_1 garbles.

More formally, if \mathcal{P}_1 knows a pass-string pw_1 such that $d(\text{pw}_0, \text{pw}_1) \leq \delta$ and carries out the mis-garbling attack described above, then in the real world, the keys produced by \mathcal{P}_0 and \mathcal{P}_1 either will or will not match based on some predicate p of \mathcal{P}_1 's choosing on the two pass-strings pw_0 and pw_1 . Therefore, in the ideal world, the keys should also match or not match based on $p(\text{pw}_0, \text{pw}_1)$; otherwise, the environment will be able to distinguish between the two worlds. In order to make that happen, since the simulator does not know the predicate p in question, the simulator must be able to recover the entire pass-string pw_0 (given a sufficiently close pw_1) through the `TestPwd` interface.

Theorem 1. *If $(\text{Gb}, \text{En}, \text{Ev}, \text{De})$ is a projective, output-projective and garbled-output random secure garbling scheme, then protocol Π_{RFE} with authenticated channels in the \mathcal{F}_{OT} -hybrid model securely realizes $\mathcal{F}_{\text{RFE}}^P$ with respect to static corruptions for any threshold δ , as long as the pass-string space and notion of distance are such that for any pass-string pw , it is easy to compute another pass-string pw' such that $d(\text{pw}, \text{pw}') > \delta$.⁵*

Proof (Sketch). For every efficient adversary \mathcal{A} , we describe a simulator \mathcal{S}_{RFE} such that no efficient environment can distinguish an execution with the real protocol Π_{RFE} and \mathcal{A} from an execution with the ideal functionality $\mathcal{F}_{\text{RFE}}^P$ and \mathcal{S}_{RFE} . \mathcal{S}_{RFE} is described in Figure 21. We prove indistinguishability in a series of hybrid steps. First, we introduce the ideal functionality as a dummy node. Next, we allow the functionality to choose the parties' keys, and we prove the indistinguishability of this step from the previous using the garbled output randomness property of our garbling scheme (Definition 9, Theorem 10). Next, we simulate an honest party's interaction with another honest party without using their pass-string, and prove the indistinguishability of this step from the previous using the obliviousness property of our garbling scheme. Finally, we simulate an honest party's interaction with a corrupted party without using the honest party's pass-string, and prove the indistinguishability of this step from the previous using the privacy property of our garbling scheme.

We give a more formal proof of Theorem 1 in Appendix C.

3.2.3 From Split Randomized Fuzzy Equality to fPAKE The Randomized Fuzzy Equality (RFE) functionality $\mathcal{F}_{\text{RFE}}^P$ assumes authenticated channels, which an fPAKE protocol cannot do. In order to adapt RFE to our setting, we use the split functionality transformation defined by Barak *et al.* [BCL⁺05]. Barak *et al.* provide a generic transformation from protocols which require authenticated channels to protocols which do not. In the “transformed” protocol, an adversary can engage in two separate instances of the protocol with the sender and receiver, and they will not realize that they are not talking to one another. However, it does guarantee that the adversary cannot do anything

⁵ This is used in the argument of indistinguishability of Games \mathbf{G}_2 and \mathbf{G}_3 in Appendix C.

The functionality $s\mathcal{F}_{\text{RFE}}^P$ is parameterized by a security parameter λ . It interacts with an adversary \mathcal{S} and two parties \mathcal{P}_0 and \mathcal{P}_1 via the following queries:

- **Initialization**
 - **Upon receiving a query (Init, sid) from a party $\mathcal{P}_i \in \{\mathcal{P}_0, \mathcal{P}_1\}$** , send (Init, sid, \mathcal{P}_i) to the adversary \mathcal{S} .
 - **Upon receiving a query (Init, sid, \mathcal{P}_i , H , sid $_H$) from the adversary \mathcal{S} :**
 - * Verify that $H \subseteq \{\mathcal{P}_0, \mathcal{P}_1\}$, that $\mathcal{P}_i \in H$, and that if a previous set H' was recorded, either (1) $H \cap H'$ contains only corrupted parties and $\text{sid}_H \neq \text{sid}_{H'}$, or (2) $H = H'$ and $\text{sid}_H = \text{sid}_{H'}$.
 - * If verification fails, do nothing.
 - * Otherwise, record the pair (H, sid_H) (if it was not already recorded), output (Init, sid, sid $_H$) to \mathcal{P}_i , and locally initialize a new instance of the original RFE functionality \mathcal{F}_{RFE} denoted $H\mathcal{F}_{\text{RFE}}^P$, letting the adversary play the role of $\{\mathcal{P}_0, \mathcal{P}_1\} - H$ in $H\mathcal{F}_{\text{RFE}}^P$.
- **RFE**
 - **Upon receiving a query from a party $\mathcal{P}_i \in \{\mathcal{P}_0, \mathcal{P}_1\}$** , find the set H such that $\mathcal{P}_i \in H$, and forward the query to $H\mathcal{F}_{\text{RFE}}^P$. Otherwise, ignore the query.
 - **Upon receiving a query from the adversary \mathcal{S} on behalf of \mathcal{P}_i corresponding to set H** , if $H\mathcal{F}_{\text{RFE}}^P$ is initialized and $\mathcal{P}_i \notin H$, then forward the query to $H\mathcal{F}_{\text{RFE}}^P$. Otherwise, ignore the query.

Fig. 6. Functionality $s\mathcal{F}_{\text{RFE}}^P$

beyond this attack. In other words, it provides “session authentication”, meaning that each party is guaranteed to carry out the entire protocol with the same partner, but not “entity authentication”, meaning that the identity of the partner is not guaranteed.

Barak *et al.* achieve this transformation in three steps. First, the parties generate signing and verification keys, and send one another their verification keys. Next, the parties sign the list of all keys they have received (which, in a two-party protocol, consists of only one key), sign that list, and send both list and signature to all other parties. Finally, they verify all of the signatures they have received. After this process — called “link initialization” — has been completed, the parties use those public keys they have exchanged to authenticate subsequent communication.

We describe the Randomized Fuzzy Equality Split Functionality in Figure 6. It is simplified from Figure 1 in Barak *et al.* [BCL⁺05] because we only need to consider two parties and static corruptions.

It turns out that $s\mathcal{F}_{\text{RFE}}^P$ is enough to realize $\mathcal{F}_{\text{fPAKE}}^P$. In fact, the protocol Π_{RFE} with the split functionality transformation directly realizes $\mathcal{F}_{\text{fPAKE}}^P$. In Appendix D, we prove that this is the case.

3.3 An Efficient Circuit f for Hamming Distance

The Hamming distance of two pass-strings $\text{pw}, \text{pw}' \in \mathbb{F}_p^n$ is equal to the number of locations at which the two pass-strings have the same character. More formally,

$$d(\text{pw}, \text{pw}') := |\{j \mid \text{pw}[j] \neq \text{pw}'[j], j \in [n]\}|.$$

We design f for Hamming distance as follows:

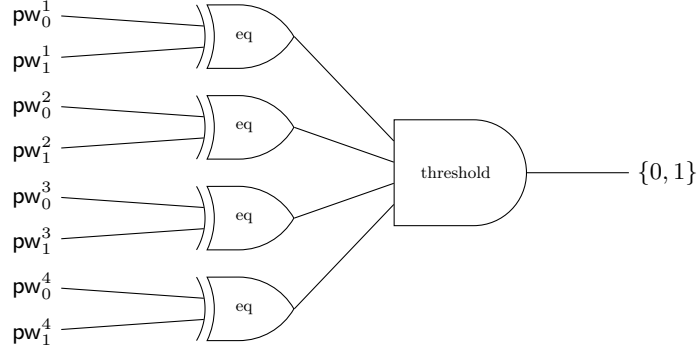


Fig. 7. The f circuit

1. First, f XORs corresponding (binary) pass-string characters, resulting in a list of bits indicating the (in)equality of those characters.
2. Then, f feeds those bits into a threshold gate, which returns 1 if at least $n - \delta$ of its inputs are 0, and returns 0 otherwise. f returns the output of that threshold gate, which is 1 if and only if at least $n - \delta$ pass-string characters match.

This circuit, illustrated in Figure 7, is very efficient to garble; it only requires n ciphertexts. Below, we briefly explain this garbling. Our explanation assumes familiarity with YGC literature [Yak17, and references therein]. Briefly, garbled gadget labels [BMR16] enable the evaluation of modular addition gates for free (there is no need to include any information in the garbled circuit to enable this addition). However, for a small modulus m , converting the output of that addition to a binary decision requires $m - 1$ ciphertexts. We utilize garbled gadgets with a modulus of $n + 1$ in our efficient garbling as follows:

1. The input wire labels encode 0 or 1 modulo $n + 1$. However, instead of having those input wire labels encode the characters of the two pass-strings directly, they encode the outputs of the comparisons of corresponding characters. If the j th character of \mathcal{P}_i 's pass-string is 0, then \mathcal{P}_i puts the 0 label first; however, if the j th character of \mathcal{P}_i 's pass-string is 1, then \mathcal{P}_i flips the labels. Then, when \mathcal{P}_{1-i} is using oblivious transfer to retrieve the label corresponding to her j th pass-string character, she will retrieve the 0 label if the two characters are equal, and the 1 label otherwise. (Note that this pre-processing on the garbler's side eliminates the need to send $X_{0,0}$ and $X_{1,1}$ in Figure 5.)
2. Compute a n -input threshold gate, as illustrated in Figure 6 of Yakoubov [Yak17]. This gate returns 0 if the sum of the inputs is above a certain threshold (that is, if at least $n - \delta$ pass-string characters differ), and 1 otherwise. This will require n ciphertexts.

Thus, a garbling of f consists of n ciphertexts. Since fPAKE requires two such garbled circuits (Figure 5), $2n$ ciphertexts will be exchanged.

Larger Pass-string Characters. If larger pass-string characters are used, then Step 1 above needs to change to check (in)equality of the larger characters instead of bits. Step 2 will

remain the same. There are several ways to perform an (in)equality check on characters in \mathbb{F}_p for $p \geq 2$:

1. Represent each character in terms of bits. Step 1 will then consist of XORing corresponding bits, and taking an OR or the resulting XORs of each character to get negated equality. This will take an additional $n \log(p)$ ciphertexts for every pass-string character.
2. Use garbled gadget labels from the outset. We will require a larger OT (1-out-of- p instead of 1-out-of-2), but nothing else will change.

4 Specialized Construction For Hamming Distance

In Appendix H, we show that it is not straightforward to build a secure fPAKE from primitives that are, by design, well-suited for correcting errors. However, PAKE protocols are appealingly efficient compared to the garbled circuits used in the prior construction. In this section, we ask whether the failed approach can be rescued in an efficient way, and we answer this question in the affirmative.

4.1 Building Blocks

4.1.1 Robust Secret Sharing We recall the definition of a robust secret sharing scheme, slightly simplified for our purposes from Cramer *et al.* [CDD⁺15]. For a vector $c \in \mathbb{F}_q^n$ and a set $A \subseteq [n]$, we denote with c_A the projection $\mathbb{F}_q^n \rightarrow \mathbb{F}_q^{|A|}$, i.e., the sub-vector $(c_i)_{i \in A}$.

Definition 2. Let $\lambda \in \mathbb{N}$, q a λ -bit prime, \mathbb{F}_q a finite field and $n, t, m, r \in \mathbb{N}$ with $t < r \leq n$ and $m < r$. An $(n, t, r)_q$ robust secret sharing scheme (RSS) consists of two probabilistic algorithms $\text{Share} : \mathbb{F}_q \rightarrow \mathbb{F}_q^n$ and $\text{Reconstruct} : \mathbb{F}_q^n \rightarrow \mathbb{F}_q$ with the following properties:

- t -privacy: for any $s, s' \in \mathbb{F}_q, A \subset [n]$ with $|A| \leq t$, the projections c_A of $c \xleftarrow{\$} \text{Share}(s)$ and c'_A of $c' \xleftarrow{\$} \text{Share}(s')$ are identically distributed.
- r -robustness: for any $s \in \mathbb{F}_q, A \subset [n]$ with $|A| \geq r$, any c output by $\text{Share}(s)$, and any \tilde{c} such that $c_A = \tilde{c}_A$, it holds that $\text{Reconstruct}(\tilde{c}) = s$.

In other words, an $(n, t, r)_q$ -RSS is able to reconstruct the shared secret even if the adversary tampered with up to $n - r$ shares, while each set of t shares is distributed independently of the shared secret s and thus reveals nothing about it. We note that we allow for a gap, i.e., $r \geq t + 1$. Schemes with $r > t + 1$ are called *ramp* RSS.

Definition 3 (Smoothness). We say that an $(n, t, r)_q$ -RSS is

- m -smooth if for any $s \in \mathbb{F}_q, A \subset [n]$ with $|A| \leq m$, any c output by $\text{Share}(s)$, and any \tilde{c} such that $\tilde{c}_A = c_A$, $\tilde{c}_{\bar{A}} \xleftarrow{\$} \mathbb{F}_q^{n-|A|}$, for all PPT \mathcal{A} it holds that

$$|\Pr[1 \leftarrow \mathcal{A}(1^\lambda, \text{Reconstruct}(\tilde{c}))] - \Pr[1 \leftarrow \mathcal{A}(1^\lambda, u)]|$$

is negligible in λ , where the probability is taken over the random coins of \mathcal{A} and Reconstruct and $u \xleftarrow{\$} \mathbb{F}_q$.

- m -smooth on random secrets if it is m -smooth for randomly chosen $s \xleftarrow{\$} \mathbb{F}_q$ and the probabilities are additionally taken over the coins consumed by this choice.

Definition 4 (Strong t -privacy). We say that an $(n, t, r)_q$ -RSS has strong t -privacy, if for any $s \in \mathbb{F}_q$, $A \subset [n]$ with $|A| \leq t$, the projection c_A of $c \xleftarrow{\$} \text{Share}(s)$ is distributed uniformly randomly in $\mathbb{F}_q^{|A|}$.

Note that strong t -privacy implies t -privacy. The opposite does not necessarily hold. (To see that the opposite might not hold, imagine a **Share** algorithm creating shares that start with “I’m a share!”). Also note that, in case of random errors occurring, as long as there are fewer than t undisturbed shares, a strong t -private scheme actually hides the locations (and with this also the number) of errors.

4.1.2 Linear Codes A linear q -ary code of length n and rank k is a subspace C with dimension k of the vector space \mathbb{F}_q^n . The vectors in C are called codewords. The size of a code is the number of codewords it contains, and is thus equal to q^k . The weight of a word $w \in \mathbb{F}_q^n$ is the number of its non-zero components, and the distance between two words is the Hamming distance between them (equivalently, the weight of their difference). The minimal distance d of a linear code C is the minimum weight of its non-zero codewords, or equivalently, the minimum distance between any two distinct codewords.

A code for an alphabet of size q , of length n , rank k , and minimal distance d is called an $(n, k, d)_q$ -code. Such a code can be used to detect up to $d - 1$ errors (because if a codeword is sent and fewer than $d - 1$ errors occur, it will not get transformed to another codeword), and correct up to $\lfloor (d - 1)/2 \rfloor$ errors (because for any received word, there is a unique codeword within distance $\lfloor (d - 1)/2 \rfloor$). For linear codes, the encoding of a (row vector) word $W \in \mathbb{F}_q^k$ is performed by an algorithm $C.\text{Encode} : \mathbb{F}_q^k \rightarrow \mathbb{F}_q^n$, which is the multiplication of W by a so-called “generating matrix” $G \in \mathbb{F}_q^{k \times n}$ (which defines an injective linear map). This leads to a row-vector codeword $W \cdot G =: c \in C \subset \mathbb{F}_q^n$.

The Singleton bound states that for any linear code, it holds that $k + d \leq n + 1$. A *maximum distance separable* (or MDS) code satisfies $k + d = n + 1$. Since $d = n - k + 1$, MDS codes are fully described by the parameters (q, n, k) . Such an $(n, k)_q$ -MDS code can correct up to $\lfloor (n - k)/2 \rfloor$ errors; it can detect if there are errors whenever there are no more than $n - k$ of them. For a matrix G generating an MDS code, any set of k columns of G are linearly independent.

For a thorough introduction to linear codes and proofs of all statements in this short overview we refer the reader to [Rot06].

Observe that a linear code, due to the linearity of its encoding algorithm, is not a primitive designed to hide anything about the encoded message (e.g., a popular choice for the generating matrix is $G := (I_k | *)$ with I_k being the $k \times k$ identity matrix). However, we show in the following lemma how to turn an MDS code into a RSS scheme with additional smoothness guarantees.

Lemma 5. Let C be a $(n + 1, k)_q$ -MDS code. We set L to be the last column of the generating matrix G of the code C and we denote by C' the $(n, k)_q$ -MDS code whose

generating matrix G' is G without the last column. Let further algorithm **Decode** of the MDS code C' be of the following form:

1. On input a word $c \in \mathbb{F}_q^n$, **Decode** chooses $D \subseteq [n]$ with $|D| = k$.
2. Let G'_D denote the matrix obtained from G' by eliminating all columns with indices not in D . **Decode** now outputs $c_D \cdot G'^{-1}_D$.

Let **Share** and **Reconstruct** work as follows:

- **Share**(s) for $s \in \mathbb{F}_q$ first chooses a random row vector $W \in \mathbb{F}_q^k$ such that $W \cdot L = s$, and outputs $c \leftarrow C'.\text{Encode}(W)$ (equivalently, we can say that **Share**(s) chooses a uniformly random codeword of C whose last coordinate is s , and outputs the first n coordinates as c).
- **Reconstruct**(w) for $w \in \mathbb{F}_q^n$ first runs $C'.\text{Decode}(w)$. If it gets a vector W' , then output $s = W' \cdot L$, otherwise output $s \xleftarrow{\$} \mathbb{F}_q$.

Then **Share** and **Reconstruct** form a t -smooth, strongly t -private $(n, t, r)_q$ -RSS for $t = k - 1$ and $r = \lceil (n + k)/2 \rceil$ that is $(r - 1)$ -smooth on random secrets.

Proof. Let us consider all required properties from Definitions 2, 3 and 4.

- strong t -privacy: Assume $|A| = t$ (privacy for smaller A will follow immediately by adding arbitrary coordinates to it to get to size t). Let $J = A \cup \{n + 1\}$; note that $|J| = t + 1 = k$. Note that for the code C , any k coordinates of a codeword determine uniquely the input to **Encode** that produces this codeword (otherwise, there would be two codewords that agreed on k elements and thus had distance $n - k + 1$, which is less than the minimum distance of C). Therefore, the mapping given by $\text{Encode}_J : \mathbb{F}_q^k \rightarrow \mathbb{F}_q^{|J|}$ is bijective; thus coordinates in J are uniform when the input to **Encode** is uniform. The algorithm **Share** chooses the input to **Encode** uniformly subject to fixing the coordinate $n + 1$ of the output. Therefore, the remaining coordinates (i.e., the coordinates in A) are uniform.
- r -robustness: Note that C has minimum distance $n - k + 2$, and therefore C' has minimum distance $n - k + 1$ (because dropping one coordinate reduces the distance by at most 1). Therefore, C' can correct $\lfloor (n - k)/2 \rfloor = n - r$ errors. Since $c_A = \tilde{c}_A$ and $|A| \geq r$, there are at most $n - r$ errors in \tilde{c} , so the call to $C'.\text{Decode}(\tilde{c})$ made by **Reconstruct**(\tilde{c}) will output $\tilde{W} = W$. Then **Reconstruct**(\tilde{c}) will output $s = \tilde{W} \cdot L = W \cdot L$.
- t -smoothness: to prove this, we show that disturbing one share uniformly random already randomizes the output of **Reconstruct**. Let D denote the set chosen by **Decode**. Since every codeword is uniquely determined by k elements, the mappings $f_i : \mathbb{F}_q \rightarrow \mathbb{F}_q$, $x \mapsto \text{Encode}(G'^{-1}_D(c))_{n+1}$ with $c \leftarrow \mathbb{F}_q^k$, $c_i = x$ are bijective for all $i \in [k]$. Since $t = k - 1$, \tilde{c}_D contains at least one entry that is chosen uniformly at random and thus the claim follows from the fact that the output of **Reconstruct** is computed as $\text{Encode}(G'^{-1}_D(\cdot))$.
- $(r - 1)$ -smoothness on random secrets: first, it holds that $r - 1 > k$ and thus \tilde{c} contains more than k undisturbed shares. We distinguish two cases. Either D chosen

by **Decode** contains only undisturbed shares (i.e., $D \subseteq A$), then **Reconstruct** will output s which is distributed uniformly random in \mathbb{F}_q . Else, $D \not\subseteq A$. In this case, at least one element of \tilde{c}_D is distributed uniformly random and the randomness of the output of **Reconstruct** follows as in the proof of t -smoothness.

Note that the Shamir’s secret sharing scheme is exactly the above construction with Reed-Solomon codes [MS81]. Further, we are not aware of any decoding algorithm for linear MDS codes that cannot be adopted to comply with our restrictions on **Decode**.

4.1.3 Implicit-Only PAKE PAKE protocols can have two types of authentication: implicit authentication, where at the end of the protocol the two parties share the same key if they used the same pass-string and random independent keys otherwise; or explicit authentication where, in addition, they actually know which of the two situations they are in. A PAKE protocol that only achieves implicit authentication can provide explicit authentication by adding key-confirmation flows [BPR00].

The standard PAKE functionality $\mathcal{F}_{\text{pwKE}}$ from [CHK⁺05] (see Figure 17) is designed with explicit authentication in mind, or at least considers that success or failure will later be detected by the adversary when he will try to use the key. Thus, it reveals to the adversary whether a pass-string guess attempt was successful or not. However, some applications could require a PAKE that does not provide any feedback, and so does not reveal the situation before the keys are actually used. Observe that, regarding honest players, $\mathcal{F}_{\text{pwKE}}$ already features implicit authentication since the players do not learn anything but their own session key.

Definition of implicit-only PAKE. We introduce a new notion called implicit-only PAKE or iPAKE (see Figure 8). The iPAKE ideal functionality is designed to implement implicit authentication with respect to the adversary as well as honest players, by not providing the adversary with any feedback in response to a password guess. Of course, in many cases, the parties and the adversary can later check whether their session keys match or not, and so whether the pass-strings were the same or not. We stress that this is not leakage from the PAKE protocol itself, but from the environment.

In addition to the changes we already make to the $\mathcal{F}_{\text{pwKE}}$ ideal functionality in our fPAKE functionality description, we make one more change in our iPAKE functionality. In response to a **TestPw** query, the functionality silently updates the internal state of the record (from **fresh** to either **compromised** or **interrupted**), meaning that the query outcome is not given to the adversary \mathcal{S} . Without going into details, it is clear that the simulation of an honest party is hard if the simulator \mathcal{S} does not know whether the password it extracted from the corrupt party is correct or not. However, the simulator gets help from the functionality, which sets honest parties’ keys appropriately.

We further alter this functionality to allow for public labels, as shown in Appendix F, Figure 26. The resulting functionality $\mathcal{F}_{\ell\text{-iPAKE}}$ idealizes what we call *labeled implicit-only PAKE* (or ℓ -iPAKE for short), resembling the notion of labeled public key encryption as formalized in [Sho01]. In a nutshell, labels are public authenticated strings that are chosen by each user individually for each execution of the protocol. Authentication here

The functionality $\mathcal{F}_{\text{iPAKE}}$ is parameterized by a security parameter λ . It interacts with an adversary \mathcal{S} and the (dummy) parties \mathcal{P}_0 and \mathcal{P}_1 via the following queries:

- **Upon receiving a query** ($\text{NewSession}, \text{sid}, \mathcal{P}_i, \text{role}$) **from party \mathcal{P}_i :**
 - Send ($\text{NewSession}, \text{sid}, \mathcal{P}_i, \text{role}$) to \mathcal{S} ;
 - If one of the following is true, record $(\mathcal{P}_i, \text{pw}_i)$ and mark this record **fresh**:
 - * This is the first **NewSession** query
 - * This is the second **NewSession** query and there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$
- **Upon receiving a query** ($\text{TestPwd}, \text{sid}, \mathcal{P}_i, \text{pw}'_i$) **from \mathcal{S} :**
If there is a **fresh** record $(\mathcal{P}_i, \text{pw}_i)$, then:
 - If $\text{pw}_i = \text{pw}'_i$, mark the record **compromised**;
 - If $\text{pw}_i \neq \text{pw}'_i$, mark the record **interrupted**.
- **Upon receiving a query** ($\text{NewKey}, \text{sid}, \mathcal{P}_i, \text{sk}$) **from \mathcal{S} , where $|\text{sk}| = \lambda$:**
If there is no record of the form $(\mathcal{P}_i, \text{pw}_i)$, or if this is not the first **NewKey** query for \mathcal{P}_i , then ignore this query. Otherwise:
 - If at least one of the following is true, then output (sid, sk) to player \mathcal{P}_i :
 - * The record is **compromised**
 - * \mathcal{P}_i is corrupted
 - * The record is **fresh**, \mathcal{P}_{1-i} is corrupted, and there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ with $\text{pw}_i = \text{pw}_{1-i}$
 - If this record is **fresh**, both parties are honest, there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ with $\text{pw}_i = \text{pw}_{1-i}$, a key sk' was sent to \mathcal{P}_{1-i} , and $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ was **fresh** at the time, then output (sid, sk') to \mathcal{P}_i ;
 - In any other case, pick a new random key sk' of length λ and send (sid, sk') to \mathcal{P}_i .
 - Mark the record $(\mathcal{P}_i, \text{pw}_i)$ as **completed**.

Fig. 8. Functionality $\mathcal{F}_{\text{iPAKE}}$

means that tampering with the label can be efficiently detected. Such labels can be used to, e.g., distribute public information such as public keys reliably over unauthenticated channels.

A UC-Secure ℓ -iPAKE Protocol. In the seminal paper by Bellare and Merritt [BM92], the Encrypted Key Exchange protocol (EKE) is proposed, which is essentially a Diffie-Hellman [DH76] key exchange. The two flows of the protocol are encrypted using the pass-string as the encryption key with an appropriate symmetric encryption scheme. The EKE protocol has been further formalized by Bellare *et al.* [BPR00] under the name EKE2. We present its labeled variant in Figure 9. The idea of appending the label to the symmetric key is taken from [ACCP08]. We prove security of this protocol in the $\mathcal{F}_{\text{RO}}, \mathcal{F}_{\text{IC}}, \mathcal{F}_{\text{CRS}}$ -hybrid model. That is, we use an ideal random oracle functionality \mathcal{F}_{RO} to model the hash function, and ideal cipher functionality \mathcal{F}_{IC} to model the encryption scheme and assume a publicly available common reference string modeled by \mathcal{F}_{CRS} . Formal definitions of these functionalities are given in Appendix A.

Theorem 6. *If the CDH assumption holds in \mathbb{G} , the protocol EKE2 depicted in Figure 9 securely realizes $\mathcal{F}_{\ell\text{-iPAKE}}$ in the $\mathcal{F}_{\text{RO}}, \mathcal{F}_{\text{IC}}, \mathcal{F}_{\text{CRS}}$ -hybrid model with respect to static corruptions.*

We note that this result is not surprising, given that other variants of EKE2 have already been proven to UC-emulate $\mathcal{F}_{\text{pwKE}}$. Intuitively, a protocol with only two flows not

$A(\text{pw} \in \mathbb{F}_p)$	$B(\text{pw}' \in \mathbb{F}_p)$
$x \xleftarrow{\$} \mathbb{F}_P, \ell \leftarrow \mathcal{L}, X \leftarrow g^x$ $X^* \leftarrow \mathcal{E}_{\text{pw} \ell}(X)$	$y \xleftarrow{\$} \mathbb{F}_P, \ell' \leftarrow \mathcal{L}, Y \leftarrow g^y$ $Y^* \leftarrow \mathcal{E}_{\text{pw}' \ell'}(Y)$
$Z \leftarrow \mathcal{D}_{\text{pw} \ell'}(Y^*)^x$ $k \leftarrow H(X^*, Y^*, Z)$ output (ℓ', k)	$Z' \leftarrow \mathcal{D}_{\text{pw}' \ell}(X^*)^y$ $k' \leftarrow H(X^*, Y^*, Z')$ output (ℓ, k')

Fig. 9. Protocol EKE2, in a group $\mathbb{G} = \langle g \rangle$ of prime order P , with a hash function $H : \mathbb{G}^3 \rightarrow \{0, 1\}^k$ and a symmetric cipher \mathcal{E}, \mathcal{D} onto \mathbb{G} for keys in $\mathbb{F}_p \times \mathcal{L}$, where \mathcal{L} is the label space.

depending on each other does not leak the outcome to the adversary via the transcript, which explains why EKE2 is implicit-only. Hashing of the transcript keeps the adversary from biasing the key unless he knows the correct pass-string or breaks the ideal cipher. For completeness, we include the full proof in Appendix F.

4.2 Construction

We show how to combine an RSS with a signature scheme and an ℓ -iPAKE to obtain an fPAKE. The high-level idea is to fix the issue that arose in the protocol from Appendix H due to pass-strings being used as one-time pads. Instead, we first expand the pass-string characters to keys — which we refer to as “character keys” — with large entropy using ℓ -iPAKE. The resulting character keys are then used as one-time pads for shares of a chosen output session key. We also apply known techniques from the literature, i.e., adding signatures and labels to prevent man-in-the-middle attacks. Our full protocol is depicted in Figure 10. It works as follows:

1. In the first phase, the two parties aim at enhancing their pass-strings to a vector of high-entropy character keys. For this, the pass-strings are viewed as vectors of characters. The parties repeatedly execute a PAKE on each of these characters separately. The PAKE will ensure that the character key vectors held by the two parties match in all positions where their pass-strings matched, and are independent in all other positions.
2. In the second phase of the protocol one party, the sender, will pick the final session key uniformly at random and send it in such a way that it reaches the other party only if enough of the character keys match. This is done by applying an RSS to the key, and sending it to the other party using the character key vector as a one time pad. The robustness property of the RSS ensures that a few non-matching password digits do not prevent the receiver from recovering the sender’s key.

When using the MDS code-based RSS which is described in Lemma 5, the one-time pad encryption of the shares (which form a codeword) can be viewed as the code-offset construction for information reconciliation (aka secure sketch) [JW99, DRS04] applied to

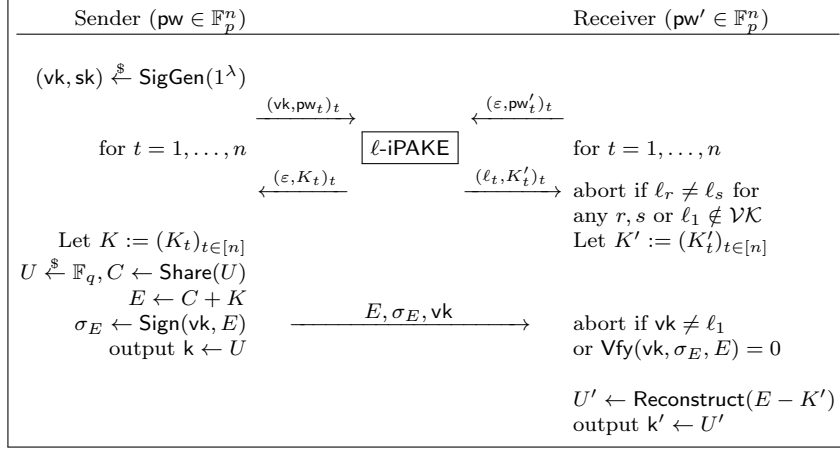


Fig. 10. Protocol $\text{fPAKE}_{\text{RSS}}$ where $q \approx 2^\lambda$ is a prime number and $+$ denotes the group operation in \mathbb{F}_q^n . ε denotes the empty string. $(\text{Share}, \text{Reconstruct})$ is a Robust Secret Sharing scheme with $\text{Share} : \mathbb{F}_q \rightarrow \mathbb{F}_q^n$, and $(\text{SigGen} \rightarrow \mathcal{VK} \times \mathcal{SK}, \text{Sign}, \text{Vfy})$ is a signature scheme. The parties repeatedly execute a labeled implicit-only PAKE protocol with label space \mathcal{VK} and key space \mathbb{F}_q , which takes inputs from $\mathcal{VK} \times \mathbb{F}_p$. If at any point an expected message fails to arrive (or arrives malformed), the parties output a random key.

the character key vectors. We present our construction in terms of RSS, but we could instead present this construction using information reconciliation. The syndrome construction of secure sketches 5 can also be used here instead of the code-offset construction.

4.3 Security of $\text{fPAKE}_{\text{RSS}}$

We show that our protocol realizes functionality $\mathcal{F}_{\text{fPAKE}}^M$ in the $\mathcal{F}_{\ell\text{-iPAKE}}$ -hybrid model. In a nutshell, the idea is to simulate without the pass-strings by adjusting the character keys outputted by $\mathcal{F}_{\ell\text{-iPAKE}}$ to the mask of the pass-strings, which is leaked by $\mathcal{F}_{\text{fPAKE}}^M$.

Theorem 7. *If $\text{RSS} := (\text{Share} : \mathbb{F}_q \rightarrow \mathbb{F}_q^n, \text{Reconstruct} : \mathbb{F}_q^n \rightarrow \mathbb{F}_q)$ is a strongly t -private, t -smooth $(n, t, r)_q$ -RSS that in addition is $(r - 1)$ -smooth on random secrets, and $(\text{SigGen}, \text{Sign}, \text{Vfy})$ is an EUF-CMA secure one-time signature scheme, protocol $\text{fPAKE}_{\text{RSS}}$ securely realizes $\mathcal{F}_{\text{fPAKE}}^M$ with $\gamma = n - t - 1$ and $\delta = n - r$ in the $\mathcal{F}_{\ell\text{-iPAKE}}$ -hybrid model with respect to static corruptions.*

In particular, if we wish key agreement to succeed as long as there are fewer than δ errors, we instantiate RSS using the construction of Lemma 5 based on a $(n + 1, k)_q$ MDS code, with $k = n - 2\delta$. This will give $r = \lceil (n + k)/2 \rceil = n - \delta$, so δ will be equal to $n - r$, as required. It will also give $\gamma = n - t - 1 = 2\delta$.

We thus obtain the following corollary:

Corollary 8. *For any δ and $\gamma = 2\delta$, given an $(n + 1, k)_q$ -MDS code for $k = n - 2\delta$ (with minimal distance $d = n - k + 2$) and an EUF-CMA secure one-time signature scheme, protocol $\text{fPAKE}_{\text{RSS}}$ securely realizes $\mathcal{F}_{\text{fPAKE}}^M$ in the $\mathcal{F}_{\ell\text{-iPAKE}}$ -hybrid model with respect to static corruptions.*

Proof sketch of Theorem 7. We start with the real execution of the protocol and indistinguishably switch to an ideal execution with dummy parties relaying their inputs to and obtaining their outputs from $\mathcal{F}_{\text{iPAKE}}^M$. To preserve the view of the distinguisher (the environment \mathcal{Z}), a simulator \mathcal{S} plays the role of the real world adversary by controlling the communication between $\mathcal{F}_{\text{iPAKE}}^M$ and \mathcal{Z} . During the proof, we build $\mathcal{F}_{\text{iPAKE}}^M$ and \mathcal{S} by subsequently randomizing pass-strings (since the final simulation has to work without them) and session keys (since $\mathcal{F}_{\text{iPAKE}}^M$ hands out random session keys in certain cases). We have to tackle the following difficulties, which we will describe in terms of attacks.

- Passive attack: in this attack, \mathcal{Z} picks two pass-strings and then observes the transcript and outputs of the protocol, without having access to any internal state of the parties. We show that \mathcal{Z} cannot distinguish between transcript and outputs that were either produced using \mathcal{Z} 's pass-strings or random pass-strings. Regarding the outputs, we argue that even in the real execution the session keys were chosen uniformly at random (with \mathcal{Z} not knowing the coins consumed by this choice) as long as the RSS either reconstructs the uniformly random secret correctly or outputs a different uniformly random value. Depending on the distance of the passwords this is guaranteed either by robustness or smoothness for random secrets, both of which are properties of the RSS. Regarding the transcript, randomization is straightforward using properties of the one-time pad.
- Man-in-the-middle attack: in this attack, \mathcal{Z} injects a malicious message into a session of two honest parties. There are several ways to secure protocols that have to run in unauthenticated channels and are prone to this attack. Basically, all of them introduce methods to bind messages together to prevent the adversary from injecting malicious messages. To do this, we need the *labeled* version of our iPAKE and a one-time signature scheme⁶. Unless \mathcal{Z} is able to break a one-time-signature scheme, this attack always results in an abort.
- Honest-but-curious/Active attack: in this attack, \mathcal{Z} corrupts one of the parties. The simulator will get help from $\mathcal{F}_{\text{iPAKE}}^M$ by issuing a `TestPwd` query, which will inform him whether the passwords used by both parties are close and, if so, in which positions they match (i.e., their mask).
 - If the sender is honest, we show how to use this information to simulate the transcript. Note that knowledge of the mask is necessary since, due to corruption, \mathcal{Z} can now actually decrypt the one-time pad and thus the transcript reveals the positions of the errors in the pass-strings, which are of course already known to \mathcal{Z} . If the simulator does not learn a mask, then the passwords are too far away and it follows from the strong privacy of the RSS that real and simulated transcript are indistinguishable from \mathcal{Z} 's view. Note that here, since \mathcal{Z} is aware of the non-matching pass-string positions, it is crucial that strong privacy guarantees that the locations of all errors are hidden from \mathcal{Z} .

⁶ Instead of labels and one-time signature, one could just sign all the messages, as would be done using the split-functionality [BCL⁺05], but this would be less efficient. This trade-off, with labels, is especially useful when we use a PAKE that admits adding labels basically for free, as it is the case with the special PAKE protocol we use.

- If the receiver is honest and \mathcal{Z} injects a malicious message on behalf of the sender, the simulator uses the mask to compute the output of the honest receiver. If no mask is obtained then again the pass-strings are too far away from each other, and the smoothness property of the RSS (for arbitrarily chosen secrets) says that the receiver’s output can be simulated by choosing it uniformly at random.

One interesting subtlety that arises is the usage of the iPAKE. Observe that the UC security notion for a regular PAKE as defined by Canetti *et al.* [CHK⁺05] and recalled in Appendix A provides an interface to the adversary to test a pass-string once and learn whether it is right or wrong. Using this notion, our simulator would have to answer to such queries from \mathcal{Z} . Since this is not possible without $\mathcal{F}_{\text{iPAKE}}^M$ leaking the mask all the time, it is crucial to use the iPAKE variant that we introduced in section 4.1.3. Using this stronger notion, the adversary is still allowed one pass-string guess which may affect the output, but the adversary learns nothing more about the outcome of his guess than he can infer from whatever access he has to the outputs alone. Since our protocol uses the outputs of the PAKE as one-time pad keys, it is intuitively clear that by preventing \mathcal{Z} from getting additional leakage about these keys, we protect the secrets of honest parties.

4.4 Further Discussion: Removing Modeling Assumptions

All of the assumptions in our protocol come from the realization of the $\mathcal{F}_{\ell\text{-iPAKE}}$ functionality. The $\ell\text{-iPAKE}$ protocol from section 4.1.3 requires a random oracle, an ideal cipher and a CRS. We note that we can remove everything except for the CRS by, e.g., taking the PAKE protocol introduced by Katz and Vaikuntanathan [KV11]. This protocol also securely realizes our $\mathcal{F}_{\ell\text{-iPAKE}}$ functionality⁷. However, it is more costly than our $\ell\text{-iPAKE}$ protocol since both messages each contain one non-interactive zero knowledge proof.

Since fPAKE implies a regular PAKE (simply set $\delta = 0$), Canetti *et al.* [CHK⁺05] give strong evidence that we cannot hope to realize $\mathcal{F}_{\text{fPAKE}}$ without a CRS.

5 Comparison of fPAKE Protocols

In this section, we give a brief comparison of our fPAKE protocols. First, in Figure 11, we describe the assumptions necessary for the two constructions, and the security parameters that they can achieve.

⁷ In a nutshell, their protocol is implicit-only for the same reason as the $\ell\text{-iPAKE}$ protocol we use here: there are only two flows that do not depend on each other, so the transcript cannot reveal the outcome of a guess unless it reveals the pass-string to anyone. Regarding the session keys, usage of a hash function takes care of randomizing the session key in case of a failed dictionary attack. Furthermore, the protocol already implements labels. A little more detailed, looking at the proof by Katz and Vaikuntanathan [KV11], the simulator does not make use of the answer of `TestPwd` to simulate any messages. Regarding the session key that an honest player receives in a corrupted session, they are chosen to be random in the simulation (in `Expt3`). Letting this happen already in the functionality makes the simulation independent of the answer of `TestPwd` also regarding the computation of the session keys.

	Assumptions	Threshold δ	Gap $\gamma - \delta$
$\text{fPAKE}_{\text{RSS}}$	UC-secure ℓ -iPAKE	$< n/2$	δ
$\text{fPAKE}_{\text{YGC}}$	(1) UC-secure OT (2) projective, output-projective and garbled-output random secure garbling scheme	Any	None

Fig. 11. Assumptions, Distance Thresholds and Functionality/Security Gaps achieved by the two schemes. $\text{fPAKE}_{\text{RSS}}$ is the construction in Figure 10. $\text{fPAKE}_{\text{YGC}}$ is the construction in Figure 5 with the split functionality transformation of Barak *et al.* [BCL⁺05].

Then, in Figure 12, we describe the efficiency of the constructions when concrete primitives (OT / ℓ -iPAKE) are used to instantiate them. $\text{fPAKE}_{\text{RSS}}$ is instantiated as the construction in Figure 10 with the ℓ -iPAKE in Figure 9 and an RSS. $\text{fPAKE}_{\text{YGC}}$ is instantiated as the construction in Figure 5 with the UC-secure oblivious transfer protocol of Chou and Orlandi [CO15] described in Figure 25, with the garbling scheme of Bal *et al.* [BMR16], and with the split functionality transformation of Barak *et al.* [BCL⁺05]. Though $\text{fPAKE}_{\text{YGC}}$ can handle any efficiently computable notion of distance, Figure 12 assumes that both constructions use Hamming distance (and that, specifically, $\text{fPAKE}_{\text{YGC}}$ uses the circuit described in Figure 7). We describe efficiency in terms of sub-operations (per-party, not in aggregate).

Note that these concrete primitives each have their own set of required assumptions. Specifically, the ℓ -iPAKE in Figure 9 requires a random oracle (RO, described in Figure 14), ideal cipher (IC, described in Figure 15) and common reference string (CRS, described in Figure 13). The oblivious transfer protocol in Figure 25 requires a random oracle. The garbling scheme of Bal *et al.* [BMR16] requires a mixed modulus circular correlation robust hash function, which is a weakening of the random oracle assumption.

For $\text{fPAKE}_{\text{RSS}}$, the factor of n arises from the n times EKE2 is executed. For $\text{fPAKE}_{\text{YGC}}$, the factor of n comes from the garbled circuit. Additionally, in $\text{fPAKE}_{\text{YGC}}$, three communication flows come from OT. The last of these is combined with sending the garbled circuits. Two additional flows of communication come from the split functionality transformation. The need for signatures also arises from the split functionality transformation.

Efficiency Optimizations to $\text{fPAKE}_{\text{YGC}}$ We can make several small efficiency improvements to the $\text{fPAKE}_{\text{YGC}}$ construction which are not reflected in Figure 12. First, instead of using the split functionality transformation of Barak *et al.* [BCL⁺05], we can use the split split functionality of Camenisch *et al.* [CCGS10]. It uses a split key exchange functionality to establish symmetric keys, and then uses those to symmetrically encrypt and authenticate each flow. While this does not save any communication rounds, it does reduce the number of public key operations needed. Second, if the pass-strings are more than λ bits long (where λ is the security parameter), OT extensions that are secure against malicious adversaries [AHMR15] can be used. If the pass-strings are fewer than λ bits long, then nothing is to be gained from using OT extensions, since OT extensions require λ “base OTs”. However, if the pass-strings are longer — say, if they are some

		Output Key Format	# (Bidirectional) Communication Flows	# Exponentiations	# Hashes	# Encryptions	# Decryptions	# Share	# Reconstruct	# SigKeyGens	# Signs	# SigVerifies
fPAKE _{RSS}	sender	\mathbb{F}_q	2	$2n$	n	n	n	1	0	1	1	0
	receiver							0	1	0	0	1
fPAKE _{YGC}		$\{0, 1\}^\lambda$	5	$3n + 2$	$4n + 7$	$2n$	n	—	—	1	5	5

Fig. 12. Efficiency (in Terms of Sub-Operations) of the Two Constructions. Here, by “bidirectional communication flow” we mean two flows, one in each direction, which do not depend on one another. fPAKE_{RSS} is the construction in Figure 10 instantiated with the ℓ -iPAKE in Figure 9. The first fPAKE_{RSS} row describes the sender’s efficiency; the second row describes the receiver’s efficiency. fPAKE_{YGC} is the construction in Figure 5 instantiated with the UC-secure oblivious transfer protocol of Chou and Orlandi [CO15] described in Figure 25, the garbling scheme of Bal *et al.* [BMR16], and with the split functionality transformation of Barak *et al.* [BCL⁺05]. fPAKE_{YGC} is described in a single row, since it is a symmetric protocol.

biometric measurement that is thousands of bits long — then OT extensions would save on the number of public key operations, at the cost of an extra round of communication.

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A Ideal UC Functionalities

COMMON REFERENCE STRING. The Common Reference String (CRS) functionality was already defined in [Can07]. We recall it in Figure 13 for completeness. Note that we do not let \mathcal{F}_{CRS} check whether a party is allowed to obtain the CRS — it is assumed public.

The functionality $\mathcal{F}_{\text{CRS}}^{\mathcal{D}}$ is parameterized with a distribution \mathcal{D} and proceeds as follows:

- Upon receiving (sid, crs) :
 - If there is no value r recorded, then choose and record a value $r \xleftarrow{\$} \mathcal{D}$.
 - Reply with (sid, r) .

Fig. 13. Functionality \mathcal{F}_{CRS}

RANDOM ORACLES. The Random Oracle (RO) functionality was already defined by Hofheinz and Müller-Quade in [HM04]. We recall it in Figure 14 for completeness. It is clear that the random oracle model UC-emulates this functionality.

IDEAL CIPHER. An ideal cipher [BPR00] is a block cipher that takes a plaintext or a ciphertext as input. We describe the ideal cipher functionality \mathcal{F}_{IC} in Figure 15, in the

The functionality \mathcal{F}_{RO} proceeds as follows, running on security parameter k , with a set of (dummy) parties $\mathcal{P}_1, \dots, \mathcal{P}_n$ and an adversary \mathcal{S} :

- \mathcal{F}_{RO} keeps a list L (which is initially empty) of pairs of bit strings.
 - Upon receiving a value (sid, m) (with $m \in \{0, 1\}^*$) from some party \mathcal{P}_i or from \mathcal{S} , do:
 - If there is a pair (m, \tilde{h}) for some $\tilde{h} \in \{0, 1\}^k$ in the list L , set $h := \tilde{h}$.
 - If there is no such pair, choose uniformly $h \in \{0, 1\}^k$ and store the pair $(m, h) \in L$.
- Once h is set, reply to the activating machine (i.e., either \mathcal{P}_i or \mathcal{S}) with (sid, h) .

Fig. 14. Functionality \mathcal{F}_{RO}

same vein as the above random oracle functionality. It is clear that the ideal cipher model UC-emulates this functionality. Note that this functionality characterizes a perfectly random permutation for each key by ensuring injectivity for each query simulation: to this aim, it uses a list L and projections M_{sk} and C_{sk} , that are global, independently of the sid .

The functionality \mathcal{F}_{IC} takes as input the security parameter k , and interacts with an adversary \mathcal{S} and with a set of (dummy) parties $\mathcal{P}_1, \dots, \mathcal{P}_n$ by means of these queries:

- \mathcal{F}_{IC} keeps a (initially empty) list L containing 3-tuples of bit strings and two (initially empty) sets C_{sk} and M_{sk} for every sk . (The sets are not created until sk is first used, thus avoiding the need to instantiate exponentially many sets.)
- **Upon receiving a query $(\text{sid}, \mathcal{E}, \text{sk}, m)$ (with $m \in \{0, 1\}^k$) from some party \mathcal{P}_i or \mathcal{S} , do:**
 - If there is a 3-tuple $(\text{sk}, m, \tilde{c})$ for some $\tilde{c} \in \{0, 1\}^k$ in the list L , set $c := \tilde{c}$.
 - If there is no such record, choose uniformly $c \in \{0, 1\}^k \setminus C_{\text{sk}}$ which is the set consisting of ciphertexts not already used with sk . Next, it stores the 3-tuple $(\text{sk}, m, c) \in L$ and sets both $M_{\text{sk}} \leftarrow M_{\text{sk}} \cup \{m\}$ and $C_{\text{sk}} \leftarrow C_{\text{sk}} \cup \{c\}$.

Once c is set, reply to the activating machine with (sid, c) .
- **Upon receiving a query $(\text{sid}, \mathcal{D}, \text{sk}, c)$ (with $c \in \{0, 1\}^k$) from some party \mathcal{P}_i or \mathcal{S} , do:**
 - If there is a 3-tuple $(\text{sk}, \tilde{m}, c)$ for some $\tilde{m} \in \{0, 1\}^k$ in L , set $m := \tilde{m}$.
 - If there is no such record, choose uniformly $m \in \{0, 1\}^k \setminus M_{\text{sk}}$ which is the set consisting of plaintexts not already used with sk . Next, it stores the 3-tuple $(\text{sk}, m, c) \in L$ and sets both $M_{\text{sk}} \leftarrow M_{\text{sk}} \cup \{m\}$ and $C_{\text{sk}} \leftarrow C_{\text{sk}} \cup \{c\}$.

Once m is set, reply to the activating machine with (sid, m) .

Fig. 15. Functionality \mathcal{F}_{IC}

OBLIVIOUS TRANSFER. The Oblivious Transfer (OT) functionality was defined by Canetti *et al.* [CLOS02]. We recall it in Figure 16.

PASSWORD-AUTHENTICATED KEY EXCHANGE. The initial PAKE functionality $\mathcal{F}_{\text{pwKE}}$ has been defined by Canetti *et al.* [CHK⁺05]. We recall it in Figure 17. We stress that this functionality immediately leaks the result of the **TestPwD**-query, which models explicit authentication; when the adversary tries a password, it learns whether the guess was correct or not.

The functionality \mathcal{F}_{OT} is parameterized by a security parameter λ . It interacts with an adversary \mathcal{S} and the players \mathcal{S} (the sender) and \mathcal{R} (the receiver) via the following queries:

- **Upon receiving a query (Send, sid, x_0, x_1) from \mathcal{S} ,** where $x_0, x_1 \in \{0, 1\}^\lambda$, record the tuple (x_0, x_1) .
- **Upon receiving a query (Receive, sid, i) from \mathcal{R} :**
If there is a record (x_0, x_1) , then send (sid, x_i) to \mathcal{R} and sid to \mathcal{S} , and halt. Otherwise, ignore the query.

Fig. 16. Functionality \mathcal{F}_{OT}

The functionality $\mathcal{F}_{\text{pwKE}}$ is parameterized by a security parameter k . It interacts with an adversary \mathcal{S} and a set of (dummy) parties $\mathcal{P}_1, \dots, \mathcal{P}_n$ via the following queries:

- **Upon receiving a query (NewSession, sid, $\mathcal{P}_i, \mathcal{P}_j, \text{pw}$, role) from party \mathcal{P}_i :**
 - Send (NewSession, sid, $\mathcal{P}_i, \mathcal{P}_j, \text{role}$) to \mathcal{S} ;
 - If one of the following is true, record $(\mathcal{P}_i, \mathcal{P}_j, \text{pw})$ and mark this record **fresh**:
 - * This is the first NewSession query
 - * This is the second NewSession query and there is a record $(\mathcal{P}_j, \mathcal{P}_i, \text{pw}')$
- **Upon receiving a query (TestPwd, sid, $\mathcal{P}_i, \text{pw}'$) from \mathcal{S} :**
If there is a **fresh** record $(\mathcal{P}_i, \mathcal{P}_j, \text{pw})$, then do:
 - If $\text{pw} = \text{pw}'$, mark the record **compromised**, and reply to \mathcal{S} with *correct guess*;
 - If $\text{pw} \neq \text{pw}'$, mark the record **interrupted**, and reply to \mathcal{S} with *wrong guess*.
- **Upon receiving a query (NewKey, sid, \mathcal{P}_i, sk) from the \mathcal{S} , where $|\text{sk}| = k$:**
If there is a record of the form $(\mathcal{P}_i, \mathcal{P}_j, \text{pw})$, and this is the first NewKey query for \mathcal{P}_i , then:
 - If this record is **compromised**, or either \mathcal{P}_i or \mathcal{P}_j is corrupted, then output (sid, sk) to player \mathcal{P}_i ;
 - If this record is **fresh**, there is a record $(\mathcal{P}_j, \mathcal{P}_i, \text{pw}')$ with $\text{pw} = \text{pw}'$, a key sk' was sent to \mathcal{P}_j , and $(\mathcal{P}_j, \mathcal{P}_i, \text{pw})$ was **fresh** at the time, then output (sid, sk') to \mathcal{P}_i ;
 - In any other case, pick a new random key sk' of length k and send (sid, sk') to \mathcal{P}_i .
 Either way, mark the record $(\mathcal{P}_i, \mathcal{P}_j, \text{pw})$ as **completed**.

Fig. 17. Functionality $\mathcal{F}_{\text{pwKE}}$

In our paper, we start from this functionality to derive the basic functionality fPAKE , in Section 2, after a few changes:

- we consider only two parties — \mathcal{P}_0 and \mathcal{P}_1 —, which is enough since universal composability takes care of ensuring that a two-party functionality remains secure in a multi-party world;
- we do not allow the adversary to set \mathcal{P}_i 's key if \mathcal{P}_{1-i} is corrupted but did not guess \mathcal{P}_i 's password. We make this change in order to protect an honest \mathcal{P}_i from, for instance, revealing sensitive information to an adversary who did not successfully guess her password, but did corrupt her partner.

B Garbled Output Randomness: A New Yao's Garbled Circuit Definition

We refer to Yakoubov [Yak17] for a gentle introduction to Yao's Garbled Circuits. Note that the authenticity property implies that, in an output-projective garbling scheme, if the output is a single bit, the second output label and the second token of d are hard for the evaluator to guess (no probabilistic polynomial-time adversary can guess it with non-negligible probability). However, for our fPAKE construction (Section 3), we require a stronger property: not only should the second output label be hard to guess, but it should be *indistinguishable from random*. We call this *garbled-output randomness*.

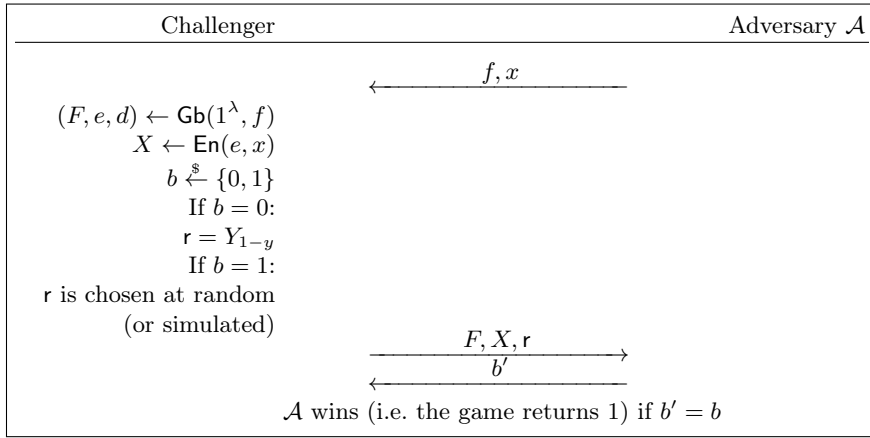


Fig. 18. The $\text{GOutRand}_{\mathcal{G}}^{\mathcal{A}}(1^\lambda)$ Game, where $y = f(x) \in \{0, 1\}$, Y_y is the corresponding garbled output (or output label), and Y_{1-y} is the other output label.

Define the adversary \mathcal{A} 's advantage in the the garbled-output randomness game (Figure 18) as

$$\text{GOutRandAdv}_{\mathcal{G}}(1^\lambda, \mathcal{A}) = \left| \Pr[\text{GOutRand}_{\mathcal{G}}^{\mathcal{A}}(1^\lambda) = 1] - \frac{1}{2} \right|.$$

Definition 9. An output-projective binary output garbling scheme $\mathcal{G} = (\text{Gb}, \text{En}, \text{Ev}, \text{De})$ is garbled-output random if for all sufficiently large security parameters λ , for any polynomial time adversary \mathcal{A} ,

$$\text{GOutRandAdv}_{\mathcal{G}}(1^\lambda, \mathcal{A}) = \text{negl}.$$

In order to achieve this, we modify the scheme of Bal *et al.* [BMR16] to put the output wire label through the hash function H one more time; the two labels will thus be $Y_0 = H(\text{finaloutput}, W_{\text{output}}^0)$ and $Y_1 = H(\text{finaloutput}, W_{\text{output}}^0 \oplus R)$, where W_{output}^0 and $W_{\text{output}}^0 \oplus R$ were the labels in the scheme of Bal *et al.*

Theorem 10. This modified scheme is garbled-output random (Definition 9) when the key derivation function H is mixed-modulus circular correlation robust (Definition 1 of Bal *et al.* [BMR16]).

Proof (Proof Sketch).

Definition 9 requires the indistinguishability of (real garbled circuit and inputs, real second output label) and (real garbled circuit and inputs, random second output label); in shorthand, we want to show that $(\text{real}, \text{real}) \sim (\text{real}, \text{random})$, where \sim denotes computational indistinguishability.

We use the setting (simulated garbled circuit and inputs, random second output label) — $(\text{simulated}, \text{random})$ for short — as a hybrid step. We show that $(\text{real}, \text{real}) \sim (\text{simulated}, \text{random})$. We do this by having the adversary \mathcal{A} (modeled after the one in Choi *et al.* [CKKZ12]) compute the second output label $\mathcal{O}(\text{finaloutput}, 2, 2, W_{\text{output}}^b, 1, 0)$ (using the mixed-modulus circular correlation robustness oracle) and send it to the obliviousness adversary \mathcal{B} along with the garbled circuit and garbled inputs. If the oracle is random, \mathcal{B} will see $(\text{simulated}, \text{random})$. Otherwise, \mathcal{B} will see $(\text{real}, \text{real})$. If \mathcal{B} can distinguish between those two, then \mathcal{A} can use that to break mixed-modulus circular correlation robustness. Hence, $(\text{real}, \text{real}) \sim (\text{simulated}, \text{random})$.

Because the garbling scheme is oblivious, we know that $(\text{simulated}, \text{random}) \sim (\text{real}, \text{random})$, since we can always add a random value to the adversary's view in the obliviousness game.

Now that we have $(\text{real}, \text{real}) \sim (\text{simulated}, \text{random})$ and $(\text{simulated}, \text{random}) \sim (\text{real}, \text{random})$, we can conclude that $(\text{real}, \text{real}) \sim (\text{real}, \text{random})$.

C Proof of Theorem 1

We proceed in a series of games, where no probabilistic polynomial-time environment can distinguish the view of the adversary \mathcal{A} in each game from that in the previous game. We start with the real execution of the protocol and end with the ideal execution. Figure 19 summarizes the changes made in each game.

Game $\mathbf{G_0}$: Real

This is the real execution of Π_{RFE} where the environment \mathcal{Z} runs the protocol (described in Figure 5) with parties \mathcal{P}_0 and \mathcal{P}_1 , both having access to an ideal OT

	Game	Functionality \mathcal{F}			\mathcal{S}_{RFE}	Property Used
		NewSession	TestPwd	NewKey		
	Game \mathbf{G}_0	N/A	N/A	N/A	N/A	
	Game \mathbf{G}_1	forwards inputs to \mathcal{S}_{RFE}		forwards outputs to dummy parties	runs protocol for honest parties	
	Game \mathbf{G}_2	records inputs			creates NewKey queries from party outputs	
both parties honest	Game \mathbf{G}_3			chooses keys for both parties when $d(\text{pw}_0, \text{pw}_1) \leq \delta$		garbled output randomness
	Game \mathbf{G}_4			chooses key for \mathcal{P}_0 when $d(\text{pw}_0, \text{pw}_1) > \delta$		garbled output randomness
	Game \mathbf{G}_5			chooses keys for both parties when $d(\text{pw}_0, \text{pw}_1) > \delta$		garbled output randomness
	Game \mathbf{G}_6				simulates F_0, X_0	obliviousness
	Game \mathbf{G}_7	does not forward pw_0, pw_1			simulates F_1, X_1	obliviousness
\mathcal{P}_i honest, \mathcal{P}_{1-i} corrupt	Game \mathbf{G}_8	replaces the malicious NewSession input with the one given by \mathcal{S}_{RFE}			extracts malicious pw'_{1-i} from OT, and tells \mathcal{F} to replace the malicious NewSession input with pw'_{1-i}	
	Game \mathbf{G}_9			chooses key for \mathcal{P}_i when $d(\text{pw}_0, \text{pw}_1) > \delta$ (now fully implemented)		garbled output randomness
	Game \mathbf{G}_{10}				simulates F_i, X_i using $d(\text{pw}_i, \text{pw}'_{1-i})$	privacy
	Game \mathbf{G}_{11}	does not forward pw_i	fully implemented		makes TestPwd query to set pw'_i	

Fig. 19. A Summary of the Sequence of Games in the Proof of Theorem 1

functionality \mathcal{F}_{OT} , and an adversary \mathcal{A} that, w.l.o.g., can be assumed to be the dummy adversary as shown in [Can01, section 4.4.1].

Game \mathbf{G}_1 : Adding Ideal Layout

This is the real game, but with dummy party and ideal functionality nodes thrown in and all previously existing nodes (except the environment) grouped into one machine, called the simulator (\mathcal{S}_{RFE} , or \mathcal{S} for short). Please refer to Figure 20 for the differences between \mathbf{G}_0 and \mathbf{G}_1 .

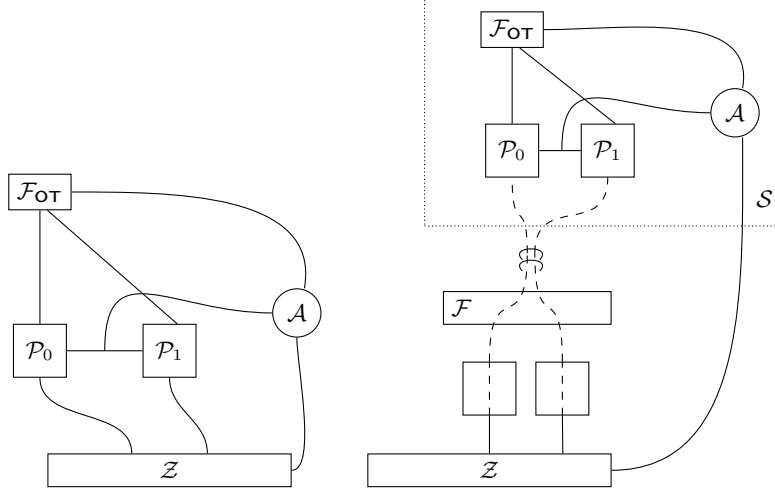


Fig. 20. Transition from game \mathbf{G}_0 (left) to game \mathbf{G}_1 (right), showing a setting where both parties are honest.

Game \mathbf{G}_2 : Adding \mathcal{F} 's Record-Keeping and TestPwd Interface

Modifications to \mathcal{F} : We now allow \mathcal{F} to do all of the record-keeping described in Figure 4.

\mathcal{F} still forwards **NewSession** queries from the dummy parties in their entirety (including the pass-string) to \mathcal{S}_{RFE} , but also records them. Since this is a matter of internal record-keeping only, this does not affect \mathcal{A} 's view.

Modifications to \mathcal{S}_{RFE} : \mathcal{S}_{RFE} creates **NewKey** queries for \mathcal{F} from whatever output the simulated parties produce. In this game, \mathcal{F} still simply forwards the keys it is given to the dummy parties without modifying them, so this does not affect \mathcal{A} 's view.

Game \mathbf{G}_3 : Allowing \mathcal{F} to Choose Keys For Two Honest Parties With Close Pass-strings

Modifications to \mathcal{F} : We now allow \mathcal{F} to follow the instructions in Figure 4 to choose the key when \mathcal{P}_0 and \mathcal{P}_1 are both honest, and $d(\text{pw}_0, \text{pw}_1) \leq \delta$.

We can use any environment who can distinguish this game from Game \mathbf{G}_2 to build an adversary \mathcal{B} that can break the garbled output randomness property (Definition 9, Theorem 10) of our garbling scheme.

Since both parties are honest, in order to use the environment as a distinguisher, the adversary needs to give the environment a transcript of the parties' interactions

$(F_0, X_{0,0}, F_1, X_{1,1})$ as well as the parties' output keys. Because we are in the OT hybrid model, the environment sees neither inputs to the OT nor its outputs.

Our adversary \mathcal{B} executes \mathcal{S}_{RFE} 's simulation of \mathcal{P}_0 with the \mathcal{F} of Game \mathbf{G}_2 with some modifications. First, it finds a pass-string pw such that $d(\text{pw}_0, \text{pw}) > \delta$. (Note that in order for this reduction to work, such a pass-string must be efficiently computable, which it is by the assumption in the statement of Theorem 1.) Instead of running Gb , \mathcal{B} queries the garbled output randomness challenger on $(f, (\text{pw}_0, \text{pw}))$ to obtain (F_0, X_0, r) . Let $X_0 = (X_{0,0}, X_{0,1})$. Note that F_0 and $X_{0,0}$ are generated by the challenger exactly as they would be by \mathcal{S}_{RFE} , so those values do not change. ($X_{0,1}$ is different, since pw is different from pw_1 , but $X_{0,1}$ is not visible to the environment.) Since the adversary used a pass-string pw that is dissimilar to pw_1 , it uses the value r — corresponding to the output label *not* returned by $\text{Ev}(F_0, X_0)$ — as $k_{0,\text{correct}}$. ($\text{Ev}(F_0, X_0)$ would give $k_{0,\text{wrong}}$.) If $b = 0$, the challenger will return the actual $k_{0,\text{correct}}$ as r , and the environment's view will be that of Game \mathbf{G}_2 . If $b = 1$, the challenger will give a random value as r . If r is truly random, then so is $r \oplus Y_1$; so, the environment's view will be that of Game \mathbf{G}_3 . (Note that we do not change the way in which \mathcal{P}_1 generates F_1, X_1 , but we do set \mathcal{P}_1 's output key to be the same as \mathcal{P}_0 's. This will be the key that an honest execution of the protocol would produce if $b = 0$, and random otherwise.) The adversary \mathcal{B} then returns the environment's guess as b' . The advantage of \mathcal{B} in the garbled output randomness game will be exactly the same as that of the environment in distinguishing between Game \mathbf{G}_2 and Game \mathbf{G}_3 .

Game \mathbf{G}_4 : Allowing \mathcal{F} to Choose Keys For One of Two Honest Parties With Dissimilar Pass-strings

Modifications to \mathcal{F} : We now allow \mathcal{F} to follow the instructions in Figure 4 to choose the key for \mathcal{P}_0 when \mathcal{P}_0 and \mathcal{P}_1 are both honest, and $d(\text{pw}_0, \text{pw}_1) > \delta$.

We can use any environment who can distinguish this game from Game \mathbf{G}_3 to build an adversary \mathcal{B} that can break the garbled output randomness property (Definition 9, Theorem 10) of our garbling scheme.

Our adversary \mathcal{B} executes \mathcal{S}_{RFE} 's simulation of \mathcal{P}_0 with the \mathcal{F} of Game \mathbf{G}_3 with some modifications. Instead of running Gb , \mathcal{B} queries the garbled output randomness challenger on $(f, (\text{pw}_0, \text{pw}_1))$ to obtain (F_0, X_0, r) . Note that F_0 and X_0 are generated by the challenger exactly as they would be by \mathcal{S}_{RFE} , so these values do not change. If $b = 0$, the challenger will return the actual $k_{0,\text{correct}}$ as r , and the environment's view will be that of Game \mathbf{G}_3 . If $b = 1$, the challenger will give a random value as r . If r is truly random, then so is $r \oplus Y_1$; so, the environment's view will be that of Game \mathbf{G}_4 . The adversary \mathcal{B} then returns the environment's guess as b' . The advantage of \mathcal{B} in the garbled output randomness game will be exactly the same as that of the environment in distinguishing between Game \mathbf{G}_3 and Game \mathbf{G}_4 .

Game \mathbf{G}_5 : Allowing \mathcal{F} to Choose Keys For Both Honest Parties With Dissimilar Pass-strings

Modifications to \mathcal{F} : We now allow \mathcal{F} to follow the instructions in Figure 4 to choose the key for \mathcal{P}_1 as well as for \mathcal{P}_0 when \mathcal{P}_0 and \mathcal{P}_1 are both honest, and $d(\text{pw}_0, \text{pw}_1) > \delta$.

We can use any environment who can distinguish this game from Game \mathbf{G}_4 to build an adversary that can break the garbled output randomness property (Definition 9, Theorem 10) of our garbling scheme, exactly as we did in the reduction above.

Game \mathbf{G}_6 : Simulating F, X for One of Two Honest Parties

Modifications to \mathcal{S}_{RFE} : Consider the case when both \mathcal{P}_0 and \mathcal{P}_1 are honest. In this game, the simulator replaces \mathcal{P}_0 's garbled circuit and input with simulated ones. \mathcal{S}_{RFE} does not need to simulate anything relating to the OT, since the environment cannot observe OT functionality inputs or outputs if both participating parties are honest. \mathcal{S}_{RFE} uses the obliviousness simulator to generate F_0, X_0 (while continuing to generate F_1, X_1 honestly), and sends the garbled circuits and the appropriate parts of the garbled inputs between the parties. \mathcal{S}_{RFE} outputs \perp bot as both parties' keys, since the outputs don't matter - \mathcal{F} takes care of outputting appropriate keys as of a few games ago (Game \mathbf{G}_3 if $d(\text{pw}_0, \text{pw}_1) \leq \delta$, and Games $\mathbf{G}_4, \mathbf{G}_5$ otherwise), so this change is not observable by the environment.

We can use any environment who can distinguish this game from Game \mathbf{G}_5 to build an adversary \mathcal{B} that can break the obliviousness property of our garbling scheme. \mathcal{B} executes \mathcal{S}_{RFE} 's simulation of \mathcal{P}_0 as in Game \mathbf{G}_5 , but instead of generating (F_0, X_0) and (F_1, X_1) according to the protocol, it queries the obliviousness challenger on $(f, (\text{pw}_0, \text{pw}_1))$ to obtain (F_0, X_0) . If $b = 0$, the challenger will return actual (F_i, X_i) values, and the environment's view we will be that of Game \mathbf{G}_5 . If $b = 1$, the challenger will return simulated values, and the environment's view will be that of this game. The adversary \mathcal{B} then returns the environment's guess as b' . The advantage of \mathcal{B} in the obliviousness game will be exactly the same as that of the environment in distinguishing between Game \mathbf{G}_5 and Game \mathbf{G}_6 .

Game \mathbf{G}_7 : Removing Pass-string Forwarding Always

Modifications to \mathcal{F} : We now modify \mathcal{F} to forward only $(\text{NewSession}, \text{sid}, \mathcal{P}_i)$ to \mathcal{S}_{RFE} (omitting the pass-string pw_i) for $\mathcal{P}_i \in \{\mathcal{P}_0, \mathcal{P}_1\}$ when \mathcal{P}_0 and \mathcal{P}_1 are both honest.

Modifications to \mathcal{S}_{RFE} : In Game \mathbf{G}_6 , \mathcal{S}_{RFE} started simulating \mathcal{P}_0 's messages without using knowledge of pw_0 . \mathcal{S}_{RFE} now also simulates \mathcal{P}_1 's messages by using the obliviousness simulator to generate the garbled circuit and input F_1, X_1 .

We can use any environment who can distinguish this game from Game \mathbf{G}_6 to build an adversary \mathcal{B} that can break the obliviousness property of our garbling scheme. \mathcal{B} executes \mathcal{S}_{RFE} 's simulation of \mathcal{P}_1 as in Game \mathbf{G}_6 , but instead of generating (F_1, X_1) according to the protocol, it queries the obliviousness challenger on $(f, (\text{pw}_1, \text{pw}_0))$ to obtain (F_1, X_1) . If $b = 0$, the challenger will return actual (F_1, X_1) values, and the environment's view we will be that of Game \mathbf{G}_6 . If $b = 1$, the challenger will return simulated values, and the environment's view will be that of this game. The adversary \mathcal{B} then returns the environment's guess as b' . The advantage of \mathcal{B} in the obliviousness game will be exactly the same as that of the environment in distinguishing between Game \mathbf{G}_6 and Game \mathbf{G}_7 .

Game \mathbf{G}_8 : Setting the Malicious Input as in the "Standard Corruption Model"

Modifications to \mathcal{S}_{RFE} : In this game, \mathcal{S}_{RFE} sets a corrupt party \mathcal{P}_{1-i} 's NewSession input according to the standard corruption model [Can01]. It does so as soon as it

sees pw'_{1-i} , when it is given as an input to the ideal OT functionality. Since \mathcal{F} does not currently use pw_{1-i} , this does not affect the environment's view.

Game \mathbf{G}_9 : Allowing \mathcal{F} to Choose the Key For An Honest Party With a Pass-string Dissimilar to Its Corrupt Partners'

Modifications to \mathcal{F} : We now allow \mathcal{F} to follow the instructions in Figure 4 to choose all keys. Note that the only remaining scenario this affects is the one where only one party $\mathcal{P}_i \in \{\mathcal{P}_0, \mathcal{P}_1\}$ is honest, and $d(\text{pw}_0, \text{pw}_1) > \delta$. If both parties are corrupt, or if only one party is corrupt and $d(\text{pw}_0, \text{pw}_1) \leq \delta$, \mathcal{F} still simply forwards the output key.

We can use any environment who can distinguish this game from Game \mathbf{G}_8 to build an adversary \mathcal{B} that can break the garbled output randomness property (Definition 9, Theorem 10) of our garbling scheme. Our adversary \mathcal{B} executes \mathcal{S}_{RFE} 's simulation of \mathcal{P}_0 with the \mathcal{F} of Game \mathbf{G}_8 with some modifications. Instead of running \mathbf{Gb} first, it waits to see \mathcal{P}_{1-i} 's input pw'_{1-i} to the ideal OT functionality, and then queries the garbled output randomness challenger on $(f, (\text{pw}_i, \text{pw}'_{1-i}))$ to obtain (F_i, X_i, r) . Note that F_i and X_i are generated by the challenger exactly as they would be by \mathcal{S}_{RFE} , so those values do not change. The adversary then uses r as $k_{i, \text{correct}}$. If $b = 0$, the challenger will return the actual $k_{i, \text{correct}}$ as r , and the environment's view will be that of Game \mathbf{G}_8 . If $b = 1$, the challenger will give a random value as r . If r is truly random, then so is $r \oplus Y_{1-i}$; so, the environment's view will be that of Game \mathbf{G}_9 . The adversary \mathcal{B} then returns the environment's guess as b' . The advantage of \mathcal{B} in the garbled output randomness game will be exactly the same as that of the environment in distinguishing between Game \mathbf{G}_8 and Game \mathbf{G}_9 .

Game \mathbf{G}_{10} : Simulating Garbled Circuit and Inputs For An Honest Party With a Corrupt Partner

Modifications to \mathcal{S}_{RFE} : In this game, \mathcal{S}_{RFE} simulates F_i and X_i when \mathcal{P}_i is honest and \mathcal{P}_{1-i} is corrupt.

In more detail, \mathcal{S}_{RFE} proceeds as follows on behalf of \mathcal{P}_i :

- \mathcal{S}_{RFE} postpones step 1.
- In step 2, \mathcal{S}_{RFE} :
 - Plays an OT sender as follows:
 - * Waits for \mathcal{P}_{1-i} to provide their select bits to the OT. As a result, \mathcal{S}_{RFE} learns the pass-string used by \mathcal{P}_{1-i} , pw'_{1-i} .
 - * If $d(\text{pw}_i, \text{pw}'_{1-i}) \leq \delta$ sets $y = 1$, and sets $y = 0$ otherwise.
 - * Uses the privacy simulator for the garbling scheme to generate $(F_i, X_i, d_i) \leftarrow \mathcal{S}(1^\lambda, f, y)$.
 - * Parses $(X_{i,i}, X_{i,1-i}) \leftarrow X_i$.
 - * Sends $X_{i,1-i}$ to \mathcal{P}_{1-i} as the OT output.
 - Plays an OT receiver honestly.
- \mathcal{S}_{RFE} follows the instructions in Figure 5 for steps 3-8.

We can use any environment who can distinguish this game from Game \mathbf{G}_9 to build an adversary \mathcal{B} that can break the privacy property of our garbling scheme. Our adversary \mathcal{B} executes \mathcal{S}_{RFE} 's simulation of \mathcal{P}_0 with some modifications. Instead of

running the privacy simulator \mathcal{S} , it queries the privacy challenger on $(f, (\text{pw}_i, \text{pw}'_{1-i}))$ to obtain (F_i, X_i, d_i) . If $b = 0$, the challenger will return actual (F_i, X_i, d_i) values, and the environment's view will be that of Game \mathbf{G}_9 ; if $b = 1$, the challenger return simulated values, and the environment's view will be that of this game. The adversary \mathcal{B} then returns the environment's guess as b' . The advantage of \mathcal{B} in the privacy game will be exactly the same as that of the environment in distinguishing between Game \mathbf{G}_9 and Game \mathbf{G}_{10} .

Game \mathbf{G}_{11} : Removing Pass-string Forwarding To An Honest Party With a Corrupt Partner

Modifications to \mathcal{F} :

- If \mathcal{P}_i is honest and \mathcal{P}_{1-i} is corrupt, then, upon receiving a **NewSession** query, \mathcal{F} forwards only $(\text{NewSession}, \text{sid}, \mathcal{P}_i)$ to \mathcal{S}_{RFE} (omitting the pass-string pw_i).
- \mathcal{F} now processes **TestPwd** queries (which were not issued in any prior game) according to the instructions in Figure 4. Given a $(\text{TestPwd}, \text{sid}, \mathcal{P}_i)$ query ($\mathcal{P}_i \in \{\mathcal{P}_0, \mathcal{P}_1\}$), if $d(\text{pw}_0, \text{pw}_1) \leq \delta$, \mathcal{F} sends pw_i to \mathcal{S}_{RFE} .

Modifications to \mathcal{S}_{RFE} : Now that \mathcal{S}_{RFE} does not know \mathcal{P}_i 's pass-string, it must simulate the honest party's messages without that knowledge.

In more detail, \mathcal{S}_{RFE} proceeds as follows on behalf of \mathcal{P}_i :

- \mathcal{S}_{RFE} postpones step 1.
- In step 2, \mathcal{S}_{RFE} :
 - Plays an OT sender as follows:
 - * As in Game \mathbf{G}_{10} , waits for \mathcal{P}_{1-i} to provide their select bits to the OT. As a result, \mathcal{S}_{RFE} learns the pass-string used by \mathcal{P}_{1-i} , pw'_{1-i} .
 - * Makes a $(\text{TestPwd}, \text{sid}, \mathcal{P}_i)$ query to \mathcal{F} . If $d(\text{pw}_i, \text{pw}'_{1-i}) \leq \delta$ (that is, the adversary has approximately guessed \mathcal{P}_i 's pass-string), \mathcal{S}_{RFE} learns pw_i , and sets $\text{pw}'_i = \text{pw}_i$. Otherwise, it sets pw'_i at random such that $d(\text{pw}'_i, \text{pw}'_{1-i}) > \delta$, and uses pw'_i in place of pw_i in the rest of the simulation.
 - * As in Game \mathbf{G}_{10} , If $d(\text{pw}_i, \text{pw}'_{1-i}) \leq \delta$ sets $y = 1$, and sets $y = 0$ otherwise.
 - * As in Game \mathbf{G}_{10} , uses the privacy simulator for the garbling scheme to generate $(F_i, X_i, d_i) \leftarrow \mathcal{S}(1^\lambda, f, y)$.
 - * As in Game \mathbf{G}_{10} , parses $(X_{i,i}, X_{i,1-i}) \leftarrow X_i$.
 - * As in Game \mathbf{G}_{10} , sends $X_{i,1-i}$ to \mathcal{P}_{1-i} as the OT output.
 - Plays an OT receiver honestly with pw'_i .
- \mathcal{S}_{RFE} follows the instructions in Figure 5 for steps 3-8.

Nothing could have changed from the point of view of \mathcal{Z} . If $d(\text{pw}_i, \text{pw}'_{1-i}) \leq \delta$, this game is identical to Game \mathbf{G}_{10} . If $d(\text{pw}_i, \text{pw}'_{1-i}) > \delta$, a random pw'_i is used instead of pw_i . However, pw'_i only affects the OT execution with \mathcal{P}_i as receiver, where \mathcal{P}_{1-i} does not receive any outputs anyway. In this case, \mathcal{P}_i 's output key gets set randomly by \mathcal{F} as of Game \mathbf{G}_9 , so that does not change.

In Figure 21, we show the simulator \mathcal{S}_{RFE} for Π_{RFE} .

- Upon receiving $(\text{NewSession}, \text{sid}, \mathcal{P}_i)$ from $\mathcal{F}_{\text{RFE}}^P$ for honest party \mathcal{P}_i when \mathcal{P}_{1-i} is also honest, \mathcal{S}_{RFE} generates F_i, X_{1-i} using the obliviousness simulator for the garbling scheme, and sends those to \mathcal{P}_{1-i} .
- Upon receiving $(\text{NewSession}, \text{sid}, \mathcal{P}_i)$ from $\mathcal{F}_{\text{RFE}}^P$ for honest party \mathcal{P}_i when \mathcal{P}_{1-i} is corrupt, \mathcal{S}_{RFE} does the following:
 - \mathcal{S}_{RFE} postpones step 1.
 - In step 2, \mathcal{S}_{RFE} :
 - * Plays an OT sender as follows:
 - Waits for \mathcal{P}_{1-i} to provide their select bits to the OT. As a result, \mathcal{S}_{RFE} learns the pass-string used by \mathcal{P}_{1-i} , pw'_{1-i} .
 - Makes a $(\text{TestPwd}, \text{sid}, \mathcal{P}_i)$ query to \mathcal{F} . If $d(\text{pw}_i, \text{pw}'_{1-i}) \leq \delta$ (that is, the adversary has approximately guessed \mathcal{P}_i 's pass-string), \mathcal{S}_{RFE} learns pw_i , and sets $\text{pw}'_i = \text{pw}_i$. Otherwise, it sets pw'_i at random such that $d(\text{pw}'_i, \text{pw}'_{1-i}) > \delta$, and uses pw'_i in place of pw_i in the rest of the simulation.
 - If $d(\text{pw}_i, \text{pw}'_{1-i}) \leq \delta$ sets $y = 1$, and sets $y = 0$ otherwise.
 - Uses the privacy simulator for the garbling scheme to generate $(F_i, X_i, d_i) \leftarrow S(1^\lambda, f, y)$.
 - Parses $(X_{i,i}, X_{i,1-i}) \leftarrow X_i$.
 - Sends $X_{i,1-i}$ to \mathcal{P}_{1-i} as the OT output.
 - * Plays an OT receiver honestly with pw'_i .
 - \mathcal{S}_{RFE} follows the instructions in Figure 5 for steps 3-8 with pw'_i .

Additionally, \mathcal{S}_{RFE} forwards all other instructions from \mathcal{Z} to \mathcal{A} and reports all output of \mathcal{A} towards \mathcal{Z} . Instructions of corrupting a player are only obeyed if they are received before the protocol started, i.e., before \mathcal{S} received any NewSession query from $\mathcal{F}_{\text{RFE}}^P$.

Fig. 21. Simulator \mathcal{S}_{RFE} for Π_{RFE}

D Proof that $s\mathcal{F}_{\text{RFE}}^P$ is Enough to Realize $\mathcal{F}_{\text{fPAKE}}^P$

In Figure 22, we describe a protocol $\text{fPAKE}_{\text{YGC}}$ which trivially realizes $\mathcal{F}_{\text{fPAKE}}^P$ in the $s\mathcal{F}_{\text{RFE}}^P$ -hybrid model.

Upon receiving the input $(\text{sid}, \text{pw}_i)$, $\mathcal{P}_i \in \{\mathcal{P}_0, \mathcal{P}_1\}$ does the following:

- Sends $(\text{Init}, \text{sid})$ to $s\mathcal{F}_{\text{RFE}}^P$;
- Sends $(\text{NewSession}, \text{sid}, \text{pw}_i)$ to $s\mathcal{F}_{\text{RFE}}^P$;
- Waits for k from $s\mathcal{F}_{\text{RFE}}^P$, and
- Outputs k .

Fig. 22. Protocol $\text{fPAKE}_{\text{YGC}}$ realizing $\mathcal{F}_{\text{fPAKE}}^P$ in the $s\mathcal{F}_{\text{RFE}}^P$ -hybrid model.

Theorem 11. *Protocol $\text{fPAKE}_{\text{YGC}}$ realizes $\mathcal{F}_{\text{fPAKE}}^P$ in the $s\mathcal{F}_{\text{RFE}}^P$ -hybrid model.*

Proof. For every efficient adversary \mathcal{A} , we describe a simulator $\mathcal{S}_{\text{fPAKE}}$ in Figure 23 such that no efficient environment can distinguish an execution with the real protocol $\text{fPAKE}_{\text{YGC}}$ and \mathcal{A} from an execution with the ideal functionality $\mathcal{F}_{\text{fPAKE}}^P$ and $\mathcal{S}_{\text{fPAKE}}$. Since the environment does not get any information about the honest parties except their output, all the simulator needs to do is respond to queries to $s\mathcal{F}_{\text{RFE}}^P$. Since the honest party does nothing except query the ideal functionality $s\mathcal{F}_{\text{RFE}}^P$, and its output gets replaced by values chosen by $\mathcal{F}_{\text{fPAKE}}^P$, there is nothing to simulate.

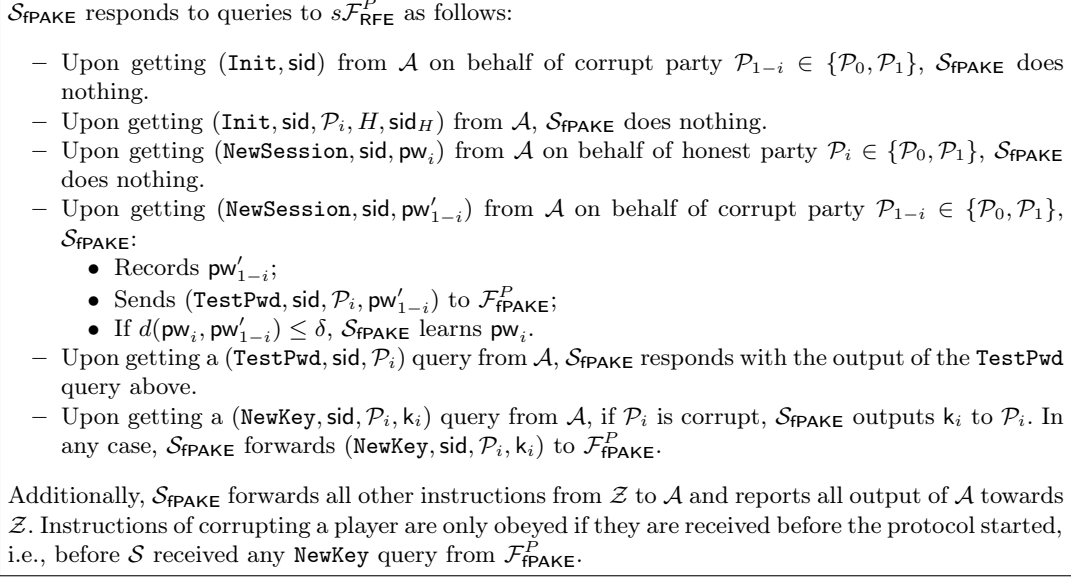


Fig. 23. Simulator $\mathcal{S}_{\text{fPAKE}}$ for $\text{fPAKE}_{\text{YGC}}$.

All that remains to show is that the values produced by $\mathcal{F}_{\text{fPAKE}}^P$ and by $s\mathcal{F}_{\text{RFE}}^P$ are identically distributed. We describe the outputs of $\mathcal{F}_{\text{fPAKE}}^P$ and $s\mathcal{F}_{\text{RFE}}^P$ in Figure 24. The table enumerates all possible cases in the functionalities. Cases in $s\mathcal{F}_{\text{RFE}}^P$ are described in terms of the distances between pass-strings: between the honest parties' pass-strings if no man in the middle attack occurred, and between adversarial and honest pass-strings if it did. (If the adversary engaged one party but not the other, only one of the distances is filled in.) Cases in $\mathcal{F}_{\text{fPAKE}}^P$ are described in terms of record markings ("fresh", "compromised" or "interrupted"). There is a one-to-one mapping between the cases in $s\mathcal{F}_{\text{RFE}}^P$ and $\mathcal{F}_{\text{fPAKE}}^P$ such that the outputs for the parties, whether they are honest or corrupt, are identically distributed. Those outputs are described in the last three columns of the table, as tuples of values the first of which is output to \mathcal{P}_0 , and the second of which is output to \mathcal{P}_1 . a, b are adversarially chosen values (which may or may not be distinct). r, s are independent, uniformly random values.

The transformation of Barak *et al.* proceeds in two steps. First, links are initialized:

1. Each party generates a signing and verification key pair, and sends the verification key to its partner.
2. Each party then signs the key it receives and sends the signature back.
3. Each party verifies the signature it receives on its own verification key using the verification key it received; if the signature does not verify, it aborts.

Second, the parties run the protocol exactly as they would over authenticated channels, signing each message with their signing key, and verifying each signature they receive.

$s\mathcal{F}_{\text{RFE}}^P$			$\mathcal{F}_{\text{fPAKE}}^P$		outputs in both $s\mathcal{F}_{\text{RFE}}^P$ and $\mathcal{F}_{\text{fPAKE}}^P$		
distance between \mathcal{P}_0 's pass-string and \mathcal{P}_1 's pass-string $d(\text{pw}_0, \text{pw}_1)$, if MITM didn't happen	distance between \mathcal{P}_0 's pass-string and the adversary's $d(\text{pw}_0, \text{pw}'_1)$, if MITM happened	distance between \mathcal{P}_1 's pass-string and the adversary's $d(\text{pw}'_0, \text{pw}_1)$, if MITM happened	\mathcal{P}_0 's record	\mathcal{P}_1 's record	\mathcal{P}_0 and \mathcal{P}_1 honest	\mathcal{P}_0 honest, \mathcal{P}_1 corrupt	\mathcal{P}_0 and \mathcal{P}_1 corrupt
close ($\leq \delta$)			fresh	fresh	r, r	a, b	a, b
far ($> \delta$)			fresh	fresh	r, s	r, b	a, b
	close ($\leq \delta$)	close ($\leq \delta$)	compromised	compromised	a, b	a, b	a, b
	close ($\leq \delta$)	far ($> \delta$)	compromised	interrupted	a, s	a, b	a, b
	close ($\leq \delta$)		compromised	fresh	a, s	a, b	a, b
	far ($> \delta$)	close ($\leq \delta$)	interrupted	compromised	r, b	r, b	a, b
	far ($> \delta$)	far ($> \delta$)	interrupted	interrupted	r, s	r, b	a, b
	far ($> \delta$)		interrupted	fresh	r, s	r, b	a, b
		close ($\leq \delta$)	fresh	compromised	r, b	r, b	a, b
		far ($> \delta$)	fresh	interrupted	r, s	r, b	a, b

Fig. 24. Output tables for $\mathcal{F}_{\text{fPAKE}}^P$ and $s\mathcal{F}_{\text{RFE}}^P$. r, s represent random outputs, while a, b represent adversarially chosen outputs.

Applying this transformation adds (1) two rounds of communication, and (2) a hash operation and a signature operation for each message, assuming the hash-and-sign paradigm is used.

E A Concrete OT

In this section, we recall a concrete UC-secure oblivious transfer protocol due to Chou and Orlandi [CO15]. While they consider the general case of 1-out-of- n transfer, we only consider $n = 2$. Say m 1-out-of-2 OTs are performed. The protocol requires the sender to compute $m + 2$ exponentiations, and the receiver to compute $2m$ exponentiations, for a total of $3m + 2$ exponentiations. Figure 25 shows a summary of the protocol. Note that this construction does require a random oracle.

F Proof of Theorem 6

We proceed in a series of games, starting with the real execution of the protocol and ending up with the ideal execution, with a simulator. To abbreviate notation, we skip all **role** tags used by $\mathcal{F}_{\text{fPAKE}}$ since they are not needed due to the symmetric layout of the protocol. For convenience, we refer to a query (**NewKey**, $\text{sid}, \mathcal{P}_i, \ell', \mathbf{k}_i$) from the adversary \mathcal{S} as *due* when:

- \mathcal{P}_i is honest

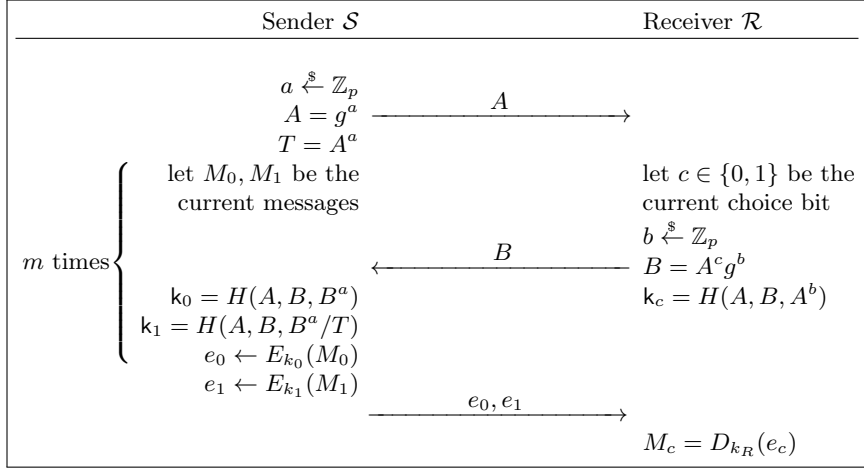


Fig. 25. A Concrete OT [CO15] to be used in Π_{RFE}

The functionality $\mathcal{F}_{\ell\text{-iPAKE}}$ is parameterized by a security parameter λ and makes use of two initially empty lists $\Lambda_{\mathcal{P}}$ and $\Lambda_{\mathcal{L}}$, storing pass-strings and labels, respectively. It interacts with an adversary \mathcal{S} and the (dummy) parties \mathcal{P}_0 and \mathcal{P}_1 via the following queries:

- **Upon receiving a query (NewSession, sid, pw_i, role, ℓ) from party \mathcal{P}_i :**
 - Send (NewSession, sid, \mathcal{P}_i , role, ℓ) to \mathcal{S} ;
 - If one of the following is true, record $(\mathcal{P}_i, \text{pw}_i)$ in $\Lambda_{\mathcal{P}}$ and mark this record **fresh**, and record (\mathcal{P}_i, ℓ) in $\Lambda_{\mathcal{L}}$ unless there already exists a record (\mathcal{P}_i, \cdot) in $\Lambda_{\mathcal{L}}$.
 - * This is the first NewSession query
 - * This is the second NewSession query and there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$
- **Upon receiving a query (TestPwd, sid, \mathcal{P}_i , pw'_i, ℓ') from \mathcal{S} :**

If there is a **fresh** record $(\mathcal{P}_i, \text{pw}_i)$ in $\Lambda_{\mathcal{P}}$, then do:

 - If $\text{pw}_i = \text{pw}'_i$, mark the record **compromised**; else mark it **interrupted**;
 - Record $(\mathcal{P}_{1-i}, \ell')$ in $\Lambda_{\mathcal{L}}$, possibly overwriting any existing record $(\mathcal{P}_{1-i}, \cdot)$.
- **Upon receiving a query (NewKey, sid, \mathcal{P}_i , sk) from \mathcal{S} , where $|\text{sk}| = \lambda$:**

If there is a record $(\mathcal{P}_{1-i}, \ell)$ in $\Lambda_{\mathcal{L}}$, extract ℓ from it; otherwise set $\ell \leftarrow \perp$.

If there is a record of the form $(\mathcal{P}_i, \text{pw}_i)$ in $\Lambda_{\mathcal{P}}$, and this is the first NewKey query for \mathcal{P}_i , then:

 - If at least one of the following is true, then output $(\text{sid}, \ell, \text{sk})$ to player \mathcal{P}_i :
 - * The record is **compromised**
 - * \mathcal{P}_i is corrupted
 - * The record is **fresh**, \mathcal{P}_{1-i} is corrupted, and there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ with $\text{pw}_{1-i} = \text{pw}_i$
 - If this record is **fresh**, both parties are honest, there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ with $\text{pw}_{1-i} = \text{pw}_i$, a key sk' was sent to \mathcal{P}_{1-i} , and $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ was **fresh** at the time, then output $(\text{sid}, \ell, \text{sk}')$ to \mathcal{P}_i ;
 - In any other case, pick a new random key sk' of length λ and send $(\text{sid}, \ell, \text{sk}')$ to \mathcal{P}_i .

No matter what, mark the record $(\mathcal{P}_i, \text{pw}_i)$ as **completed**.

Fig. 26. Functionality $\mathcal{F}_{\ell\text{-iPAKE}}$

The parties \mathcal{P}_0 and \mathcal{P}_1 are running with \mathcal{F}_{CRS} , \mathcal{F}_{RO} and \mathcal{F}_{IC} .

Protocol Steps:

1. When a party \mathcal{P}_i , $i \in \{0, 1\}$, receives an input $(\text{NewSession}, \text{sid}, \mathcal{P}_i, \text{pw}_i, \ell)$ from \mathcal{Z} , it does the following:
 - chooses $x \xleftarrow{\$} \mathbb{F}_q$
 - sends (sid, crs) to \mathcal{F}_{CRS} and receives $(\text{sid}, (g, q))$ back
 - sends $(\text{sid}, \mathcal{E}, \text{pw}_i || \ell, g^x)$ to \mathcal{F}_{IC} and receives (sid, X^*) back
 - sends (sid, ℓ, X^*) to \mathcal{P}_{1-i} and waits for an answer
2. When \mathcal{P}_i , who already obtained an input $(\text{NewSession}, \text{sid}, \mathcal{P}_i, \text{pw}_i, \ell)$ and thus holds $(x, (g, q), X^*)$, receives a message (sid, ℓ', Y^*) from \mathcal{P}_{1-i} , it
 - sends $(\text{sid}, \mathcal{D}, \text{pw}_i || \ell, Y^*)$ to \mathcal{F}_{IC} and receives (sid, X') back
 - sends $(\text{sid}, X^*, Y^*, X'^y)$ to \mathcal{F}_{RO} and receives (sid, k_i) back
 - outputs (sid, ℓ', k_i) towards \mathcal{Z} and terminates the session.

Fig. 27. A UC Execution of EKE2. We skip the **role** tags since they are not needed due to the symmetric layout of the protocol.

- there is a **fresh** record of the form $(\mathcal{P}_i, \text{pw}_i)$ in $\Lambda_{\mathcal{P}}$
- this is the first **NewKey** query for \mathcal{P}_i
- there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ in $\Lambda_{\mathcal{P}}$ with $\text{pw}_i = \text{pw}_{1-i}$ and \mathcal{P}_{1-i} is honest
- a key k_{1-i} was sent to the other party, and $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ was **fresh** at the time.

Game G_0 : The real protocol execution. This is the real execution where the environment \mathcal{Z} runs the EKE2 protocol (see Figure 27) with parties \mathcal{P}_0 and \mathcal{P}_1 , both having access to ideal CRS, RO, and IC functionalities, and an adversary \mathcal{A} that, w.l.o.g., is assumed to be the dummy adversary as shown in [Can01, section 4.4.1].

Game G_1 : Modeling the ideal layout. We first regroup and create new machines, similar to Game 1 in the proof of Theorem 7. The new machine \mathcal{S} executes the code of the CRS, RO and IC functionalities as depicted in Figures 13, 14, and 15.

Game G_2 : Simulating the ideal functionalities. We modify simulation of \mathcal{F}_{RO} and \mathcal{F}_{IC} as follows. We let \mathcal{S} implement Figure 15 by maintaining a list Λ_{IC} with entries of the form $(k, m, \alpha, \mathcal{E} | \mathcal{D}, c)$. \mathcal{S} handles encryption and decryption queries as follows:

- Upon receiving $(\text{sid}, \mathcal{E}, k, m)$ (for shortness of notation, we will also write $\mathcal{E}_k(m)$ for this query), if $k \notin \mathbb{F}_p$ or $m \notin \mathbb{G}$ then abort. Else, if there is an entry $(k, m, *, *, c)$ in Λ_{IC} , \mathcal{S} replies with (sid, c) . Else, \mathcal{S} chooses $c \xleftarrow{\$} \mathbb{G} \setminus \{1\}$. If there already is a record $(*, *, *, *, c)$ in Λ_{IC} , \mathcal{S} aborts. Else, \mathcal{S} adds $(k, m, \perp, \mathcal{E}, c)$ to Λ_{IC} and replies with (sid, c) .
- Upon receiving $(\text{sid}, \mathcal{D}, k, c)$ (or $\mathcal{D}_k(c)$, for short), if $k \notin \mathbb{F}_p$ or $c \notin \mathbb{G}$ then abort. Else, if there is an entry $(k, m, *, *, c)$ in Λ_{IC} , \mathcal{S} replies with (sid, m) . Else, \mathcal{S} chooses $\alpha \leftarrow \mathbb{F}_q^*$. If there already is a record $(*, *, g^\alpha, *, *, *)$ in Λ_{IC} , \mathcal{S} aborts. Else, \mathcal{S} adds $(k, g^\alpha, \alpha, \mathcal{D}, c)$ to Λ_{IC} and replies with (sid, g^α) .

Similarly, let Λ_{RO} denote the list that \mathcal{S} maintains upon implementing Figure 14, containing entries of the form (m, h) . We let \mathcal{S} handle queries to \mathcal{F}_{RO} as follows:

- Upon receiving $H(m)$, if $m \notin \{0, 1\}^* \times \{0, 1\}^* \times \mathbb{G}^3$, then abort. Else, if there is an entry (m, h) in Λ_{RO} , \mathcal{S} replies with (sid, h) . Else, \mathcal{S} chooses $h \xleftarrow{\$} \{0, 1\}^k$. If

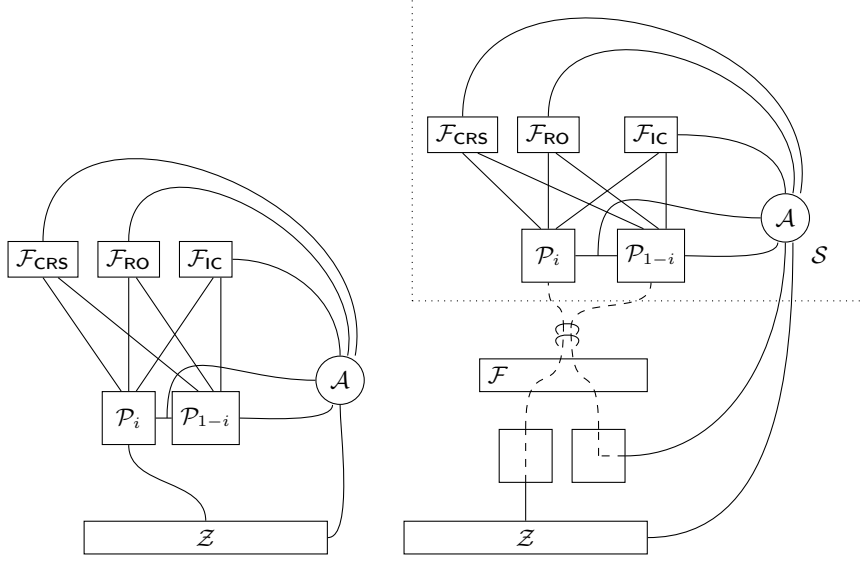


Fig. 28. Transition from game \mathbf{G}_0 (left) to game \mathbf{G}_1 (right), showing a setting where \mathcal{P}_{1-i} is corrupted.

there already is a record $(*, *, h)$ in Λ_{RO} , \mathcal{S} aborts. Else, \mathcal{S} adds (m, h) to Λ_{RO} and replies with (sid, h) .

These modifications later allow \mathcal{S} to extract unique inputs from values obtained from the two functionalities. Especially, note that Λ_{IC} will never contain $(*, k_i, *, *, \mathcal{E}, c)$, $(*, k_{1-i}, *, *, \mathcal{E}, c)$ with $k_i \neq k_{1-i}$. The entry α serves \mathcal{S} as a trapdoor for solving discrete-log type problems.

Since q is greater than 2^λ , if the oracles are only queried a polynomial number of times, the birthday problem states that game \mathbf{G}_1 and game \mathbf{G}_2 are indistinguishable with probability overwhelming in λ .

Game \mathbf{G}_3 : Building $\mathcal{F}_{\text{iPAKE}}$. In this game, we start modeling $\mathcal{F}_{\text{iPAKE}}$. First, we let \mathcal{F} maintain two initially empty lists: $\Lambda_{\mathcal{P}}$, a list of tuples of the form $(\mathcal{P}_i, \text{pw}_i)$ and $\Lambda_{\mathcal{L}}$, a list of tuples of the form (\mathcal{P}_i, ℓ) . Upon receiving a query $(\text{NewSession}, \text{sid}, \text{pw}_i, \ell)$ from (dummy) party \mathcal{P}_i , if this is the first **NewSession** query, or if this is the second **NewSession** query and there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i}, \ell')$, then \mathcal{F} records $(\mathcal{P}_i, \text{pw}_i)$ in $\Lambda_{\mathcal{P}}$ and marks this record as **fresh**. If $\Lambda_{\mathcal{L}}$ does not contain any record (\mathcal{P}_i, \cdot) so far, \mathcal{F} also records (\mathcal{P}_i, ℓ) in $\Lambda_{\mathcal{L}}$. Then, \mathcal{F} relays the query $(\text{NewSession}, \text{sid}, \text{pw}_i, \ell)$ to \mathcal{S} . Now that \mathcal{F} knows about pass-strings and labels, we can add a **TestPwd** interface to \mathcal{F} as described in Figure 8. We let \mathcal{S} parse outputs (sid, ℓ', k_i) towards \mathcal{F} to be of the form $(\text{NewKey}, \text{sid}, \mathcal{P}_i, \ell', k_i)$ by adding the **NewKey** tag and the name of the party who produced the output. Additionally, we let \mathcal{F} translate this back to (sid, ℓ', k_i) and send it to \mathcal{Z} via the dummy party \mathcal{P}_i , marking the corresponding record as **completed**.

None of these modifications changes the output towards \mathcal{Z} compared to the previous game \mathbf{G}_2 .

Game \mathbf{G}_4 : \mathcal{F} generates a random session key for an honest, interrupted session. Upon receiving a query ($\text{NewKey}, \text{sid}, \mathcal{P}_i, k_i$) from \mathcal{S} , if \mathcal{P}_i is not corrupted and there is a record of the form $(\mathcal{P}_i, \text{pw}_i)$ that is marked as **interrupted**, and this is the first NewKey query for \mathcal{P}_i , we let \mathcal{F} choose a random session key k^* of length λ . Additionally, \mathcal{F} derives the label as follows: if there is a record $(\mathcal{P}_{1-i}, \ell^*)$ in $\Lambda_{\mathcal{L}}$, extract ℓ^* from it; otherwise, set $\ell^* \leftarrow \perp$. Then, \mathcal{F} outputs $(\text{sid}, \ell^*, k^*)$ to \mathcal{P} . If there is no such interrupted record, \mathcal{F} continues to relay k_i and ℓ' .

Since the simulators described in game \mathbf{G}_3 and game \mathbf{G}_4 do not make use of the TestPw interface, none of the records of \mathcal{F} are marked as **interrupted** and thus the output towards \mathcal{Z} is equally distributed in both games.

Game \mathbf{G}_5 : \mathcal{S} handles dictionary attacks against the client \mathcal{P}_0 using the TestPw interface. In this game, we will only change the simulation. First note that the client, the initiator of the protocol, is intended to send the first message, and we call him \mathcal{P}_0 . If both \mathcal{P}_0 and \mathcal{P}_1 are honest, \mathcal{P}_0 obtained input and \mathcal{Z} advises \mathcal{A} to substitute (sid, ℓ', Y^*) with $(\text{sid}, \ell_{\mathcal{Z}}, Y_{\mathcal{Z}}^*)$, or if \mathcal{P}_1 is corrupted and produces $(\text{sid}, \ell_{\mathcal{Z}}, Y_{\mathcal{Z}}^*)$ as first flow, then \mathcal{S} will proceed simulation of \mathcal{P}_0 using $\ell_{\mathcal{Z}}$ and $Y_{\mathcal{Z}}^*$. In this situation, we modify \mathcal{S} as follows: upon receiving $(\text{sid}, \ell_{\mathcal{Z}}, Y_{\mathcal{Z}}^*)$, if there is an entry $(\text{pw}_{\mathcal{Z}} || \hat{\ell}, *, *, \mathcal{E}, Y_{\mathcal{Z}}^*)$ for any $\hat{\ell} \in \mathcal{L}$ in Λ_{IC} ⁸, the simulator asks a TestPw query $(\text{TestPw}, \text{sid}, \mathcal{P}_0, \text{pw}_{\mathcal{Z}}, \ell_{\mathcal{Z}})$ to \mathcal{F} . \mathcal{S} then proceeds the simulation using $\text{pw}_{\mathcal{Z}}$ and $\ell_{\mathcal{Z}}$ instead of pw_0 and ℓ' ⁹. If there is no entry $(\text{pw}_{\mathcal{Z}} || \hat{\ell}, *, *, \mathcal{E}, Y_{\mathcal{Z}}^*)$ in Λ_{IC} , \mathcal{S} sends $(\text{TestPw}, \text{sid}, \mathcal{P}_0, \text{pw}_0, \ell_{\mathcal{Z}})$ to \mathcal{F} ¹⁰.

Regarding the label, observe that \mathcal{S} 's NewKey query will contain $\ell_{\mathcal{Z}}$ which was contained in the output of the honest \mathcal{P}_0 (cf. Figure 27). Since TestPw queries overwrite any existing labels, there will be an entry $(\mathcal{P}_1, \ell_{\mathcal{Z}})$ in $\Lambda_{\mathcal{L}}$ and thus, regarding the label, the output towards \mathcal{Z} does not change compared to the previous game. Regarding the session key, we have to analyze different cases depending on whether $Y_{\mathcal{Z}}^*$ was generated using \mathcal{F}_{IC} or not. However, observe that the only changes of session keys between this and the previous game occur whenever a TestPw query of \mathcal{S} causes a record to be marked as **interrupted**.

- There is an entry $(\text{pw}_{\mathcal{Z}} || \hat{\ell}, *, *, \mathcal{E}, Y_{\mathcal{Z}}^*)$ in Λ_{IC} : if $\text{pw}_{\mathcal{Z}} = \text{pw}_0$, the record is marked **compromised** and the session key is not changed by \mathcal{F} . If $\text{pw}_{\mathcal{Z}} \neq \text{pw}_0$, on the other hand, the record is marked **interrupted** and \mathcal{F} hands out a random session key, as opposed to game \mathbf{G}_4 . However, since the session key is distributed as before, \mathcal{Z} can only detect this by reproducing \mathcal{P}_0 's input $(\text{sid}, X^*, Y_{\mathcal{Z}}^*, \text{CDH}(\mathcal{D}_{\text{pw}_0 || \ell}(X^*), \mathcal{D}_{\text{pw}_0 || \ell_{\mathcal{Z}}}(Y_{\mathcal{Z}}^*)))$ to \mathcal{F}_{RO} . Lemma 12 (see below) shows indistinguishability of game \mathbf{G}_4 and game \mathbf{G}_5 .

⁸ This entry is unique due to simulation of \mathcal{F}_{IC} as described in game \mathbf{G}_2 .

⁹ Note that, since \mathcal{F} does not leak any information at this point, \mathcal{S} cannot depend on the outcome of a TestPw query.

¹⁰ Letting \mathcal{S} guess a pass-string that he actually knows seems a little artificial. Indeed, the simulation in this case will be changed in game \mathbf{G}_{12} when \mathcal{S} becomes oblivious of \mathcal{P}_0 's pass-string.

- There is no entry $(*, *, *, \mathcal{E}, Y_{\mathcal{Z}}^*)$ in Λ_{IC} : since the **TestPwd** query will result in a compromised record, the modified simulation has no impact on the output towards \mathcal{Z} in this case.

The following lemma bounds the probability that an unsuccessful dictionary attack leads to a non-random looking session key. Since in this case the labels do not play any role (the encryption keys of the form $pass - string || label$ will not match regardless of the labels), we ignore them for the sake of simplicity.

Lemma 12. *If CDH holds in \mathbb{G} , then $\forall \text{pw}_0, Y_{\mathcal{Z}}^* \leftarrow \mathcal{Z}$, where $Y_{\mathcal{Z}}^*$ is a ciphertext generated through \mathcal{F}_{IC} with some key $\text{pw}_{\mathcal{Z}} \neq \text{pw}_0$, $\text{pw}_0 \in \mathbb{F}_p$, it holds that*

$$\Pr_{\mathbf{G}_5}[\text{CDH}(\mathcal{D}_{\text{pw}_0}(X^*), \mathcal{D}_{\text{pw}_0}(Y_{\mathcal{Z}}^*)) \leftarrow \mathcal{Z}(X^*)] = \text{negl}(\lambda).$$

Proof. We create an attacker \mathcal{B}_{CDH} given a CDH instance $(g, A = g^a, B = g^b)$. \mathcal{B}_{CDH} runs \mathcal{Z} simulating game \mathbf{G}_5 as follows: \mathcal{B}_{CDH} internally runs all of the participating machines, i.e. \mathcal{S} , \mathcal{F} and the dummy parties as in game \mathbf{G}_5 , but with some modifications. First, \mathcal{B}_{CDH} computes $X^* \leftarrow \mathcal{E}_{\text{pw}_0}(A)$ and updates Λ_{IC} accordingly, aborting if there already was an entry $(\text{pw}_0, A, *, \mathcal{E}, *)$. Upon receiving a query $\mathcal{D}_{\text{pw}_0}(Y_{\mathcal{Z}}^*)$, \mathcal{B}_{CDH} again aborts if there already is an entry $(\text{pw}_0, *, *, \mathcal{E}, Y_{\mathcal{Z}}^*)$. Otherwise, it draws $\beta \xleftarrow{\$} \mathbb{F}_q$ and sets the answer to this query to be Bg^β . This can happen multiple times (for different $Y_{\mathcal{Z}}^*$), and \mathcal{B}_{CDH} keeps track of the pairs $(\beta, Y_{\mathcal{Z}}^*)$ in a list Λ_{CDH} . The last modification concerns the part of the simulator's code of game \mathbf{G}_5 where a value $Z \leftarrow \mathcal{D}_{\text{pw}_0}(Y^*)^a$ needs to be computed, but note that \mathcal{B}_{CDH} does not know a . Instead, \mathcal{B}_{CDH} just sets $Z \leftarrow \perp$.

Finally, \mathcal{B}_{CDH} picks a random entry from Λ_{RO} asked by the environment, parses it as $((\text{pw}_0, X^*, Y_{\mathcal{Z}}^*, Z), h)$, looks for an entry $(\beta, Y_{\mathcal{Z}}^*)$ in Λ_{CDH} and outputs $Z/(g^a)^\beta$ as a CDH solution.

First, note that if \mathcal{B}_{CDH} does not abort, it perfectly emulates \mathcal{Z} 's view in game \mathbf{G}_5 , since A, Bg^β are random in \mathbb{G} and the record for \mathcal{P}_0 will be interrupted, which means that \mathcal{F} will output a random session key for \mathcal{P}_0 , overwriting \perp . Second, \mathcal{B}_{CDH} only has to abort if there is a collision upon choosing random values from \mathbb{G} .

Assume that \mathcal{Z} outputs $\text{CDH}(\mathcal{D}_{\text{pw}_0}(X^*), \mathcal{D}_{\text{pw}_0}(Y_{\mathcal{Z}}^*))$ with non-negligible probability. This is only possible if \mathcal{Z} asked both corresponding decryption queries. Existence of $(\text{pw}_{\mathcal{Z}}, *, *, \mathcal{E}, Y_{\mathcal{Z}}^*)$ with $\text{pw}_{\mathcal{Z}} \neq \text{pw}_0$ in Λ_{IC} ensures that the answer to $\mathcal{D}_{\text{pw}_0}(Y_{\mathcal{Z}}^*)$ can be chosen by \mathcal{B}_{CDH} as described above. Thus, \mathcal{B}_{CDH} finds a correct CDH solution with non-negligible probability $1/q_{\mathcal{Z}}$, where $q_{\mathcal{Z}}$ is the number of hash queries issued by \mathcal{Z} .

Game \mathbf{G}_6 : \mathcal{S} handles dictionary attacks against the server \mathcal{P}_1 using the **TestPwd** interface. Again, in this game we only change the simulation. Analogously to game \mathbf{G}_5 , we let \mathcal{S} use the **TestPwd** interface upon receiving adversarially generated $X_{\mathcal{Z}}^*, \ell_{\mathcal{Z}}$ upon simulating \mathcal{P}_1 . Observe that the only difference is due to the order of flows: if \mathcal{S} extracts an incorrect pass-string, he produces Y^* using this wrong pass-string. However, Y^* will be distributed as before and again, \mathcal{Z} can only detect the

change by reproducing \mathcal{P}_1 's input to \mathcal{F}_{RO} , namely $(\text{sid}, X_{\mathcal{Z}}^*, Y^*, \text{CDH}(\mathcal{D}_{\text{pw}_1 || \ell_{\mathcal{Z}}}(X_{\mathcal{Z}}^*), \mathcal{D}_{\text{pw}_1 || \ell'}(Y^*)))$.

Using an analogous argument to Lemma 12, indistinguishability from game \mathbf{G}_5 follows from the hardness of CDH in \mathbb{G} .

Game \mathbf{G}_7 : \mathcal{F} aligns session keys. Upon receiving a query ($\text{NewKey}, \text{sid}, \mathcal{P}_i, \ell', k_i$) from \mathcal{S} for a session, if this query is *due* then output $(\text{sid}, \ell^*, k^*)$ to \mathcal{P}_i where k^* is the session key that was formerly sent to the other party and the label ℓ^* is derived as usual: if there is a record $(\mathcal{P}_{1-i}, \ell^*)$ in $\Lambda_{\mathcal{L}}$, extract ℓ^* from it; otherwise, set $\ell^* \leftarrow \perp$.

We now analyze distinguishability of this game from game \mathbf{G}_6 . If \mathcal{Z} tampered with the transcript, any player that received a modified message will not have a fresh record anymore (cf. simulation described in games \mathbf{G}_5 and \mathbf{G}_6) and the output of this player towards \mathcal{Z} is not changed in this game. On the other hand, if \mathcal{Z} does not advise \mathcal{A} to tamper with any message, \mathcal{F} did not overwrite any labels and thus $\ell^* = \ell'$. Additionally, perfect correctness of the EKE2 protocol ensures that, in case of a due record, $k_i = k^*$.

Note that \mathcal{F} still differs from the functionality $\mathcal{F}_{\ell\text{-iPAKE}}$ described in Figure 8 in some aspects. First, it does not output randomly generated session keys towards \mathcal{Z} for honest sessions. Furthermore, it reports all pass-strings to \mathcal{S} . We will take care of these remaining differences in the next games.

Game \mathbf{G}_8 : In some cases, \mathcal{F} generates a random session key when the other party is corrupted. Upon receiving a NewKey query ($\text{NewKey}, \text{sid}, \mathcal{P}_i, \ell', k_i$) from \mathcal{S} , if there is a fresh record of the form $(\mathcal{P}_i, \text{pw}_i)$ in $\Lambda_{\mathcal{P}}$, and this is the first NewKey query for \mathcal{P}_i , \mathcal{P}_i is honest and \mathcal{P}_{1-i} corrupted and there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ in $\Lambda_{\mathcal{P}}$ with $\text{pw}_i \neq \text{pw}_{1-i}$, we let \mathcal{F} pick a new random key k^* of length λ and send $(\text{sid}, \ell^*, k^*)$ to \mathcal{P}_i , where ℓ^* , as usual, is taken from the list $\Lambda_{\mathcal{L}}$ or set to be \perp .

The simulation ensures that the record $(\mathcal{P}_i, \text{pw}_i)$ is either compromised or interrupted (cf. description of the simulator in games \mathbf{G}_5 and \mathbf{G}_6). Thus, the modification has no effect since it only concerns fresh records.

Game \mathbf{G}_9 : \mathcal{F} generates a random session key for an honest session. Upon receiving a NewKey query ($\text{NewKey}, \text{sid}, \mathcal{P}_i, \ell', k_i$) from \mathcal{S} , if there is a fresh record of the form $(\mathcal{P}_i, \text{pw}_i)$ in $\Lambda_{\mathcal{P}}$, and this is the first NewKey query for \mathcal{P}_i , both parties are honest and the NewKey query is not *due*, we let \mathcal{F} pick a new random key k^* of length k and send $(\text{sid}, \ell^*, k^*)$ to \mathcal{P}_i , where ℓ^* , as usual, is taken from the list $\Lambda_{\mathcal{L}}$ or set to be \perp .

In other words, \mathcal{F} now generates a random session key upon a first NewKey query for an honest party \mathcal{P}_i with fresh record $(\mathcal{P}_i, \text{pw}_i)$ where the other party is also honest, if (at least) one of the following events happens:

1. There is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ in $\Lambda_{\mathcal{P}}$ with $\text{pw}_i \neq \text{pw}_{1-i}$;
2. No output was sent to the other party yet;
3. If there was output to the other party, the record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ in $\Lambda_{\mathcal{P}}$ was not fresh and thus interrupted or compromised at that time

In all of these cases, \mathcal{S} chose a fresh k_i following a uniform distribution and ℓ' was contained in the NewSession query of \mathcal{P}_i 's partner, thus $\ell' = \ell^*$. Regarding the session key, \mathcal{Z} can only notice a difference if it reproduces k_i by computing \mathcal{P}_i 's

input $(\text{sid}, X^*, Y^*, \text{CDH}(\mathcal{D}_{\text{pw}_i||\ell}(X^*), \mathcal{D}_{\text{pw}_i||\ell'}(Y^*)))$ to \mathcal{F}_{RO} and sending it to \mathcal{F}_{RO} via the adversary \mathcal{A} .

The following lemma bounds the probability that a session key of an unattacked session does not look random.

Lemma 13. *If CDH holds in \mathbb{G} , then $\forall \text{pw}, \ell, \ell' \leftarrow \mathcal{Z}$ with $\text{pw} \in \mathbb{F}_p$ and $\ell, \ell' \in \mathcal{L}$ it holds that*

$$\Pr_{\mathbf{G}_9}[\text{CDH}(\mathcal{D}_{\text{pw}||\ell}(X^*), \mathcal{D}_{\text{pw}||\ell'}(Y^*)) \leftarrow \mathcal{Z}(X^*, Y^*)] = \text{negl}(\lambda).$$

Proof. We only sketch the proof since it is similar to the proof of Lemma 12. Namely, the strategy of embedding (randomized versions of) a CDH challenge into the simulation of game \mathbf{G}_9 is done by just encrypting both CDH challenge elements to obtain X^* and Y^* . For the final argument, note that \perp is not seen by \mathcal{Z} since it is either replaced using a random session key or a previously computed key.

It follows that game \mathbf{G}_8 and game \mathbf{G}_9 are indistinguishable.

Game \mathbf{G}_{10} : \mathcal{F} always takes all labels from the list $\Lambda_{\mathcal{L}}$. We modify \mathcal{F} as follows: if \mathcal{F} outputs (sid, ℓ', k_i) towards \mathcal{P}_i where ℓ', k_i are taken from a query $(\text{NewKey}, \text{sid}, \mathcal{P}_i, \ell', k_i)$ from \mathcal{S} , \mathcal{F} extracts ℓ^* from a record $(\mathcal{P}_{1-i}, \ell^*)$ in $\Lambda_{\mathcal{L}}$ or sets $\ell^* \leftarrow \perp$ if such a record does not exist. \mathcal{F} then outputs $(\text{sid}, \ell^*, k_i)$ towards \mathcal{P}_i . We additionally modify \mathcal{S} to remove the labels from the **NewKey** queries altogether.

First observe that we can remove the labels from the **NewKey** queries because, in this and the past games, we ensured that \mathcal{F} now does not access this label anymore. However, we still have to argue indistinguishability of the current and the previous game. The cases where k_i of \mathcal{S} is relayed by \mathcal{F} towards \mathcal{P}_i are the following:

- \mathcal{P}_i has a compromised record
- \mathcal{P}_i is corrupted
- \mathcal{P}_i has a fresh record, its partner is corrupted and has a record with a matching pass-string

In the first case, we have that $\ell' = \ell^*$ since the label ℓ' outputted by \mathcal{P}_i was also contained in a **TestPwd** query by \mathcal{S} and overwrote any existing label send by \mathcal{P}_i 's partner. For the second case, observe that since we restrict to static corruption, corrupted players will not have records in $\Lambda_{\mathcal{P}}$ and thus this case will never happen. In the third case, corruption of the partner ensures that \mathcal{S} issued a **TestPwd** query which overwrote any existing label with ℓ' , so $\ell' = \ell^*$ as well.

Observe that now \mathcal{F} acts like $\mathcal{F}_{\ell\text{-iPAKE}}$ regarding the output of session keys and labels. The only remaining difference is that the **NewSession** queries still contain the pass-strings of the parties. In the next games, we will make the simulation independent of these pass-strings.

Game \mathbf{G}_{11} : **Simulate without pw_1 if server \mathcal{P}_1 is honest.** In case of receiving a $(\text{NewSession}, \text{sid}, \text{pw}_1, \ell')$ from an honest \mathcal{P}_1 playing the role of a server, we modify \mathcal{F} by forwarding only $(\text{NewSession}, \text{sid}, \mathcal{P}_1, \ell')$ to \mathcal{S} . We now have to modify \mathcal{S} to proceed simulation without knowing pw_1 . Upon receiving $(\text{NewSession}, \text{sid}, \mathcal{P}_1, \ell')$

from \mathcal{F} for an honest \mathcal{P}_1 , we let \mathcal{S} draw uniformly at random a “dummy” pass-string $\text{pw}_{\mathcal{S}}$ and proceed the simulation of \mathcal{P}_1 using $\text{pw}_{\mathcal{S}}$ as a pass-string.

We first note that, if at any time \mathcal{S} sends a **NewKey** query to \mathcal{F} containing a session key k_1 for \mathcal{P}_1 , this session key is only seen by \mathcal{Z} if \mathcal{P}_1 ’s record is compromised. Otherwise, we thus only have to argue indistinguishability of the transcripts of game \mathbf{G}_{10} and game \mathbf{G}_{11} .

- \mathcal{Z} sends $\ell_{\mathcal{Z}}, X_{\mathcal{Z}}^*$, there is a record $(\text{pw}_{\mathcal{Z}} || \hat{\ell}, X, *, \mathcal{E}, X^*)$ in Λ_{IC} for some $\hat{\ell} \in \mathcal{L}$ and $\text{pw}_{\mathcal{Z}} = \text{pw}_1$: since \mathcal{S} will issue a **TestPw** query that will result in a compromised record (cf. simulation described in game \mathbf{G}_6), nothing is changed since $\text{pw}_{\mathcal{S}}$ was never used, and Y^* is generated using the correct pass-string $\text{pw}_{\mathcal{Z}}$.
- \mathcal{Z} sends $\ell_{\mathcal{Z}}, X_{\mathcal{Z}}^*$, there is a record $(\text{pw}_{\mathcal{Z}} || \hat{\ell}, X, *, \mathcal{E}, X^*)$ in Λ_{IC} for some $\hat{\ell} \in \mathcal{L}$ and $\text{pw}_{\mathcal{Z}} \neq \text{pw}_1$: since \mathcal{P}_1 will receive a random session key from \mathcal{F} in this case (the record will be marked **interrupted**), we only have to argue indistinguishability of Y^* generated with $\text{pw}_{\mathcal{Z}} || \ell'$ instead of $\text{pw}_1 || \ell'$. Simulation of \mathcal{F}_{IC} ensures that Y^* is distributed uniformly random as before. Observe that here it is crucial that even for a corrupted session, an interrupted record lets the functionality hand out a random session key, since \mathcal{S} has no means to decide whether it has to output a session key for \mathcal{P}_1 that matches the session key that \mathcal{Z} can compute on behalf of \mathcal{P}_0 from the message (ℓ', Y^*) .
- \mathcal{Z} sends $\ell_{\mathcal{Z}}, X_{\mathcal{Z}}^*$ and no \mathcal{E} record: the simulation described in game \mathbf{G}_6 tells \mathcal{S} to issue a **TestPw** query, but now using $\text{pw}_{\mathcal{S}}$ instead of pw_1 . If, coincidentally, $\text{pw}_1 = \text{pw}_{\mathcal{S}}$, nothing changes. On the other hand, if $\text{pw}_1 \neq \text{pw}_{\mathcal{S}}$, \mathcal{P}_1 obtains a random session key from \mathcal{F} as opposed to the game before and Y^* is created using $\text{pw}_{\mathcal{S}} || \ell'$ instead of $\text{pw}_1 || \ell'$. This can only be detected if \mathcal{Z} reproduces \mathcal{P}_1 ’s input to \mathcal{F}_{RO} from game \mathbf{G}_{10} , which happens only with negligible probability according to Lemma 14 (see below).
- both parties honest and no injections: \mathcal{P}_1 will obtain a uniformly random session key from \mathcal{F} in this case, and thus the only difference is that Y^* was created using $\text{pw}_{\mathcal{S}} || \ell'$ instead of $\text{pw}_1 || \ell'$. Again, this is indistinguishable since Y^* is distributed exactly as before.

The following lemma bounds the probability that an injected X^* that was not obtained using encryption leads to a non-random looking session key.

Lemma 14. *If CDH holds in \mathbb{G} , then $\forall \text{pw}_1, \ell', \ell_{\mathcal{Z}}, X_{\mathcal{Z}}^* \leftarrow \mathcal{Z}$, where $X_{\mathcal{Z}}^*$ was not generated using \mathcal{F}_{IC} , $\ell', \ell_{\mathcal{Z}} \in \mathcal{L}$ and $\text{pw}_1 \in \mathbb{F}_p$, it holds that*

$$\Pr_{\mathbf{G}_{11}} [\text{CDH}(\mathcal{D}_{\text{pw}_1 || \ell_{\mathcal{Z}}}(X_{\mathcal{Z}}^*), \mathcal{D}_{\text{pw}_1 || \ell'}(Y^*)) \leftarrow \mathcal{Z}(Y^*)] = \text{negl}(\lambda).$$

Proof. Note that the only difference to Lemma 12 is that this time, no record $(*, *, *, \mathcal{E}, X_{\mathcal{Z}}^*)$ exists so the fact that \mathcal{B}_{CDH} is able to embed an element of its CDH challenge into $\mathcal{D}_{\text{pw}_1}(X_{\mathcal{Z}}^*)$ is even more obvious. The rest of the proof is analogously to Lemma 12.

Game \mathbf{G}_{12} : Simulate without pw_0 if client \mathcal{P}_0 is honest. In a similar fashion, we now let \mathcal{F} cut the pass-string from **NewSession** queries to an honest \mathcal{P}_0 . We

again have to modify \mathcal{S} to proceed simulation without knowing pw_0 . Upon receiving $(\text{NewSession}, \text{sid}, \mathcal{P}_0, \ell)$ from \mathcal{F} for an honest \mathcal{P}_0 , we let \mathcal{S} draw uniformly at random a “dummy” pass-string $\text{pw}_\mathcal{S}$. \mathcal{S} proceeds the simulation of \mathcal{P}_0 using $\text{pw}_\mathcal{S}$ as a pass-string.

Additionally, we further change \mathcal{S} in case of a dictionary attack against client \mathcal{P}_0 , i.e., upon receiving $\ell_\mathcal{Z}, Y_\mathcal{Z}^*$ from \mathcal{Z} . After submitting a **TestPw** query with an extracted $\text{pw}_\mathcal{Z}$, we let \mathcal{S} now choose $x' \xleftarrow{\$} \mathbb{F}_P$ and add $(\text{pw}_\mathcal{Z} || \ell, g^{x'}, x', \perp, X^*)$ to Λ_{IC} and proceed the simulation of \mathcal{P}_0 using x' instead of x .

- \mathcal{Z} sends $\ell_\mathcal{Z}, Y_\mathcal{Z}^*$, there is a record $(\text{pw}_\mathcal{Z} || \hat{\ell}, Y, *, \mathcal{E}, Y^*)$ in Λ_{IC} for some $\hat{\ell} \in \mathcal{L}$ and $\text{pw}_\mathcal{Z} = \text{pw}_0$: \mathcal{S} will issue a **TestPw** query that will result in a compromised record (cf. simulation described in game **G₅**), resulting in a session key that is computed using $\text{pw}_\mathcal{Z}$ instead of $\text{pw}_\mathcal{S}$. Additionally, X^* is generated using the incorrect pass-string $\text{pw}_\mathcal{S}$. However, adjusting Λ_{IC} as described above still allows \mathcal{S} to know the exponent of $\mathcal{D}_{\text{pw}_\mathcal{Z} || \ell}(X^*)$ and continue the simulation, making it look like $\text{pw}_\mathcal{Z}$ was used from the beginning. \mathcal{Z} ’s view is distributed exactly as before since x', x are both uniformly random in \mathbb{F}_P .
- \mathcal{Z} sends $\ell_\mathcal{Z}, Y_\mathcal{Z}^*$, there is a record $(\text{pw}_\mathcal{Z} || \hat{\ell}, Y, *, \mathcal{E}, Y^*)$ in Λ_{IC} for some $\hat{\ell} \in \mathcal{L}$ and $\text{pw}_\mathcal{Z} \neq \text{pw}_0$: since \mathcal{P}_0 will receive a random session key from \mathcal{F} in this case (the record will be interrupted), we only have to argue indistinguishability of X^* generated with $\text{pw}_\mathcal{S} || \ell$ instead of $\text{pw}_0 || \ell$. Simulation of \mathcal{F}_{IC} ensures that Y^* is distributed uniformly random as before. Again, it is crucial here that \mathcal{F} helps \mathcal{S} by randomizing the session key if needed.
- \mathcal{Z} sends $\ell_\mathcal{Z}, Y_\mathcal{Z}^*$ and no \mathcal{E} record: the simulation described in game **G₅** tells \mathcal{S} to issue a **TestPw** query, but now using $\text{pw}_\mathcal{S} || \ell$ instead of $\text{pw}_0 || \ell$. If, coincidentally, $\text{pw}_0 = \text{pw}_\mathcal{S}$, nothing changes. On the other hand, if $\text{pw}_0 \neq \text{pw}_\mathcal{S}$, \mathcal{P}_0 obtains a random session key from \mathcal{F} as opposed to the game before and X^* was created using $\text{pw}_\mathcal{S} || \ell$ instead of $\text{pw}_0 || \ell$. This can only be detected if \mathcal{Z} reproduces \mathcal{P}_0 ’s input to \mathcal{F}_{RO} from game **G₁₁**, which happens only with negligible probability using an argument very similar to Lemma 14.
- both parties honest and no injections: \mathcal{P}_0 will obtain a uniformly random session key from \mathcal{F} in this case, and thus the only difference is that X^* was created using $\text{pw}_\mathcal{S} || \ell$ instead of $\text{pw}_0 || \ell$. Again, this is indistinguishable since X^* is distributed exactly as before.

Observe that in game **G₁₂**, $\mathcal{F} = \mathcal{F}_{\ell\text{-iPAKE}}^{11}$, and thus the theorem follows. The complete description of the simulator of game **G₁₂** interacting with $\mathcal{F}_{\ell\text{-iPAKE}}$ and \mathcal{Z} is given in Figure 29.

¹¹ We note that we can, w.l.o.g, assume that there are no **NewSession** queries from \mathcal{Z} to corrupted parties. Thus, it is enough to remove the pass-strings from the **NewSession** queries given as input from \mathcal{Z} to honest parties.

The simulator \mathcal{S} , initialized with a security parameter λ , first runs a group generation algorithm using λ to obtain a cyclic group \mathbb{G} with generator g of order q with $\log_2(q) \geq \lambda$. Then, \mathcal{S} initializes the dummy adversary \mathcal{A} . \mathcal{S} then interacts with an ideal functionality $\mathcal{F}_{\ell\text{-iPAKE}}$ and an environment \mathcal{Z} via the following queries:

- **Upon receiving a query ($\text{NewSession}, \text{sid}, \mathcal{P}_i, \ell$) from $\mathcal{F}_{\ell\text{-iPAKE}}$:**
 - initialize a party \mathcal{P}_i , connect it to \mathcal{A} and proceed the UC protocol execution described in Figure 27 using $\text{pw}_{\mathcal{S}} \xleftarrow{\$} \mathbb{F}_p$ as pass-string and \mathcal{S} 's random coins. (Cf. games $\mathbf{G}_0, \mathbf{G}_{11}$ and \mathbf{G}_{12} .)
- **Upon receiving a query (sid, m) from any entity:** (Cf. games \mathbf{G}_1 and \mathbf{G}_2 .)
 If $m \notin \langle g \rangle^3$, then abort. Else:
 - if there is an entry (m, h) then reply with (sid, h) .
 - else, choose $h \xleftarrow{\$} \mathbb{F}_q$ and abort if there already is an entry $(*, *, h)$ in Λ_{RO} . Else, reply with (sid, h) .
- **If an internally simulated party \mathcal{P}_i produces an output ($\text{sid}, \ell', \mathbf{k}_i$):**
 Send $(\text{NewKey}, \text{sid}, \mathcal{P}_i, \mathbf{k}_i)$ to $\mathcal{F}_{\ell\text{-iPAKE}}$. (Cf. game \mathbf{G}_1 .)
- **If \mathcal{Z} sends ($\text{sid}, \ell_{\mathcal{Z}}, Z_{\mathcal{Z}}^*$) to an honest party \mathcal{P}_i :**
 - if $(\text{pw}_{\mathcal{Z}} || \ell, *, *, \mathcal{E}, Z_{\mathcal{Z}}^*) \in \Lambda_{IC}$ for any $\ell \in \mathcal{L}$, then send $(\text{TestPwd}, \text{sid}, \mathcal{P}_i, \text{pw}_{\mathcal{Z}}, \ell_{\mathcal{Z}})$ to $\mathcal{F}_{\ell\text{-iPAKE}}$ and proceed the simulation of \mathcal{P}_i with $\text{pw}_{\mathcal{Z}}$. (Cf. games \mathbf{G}_5 and \mathbf{G}_6 .)
 - if $(*, *, *, \mathcal{E}, Z_{\mathcal{Z}}^*) \notin \Lambda_{IC}$, then send $(\text{TestPwd}, \text{sid}, \mathcal{P}_i, \text{pw}_{\mathcal{S}}, \ell_{\mathcal{Z}})$ to $\mathcal{F}_{\ell\text{-iPAKE}}$. (Cf. games \mathbf{G}_5 and \mathbf{G}_6 .)
 - if \mathcal{P}_i was started with input $(\text{NewSession}, \text{sid}, \mathcal{P}_i, \ell)$, choose $x' \xleftarrow{\$} \mathbb{F}_p$, add $(\text{pw}_{\mathcal{Z}} || \ell, g^{x'}, x', \perp, Z^*)$ to Λ_{IC} and proceed as if x' was the value drawn uniformly random from \mathbb{F}_q at the beginning of the simulation of \mathcal{P}_i . (Cf. game \mathbf{G}_{12} .)
- **Upon receiving a query ($\text{sid}, \text{Encrypt}, k, m$) from any entity:** (Cf. games \mathbf{G}_1 and \mathbf{G}_2 .)
 If $k \notin \mathbb{F}_p$ or $m \notin \mathbb{G}$ then abort. Else:
 - if there is an entry $(k, m, *, *, c)$ in Λ_{IC} , reply with (sid, c)
 - else, choose $c \xleftarrow{\$} \mathbb{G} \setminus \{1\}$. If there already is a record $(*, *, *, *, c)$ then abort. Else, add $(k, m, \perp, \mathcal{E}, c)$ to Λ_{IC} and reply with (sid, c) .
- **Upon receiving a query ($\text{sid}, \text{Decrypt}, k, c$) from any entity:** (Cf. games \mathbf{G}_1 and \mathbf{G}_2 .)
 If $k \notin \mathbb{F}_p$ or $c \notin \mathbb{G}$ then abort. Else:
 - if there is an entry $(k, m, *, *, c)$ in Λ_{IC} , reply with (sid, m) .
 - else, choose $\alpha \xleftarrow{\$} \mathbb{F}_p^*$. If there already is a record $(*, g^\alpha, *, *, *)$ then abort. Else, add $(k, g^\alpha, \alpha, \mathcal{D}, c)$ to Λ_{IC} and reply with (sid, g^α) .
- **Upon receiving a query (sid, crs) from any entity:** (Cf. game \mathbf{G}_1 .)
 - reply with $(\text{sid}, (g, q))$.

Additionally, \mathcal{S} forwards all other instructions from \mathcal{Z} to \mathcal{A} and reports all output of \mathcal{A} towards \mathcal{Z} . Instructions of corrupting a player are only obeyed if they are received before the protocol started, i.e., before \mathcal{S} received any **NewSession** query from $\mathcal{F}_{\ell\text{-iPAKE}}$.

Fig. 29. The Simulator \mathcal{S} for the EKE2 Protocol

G Proof of Theorem 7

For the proof, we describe an honest execution of the protocol $\text{fPAKE}_{\text{RSS}}$ in the UC framework in Figure 30. See [FHH14] for a detailed description on how to execute protocols within the UC framework. This real protocol execution will be the starting point for our proof. We then proceed in a series of games, to end up with the ideal execution running with only dummy parties, a simulator and the ideal functionality $\mathcal{F}_{\text{fPAKE}}^M$. For convenience, we refer to a received protocol message as *adversarially generated* if it was

not produced by either \mathcal{P}_0 or \mathcal{P}_1 . We also refer to a query $(\text{NewKey}, \text{sid}, \mathcal{P}_i, k_i)$ from the adversary \mathcal{S} with an honest party \mathcal{P}_i as *due* if

- there is a **fresh** record of the form $(\mathcal{P}_i, \text{pw}_i)$
- this is the first **NewKey** query for \mathcal{P}_i
- there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ with $d(\text{pw}_i, \text{pw}_{1-i}) \leq \delta$ and \mathcal{P}_{1-i} is honest
- a key k_{1-i} was sent to the other party while $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ was **fresh** at the time.

We also define a masking function that reveals the positions of the identical bits:

$$m(\text{pw}, \text{pw}') := \{i | \text{pw}_i = \text{pw}'_i, i \in [n]\}$$

The parties $\mathcal{P}_0, \mathcal{P}_1$ are running with $\mathcal{F}_{\ell\text{-iPAKE}}$.

Protocol Steps:

1. When a party $\mathcal{P}_i, i \in \{0, 1\}$, receives an input $(\text{NewSession}, \text{sid}, \text{pw}_i, \text{sender})$ from \mathcal{Z} , it does the following:
 - compute $(\text{vk}, \text{sk}) \xleftarrow{\$} \text{SigGen}(1^\lambda)$
 - query n times $\mathcal{F}_{\ell\text{-iPAKE}}$ with $(\text{NewSession}, \text{sid}, (\text{pw}_i)_t, \text{vk})$, $t = 1, \dots, n$, receiving back $(\text{sid}, \ell_t, K_t)$ as answer
 - choose $U \xleftarrow{\$} \mathbb{F}_q$
 - compute $C \leftarrow \text{Share}(U)$
 - compute $E \leftarrow C + (K_t)_{t=1}^n$
 - compute $\sigma_E \leftarrow \text{Sign}(\text{vk}, E)$
 - send $(\text{sid}, E, \sigma_E, \text{vk})$ to \mathcal{P}_{1-i}
 - set $k \leftarrow U$
 - send (sid, k) towards \mathcal{Z} and terminate the session.
2. When a party $\mathcal{P}_i, i \in \{0, 1\}$, receives an input $(\text{NewSession}, \text{sid}, \text{pw}_i, \text{receiver})$ from \mathcal{Z} , it does the following:
 - query n times $\mathcal{F}_{\ell\text{-iPAKE}}$ with $(\text{NewSession}, \text{sid}, (\text{pw}_i)_t, \varepsilon)$, $t = 1, \dots, n$, receiving back $(\text{sid}, \ell_t, K'_t)$ as answer
 - if not all ℓ_t are equal or $\ell_1 \neq \text{VK}$, then abort
3. When \mathcal{P}_i , who already obtained an input $(\text{NewSession}, \text{sid}, \text{pw}_i, \text{receiver})$ and thus holds a vector $(\text{sid}, \ell_t, K'_t)$ obtained from $\mathcal{F}_{\ell\text{-iPAKE}}$, receives a message $(\text{sid}, E, \sigma_E, \text{vk})$ from \mathcal{P}_{1-i} , it does the following:
 - set $K' := (K'_t)_{t \in [n]}$
 - abort if $\text{vk} \neq \ell_1$
 - abort if $\text{Vfy}(\text{vk}, \sigma_E, E) = 0$
 - compute $U' \leftarrow \text{Reconstruct}(E - K')$
 - set $k \leftarrow U'$
 - send (sid, k) towards \mathcal{Z} and terminate the session.

Fig. 30. A UC Execution of $\text{fPAKE}_{\text{RSS}}$

Game G_0 : The real protocol execution. This is the real execution of $\text{fPAKE}_{\text{RSS}}$ where the environment \mathcal{Z} runs the protocol (cf. Figure 10) with parties \mathcal{P}_0 and \mathcal{P}_1 , both having access to an ideal $\ell\text{-iPAKE}$ functionality $\mathcal{F}_{\ell\text{-iPAKE}}$, and an adversary \mathcal{A} that, w.l.o.g., can be assumed to be the dummy adversary as shown in [Can01, section 4.4.1].

Game G₁: Modeling the ideal layout. We first make some purely conceptual changes that do not modify the input/output interfaces of \mathcal{Z} . We add one relay (also referred to as *dummy party*) on each of the wires between \mathcal{Z} and a party. We also add one relay covering all the wires between the dummy parties and real parties and call it \mathcal{F} (and let \mathcal{F} relay messages according to the original wires). We group all the formerly existing instances except for \mathcal{Z} into one machine and call it \mathcal{S} . Note that this implies that \mathcal{S} executes the code of the ℓ -iPAKE functionality $\mathcal{F}_{\ell\text{-iPAKE}}$. The differences are depicted in Figure 20 with \mathcal{F}_{OT} replaced by $\mathcal{F}_{\ell\text{-iPAKE}}$.

Game G₂: Building $\mathcal{F}_{\text{iPAKE}}^M$. In this game, we start modeling $\mathcal{F}_{\text{iPAKE}}^M$. First, we let \mathcal{F} maintain a list of tuples of the form $(\mathcal{P}_i, \text{pw}_i)$. Upon receiving a query of the form $(\text{NewSession}, \text{sid}, \text{pw}_i, \text{role})$ from party \mathcal{P}_i , if this is the first **NewSession**-query, or if this is the second **NewSession**-query and there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$, then \mathcal{F} records $(\mathcal{P}_i, \text{pw}_i)$ and marks this record as **fresh**. In any case the query $(\text{NewSession}, \text{sid}, \mathcal{P}_i, \text{pw}_i, \text{role})$ is relayed to \mathcal{S} . Now that \mathcal{F} knows about pass-strings, we can add a **TestPwd** interface to \mathcal{F} as described in Figure 2, using leakage functions L_c^M, L_m^M and L_f^M . We let \mathcal{S} parse outputs towards \mathcal{F} as $(\text{NewKey}, \text{sid}, \mathcal{P}_i, k_i)$ by adding the **NewKey** tag and the name of the party who produced the output. Additionally, we let \mathcal{F} translate this back to (sid, k_i) , send it to \mathcal{Z} via \mathcal{P}_i and mark the corresponding record as **completed**.

None of these modifications changes the output towards \mathcal{Z} compared to the previous game G₁.

Game G₃: \mathcal{F} generates a random session key for an interrupted session. Upon receiving a query $(\text{NewKey}, \text{sid}, \mathcal{P}_i, k_i)$ from \mathcal{S} , if there is a record of the form $(\mathcal{P}_i, \text{pw}_i)$ that is marked as **interrupted**, and this is the first **NewKey** query for \mathcal{P}_i , we let \mathcal{F} output a random session key of length λ to \mathcal{P}_i . Otherwise, it continues to relay k_i . Since the simulators described in game G₂ and game G₃ do not make use of the **TestPwd** interface, none of the records of \mathcal{F} are marked as **interrupted** and thus the output towards \mathcal{Z} is equally distributed in both games.

Game G₄: \mathcal{S} handles dictionary attacks using the **TestPwd interface.** In this game, we only change the simulation. Consider the following setting: \mathcal{P}_i obtained input $(\text{NewSession}, \text{sid}, \text{pw}_i, \text{role})$ and \mathcal{P}_{1-i} is corrupted and already provided its inputs to $\mathcal{F}_{\ell\text{-iPAKE}}$. In this situation, \mathcal{S} will proceed simulation of \mathcal{P}_i as follows: \mathcal{S} assembles $\text{pw}_{\mathcal{Z}} \in \mathbb{F}_p^n$ from the queries to $\mathcal{F}_{\ell\text{-iPAKE}}$ that \mathcal{P}_{1-i} issued. \mathcal{S} sends $(\text{TestPwd}, \text{sid}, \mathcal{P}_i, \text{pw}_{\mathcal{Z}})$ to \mathcal{F} , obtaining either “wrong guess”, “correct guess” and perhaps also a mask $M \subseteq [n]$ from \mathcal{F} . If \mathcal{S} does not receive a mask, \mathcal{S} is not modified further. Else, let $I := [n] \setminus M$ the set of mismatching indices, and $d := |I| \leq \gamma$ their number. \mathcal{S} sets up keys $K, K' \in \mathbb{F}_q^n$ with $K_t = K'_t \stackrel{\$}{\leftarrow} \mathbb{F}_q$ for the matching indices $t \in M$ and $K_t, K'_t \stackrel{\$}{\leftarrow} \mathbb{F}_p^2$ for the mismatching indices $t \in I$, where K' denotes the $\mathcal{F}_{\ell\text{-iPAKE}}$ output of \mathcal{P}_{1-i} . \mathcal{S} now continues the simulation of \mathcal{P}_i using K as output of $\mathcal{F}_{\ell\text{-iPAKE}}$.

We have to analyze different cases depending on the different outcomes of **TestPwd**. However, note that the modifications only have an impact on the output k_i of \mathcal{P}_i .

if the record gets **interrupted**, and only affect the transcript if the answer to the **TestPwd** query contains a mask. Considering the case where **TestPwd**

- outputs \mathbf{m} and sets the record **compromised**, i.e. $d \leq \gamma$ since the distribution of K, K' only depends on the mask of the pass-strings, the view of \mathcal{Z} is identically distributed in game \mathbf{G}_4 and game \mathbf{G}_3 ;
- outputs “wrong guess” and sets the record **interrupted**, i.e. $d > \gamma$: \mathcal{P}_i will now obtain a randomly chosen session key from \mathcal{F} , substituting the key k_i computed by \mathcal{S} . If \mathcal{P}_i obtained **role** = **sender**, the output in game \mathbf{G}_4 and game \mathbf{G}_3 is equally distributed since the honest sender outputs a random \mathbb{F}_q value according to the protocol description. If \mathcal{P}_i obtained **role** = **receiver**, both outputs are indistinguishable with overwhelming probability due to the smoothness of the RSS, since in game \mathbf{G}_3 at least $\gamma + 1$ shares are random.

Game \mathbf{G}_5 : Excluding man-in-the-middle attacks. Again, in this game, we only change the simulation. We now consider the case where \mathcal{Z} injects a message into a session where both parties are honest. We modify \mathcal{S} as follows: upon receiving an adversarially generated $(\text{sid}, M_{\mathcal{Z}}, \sigma_{\mathcal{Z}}, \text{vk}_{\mathcal{Z}})$ from \mathcal{Z} intended for party \mathcal{P}_i , \mathcal{S} aborts. Observe that the simulation is only changed compared to the previous game if it is not aborted due to protocol instructions. This means that both games are equal unless all checks pass, especially $\text{Vfy}(\text{vk}_{\mathcal{Z}}, \sigma_{\mathcal{Z}}, M_{\mathcal{Z}}) = 1$. Any distinguisher between game \mathbf{G}_5 and game \mathbf{G}_4 can thus be turned into a forger of a valid message w.r.t the verification key of an honest party. Indistinguishability thus follows from the security of the one-time signature scheme.

Game \mathbf{G}_6 : \mathcal{F} aligns session keys. Upon receiving a query $(\text{NewKey}, \text{sid}, \mathcal{P}_i, k_i)$ from \mathcal{S} , if this query is *due* then output (sid, k_{1-i}) to \mathcal{P}_i where k_{1-i} is the session key that was formerly sent to the other party.

We now analyze distinguishability of this game from game \mathbf{G}_5 . If \mathcal{Z} tampered with the transcript, the simulation in game \mathbf{G}_5 ensures that the simulation aborts and there is thus no **NewKey** query for \mathcal{P}_i . On the other hand, if \mathcal{Z} does not advise \mathcal{A} to tamper with any message, perfect correctness of $\text{fPAKE}_{\text{RSS}}$ protocol ensures that, in case of a due record where the parties hold close pass-strings $\text{pw}_i, \text{pw}_{1-i}$ with $d(\text{pw}_i, \text{pw}_{1-i}) \leq n - r$, the output of \mathcal{F} towards \mathcal{Z} is the same as in the previous game \mathbf{G}_5 . Observe that perfect correctness directly follows from the perfect correctness of $\mathcal{F}_{\ell\text{-iPAKE}}$ and the r -robustness of the secret sharing, which is *always* able to correct up to $n - r$ errors.

Note that \mathcal{F} still differs from the functionality $\mathcal{F}_{\text{fPAKE}}^M$ in some aspects. First, it does not output randomly generated session keys towards \mathcal{Z} for honest sessions. Furthermore, it reports all pass-strings to \mathcal{S} . We will take care of these remaining differences in the next games.

Game \mathbf{G}_7 : In some cases, \mathcal{F} generates a random session key when the other party is corrupted. Upon receiving a **NewKey** query $(\text{NewKey}, \text{sid}, \mathcal{P}_i, k_i)$ from \mathcal{S} , if there is a fresh record of the form $(\mathcal{P}_i, \text{pw}_i)$, and this is the first **NewKey** query for \mathcal{P}_i , \mathcal{P}_i is honest and \mathcal{P}_{1-i} corrupted and there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ with $d(\text{pw}_i, \text{pw}_{1-i}) > \delta$, we let \mathcal{F} pick a new random key k from \mathbb{F}_q and send (sid, k) to \mathcal{P}_i .

The simulation ensures that the record $(\mathcal{P}_i, \text{pw}_i)$ is either compromised or interrupted (cf. description of the simulator in game \mathbf{G}_4). Thus, the modification has no effect since it only concerns fresh records.

Game \mathbf{G}_8 : \mathcal{F} generates a random session key for an honest session. Upon receiving a **NewKey** query $(\text{NewKey}, \text{sid}, \mathcal{P}_i, k_i)$ from \mathcal{S} , if there is a fresh record of the form $(\mathcal{P}_i, \text{pw}_i)$, and this is the first **NewKey** query for \mathcal{P}_i , both parties are honest and the **NewKey** query is not *due*, we let \mathcal{F} pick a new random key k from \mathbb{F}_q and send (sid, k) to \mathcal{P}_i .

In other words, \mathcal{F} now generates a random session key upon a first **NewKey** query for an honest party \mathcal{P}_i with fresh record $(\mathcal{P}_i, \text{pw}_i)$ where \mathcal{P}_{1-i} is also honest, if (at least) one of the following events happen:

1. There is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ with $d(\text{pw}_i, \text{pw}_{1-i}) > \delta$; then, the probability that k_i was already random in game \mathbf{G}_7 is overwhelming due to the $r - 1$ -smoothness of the RSS on random secrets. Note that to apply this property it is crucial that both parties are honest and thus the value U is randomly chosen.
2. No session key was sent to \mathcal{P}_{1-i} yet; we just have to consider the case where there is a record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ with $d(\text{pw}_i, \text{pw}_{1-i}) \leq \delta$ since we already dealt with the other case in the first event. Due to the r -robustness of the RSS, the session key in the previous game was U , which is distributed uniformly random in \mathbb{F}_q .
3. If there was a session key sent to \mathcal{P}_{1-i} , the record $(\mathcal{P}_{1-i}, \text{pw}_{1-i})$ was not fresh and thus interrupted or compromised at that time; since our simulation never issues **TestPwd** queries for honest sessions (in fact, game \mathbf{G}_5 states that \mathcal{S} aborts upon man-in-the-middle attacks with overwhelming probability), this event can not happen in our simulation.

Game \mathbf{G}_9 : **Simulating without pass-string if both parties are honest.** In case of receiving a $(\text{NewSession}, \text{sid}, \text{pw}_i, \text{role})$ from an honest \mathcal{P}_i , we modify \mathcal{F} by forwarding only $(\text{NewSession}, \text{sid}, \mathcal{P}_i, \text{role})$ to \mathcal{S} . We now have to modify \mathcal{S} to proceed simulation without knowing pw . Upon receiving $(\text{NewSession}, \text{sid}, \mathcal{P}_i, \text{role})$ from \mathcal{F} for an honest \mathcal{P}_i , we let \mathcal{S} draw uniformly at random a “dummy” pass-string $\text{pw}_\mathcal{S}$ and proceed the simulation of \mathcal{P}_i using $\text{pw}_\mathcal{S}$ as a pass-string.

We first observe that \mathcal{Z} is oblivious of k_i contained in the $(\text{NewKey}, \text{sid}, \mathcal{P}_i, k_i)$ query that \mathcal{S} will eventually send to \mathcal{F} during the simulation (since \mathcal{F} never lets the simulator determine k_i for an honest session). We thus only have to consider the case where \mathcal{P}_i has $\text{role} = \text{sender}$. We show that \mathcal{Z} , knowing $\text{pw}_i, \text{pw}_{1-i}$ and seeing two transcripts, cannot tell which one was generated using $\text{pw}_i, \text{pw}_{1-i}$ and which one was generated using $\text{pw}_\mathcal{S}, \text{pw}_{1-i}$ with a random $\text{pw}_\mathcal{S}$ unknown to \mathcal{Z} . But this is trivial since the distribution of the values U, K does not depend on the pass-strings: U is randomly chosen from \mathbb{F}_q . $\mathcal{F}_{\ell\text{-iPAKE}}$ ensures that K is randomly chosen from \mathbb{F}_q^n .

Game \mathbf{G}_{10} : **Simulating without pass-string if someone is corrupted.** Upon receiving $(\text{NewSession}, \text{sid}, \text{pw}_i, \text{role})$ from \mathcal{P}_i where \mathcal{P}_{1-i} is corrupted, we modify \mathcal{F} to only relay $(\text{NewSession}, \text{sid}, \mathcal{P}_i, \text{role})$ to \mathcal{S} . Additionally, we let \mathcal{S} draw uniformly at random a “dummy” pass-string $\text{pw}_\mathcal{S}$ and proceed the simulation of \mathcal{P}_i using $\text{pw}_\mathcal{S}$ as a pass-string. Note that due to the simulation described in game \mathbf{G}_4 , \mathcal{S} will

ask a **TestPwd** query, and after this query the simulation described in that game is already independent of pw_i except when \mathcal{F} 's reply does not contain a mask, i.e., $d(\text{pw}_i, \text{pw}_{1-i}) > \gamma$. In this case, we now let \mathcal{S} set the output of $\mathcal{F}_{\ell\text{-iPAKE}}$ towards \mathcal{P}_i to be a random $K \xleftarrow{\$} \mathbb{F}_q^n$.

Regarding indistinguishability, first note that in any case the input of \mathcal{P}_i to $\mathcal{F}_{\ell\text{-iPAKE}}$ does not impact any values and thus we only have to argue further in case \mathcal{S} is modified. Thus, \mathcal{P}_i 's record will get interrupted and \mathcal{P}_i will obtain a uniformly random session key from \mathcal{F} , meaning that we only have to consider the case where \mathcal{P}_i has **role** = **sender** and argue indistinguishability of E, σ_E, vk generated with either K depending on pw_i (as in the previous game) or $K \xleftarrow{\$} \mathbb{F}_q^n$ (as in the current game). Opposed to the situation in game **G**₉, note that now \mathcal{Z} knows K' .

Since $d(\text{pw}_i, \text{pw}_{1-i}) > \gamma = n - t - 1$, at most t components of K' are the same as K in **G**₉ with large probability $1 - \frac{n-t}{q}$, and thus w.h.p. \mathcal{Z} learns at most t components of C . Hence the transcript of the current and previous games are indistinguishable due to the strong t -privacy of the RSS.

Observe that now \mathcal{F} is equal to $\mathcal{F}_{\text{iPAKE}}^M$ and \mathcal{S} is equal to the simulator described in Figure 31. The theorem thus follows.

The simulator \mathcal{S} , initialized with a security parameter λ , initializes the dummy adversary \mathcal{A} . \mathcal{S} emulates an ideal labeled iPAKE functionality $\mathcal{F}_{\ell\text{-iPAKE}}$ as depicted in Figure 26 for all calling entities in the system^a. Additionally, \mathcal{S} interacts with an ideal functionality $\mathcal{F}_{\text{iPAKE}}^M$ and a distinguisher, the environment \mathcal{Z} , via the following queries:

- **Upon receiving a query** (**NewSession**, **sid**, \mathcal{P}_i , **role**) **from** $\mathcal{F}_{\text{iPAKE}}^M$: initialize a party \mathcal{P}_i and connect it to \mathcal{A} .
 - If \mathcal{P}_{1-i} is honest, \mathcal{S} proceeds the UC protocol execution as described in Figure 30 using $\text{pw}_{\mathcal{S}} \xleftarrow{\$} \mathbb{F}_p$ as pass-string for \mathcal{P}_i . (Cf. game **G**₉.)
 - If \mathcal{P}_{1-i} is corrupted, then \mathcal{S} waits until \mathcal{P}_{1-i} submitted n queries to $\mathcal{F}_{\ell\text{-iPAKE}}$ and then assembles $\text{pw}_{\mathcal{Z}} \in \mathbb{F}_p^n$ from them. \mathcal{S} sends (**TestPwd**, **sid**, \mathcal{P}_i , $\text{pw}_{\mathcal{Z}}$) to $\mathcal{F}_{\text{iPAKE}}^M$. If \mathcal{S} receives back a mask M , let $I := [n] \setminus M$, and \mathcal{S} sets up n \mathbb{F}_q -keys K with $K_t = K'_t \forall t \in I$ and $K_t \xleftarrow{\$} \mathbb{F}_p \neq K'_t \forall t \in M$, where K' denotes the output of $\mathcal{F}_{\ell\text{-iPAKE}}$ towards \mathcal{P}_{1-i} . \mathcal{S} now continues the simulation of \mathcal{P}_i using K as outputs of $\mathcal{F}_{\ell\text{-iPAKE}}$. (Cf. game **G**₄.) If \mathcal{S} does not receive a mask, \mathcal{S} sets the output of $\mathcal{F}_{\ell\text{-iPAKE}}$ towards \mathcal{P}_i to be $K \xleftarrow{\$} \mathbb{F}_q^n$. (Cf. game **G**₁₀.)
- **If an internally simulated party \mathcal{P}_i produces an output** (**sid**, **k** _{i}):
Send (**NewKey**, **sid**, \mathcal{P}_i , **k** _{i}) to $\mathcal{F}_{\text{iPAKE}}^M$.
- **If \mathcal{Z} sends** (**sid**, $M_{\mathcal{Z}}$, $\sigma_{\mathcal{Z}}$, $\text{vk}_{\mathcal{Z}}$) **to an honest party \mathcal{P}_i** : if \mathcal{P}_{1-i} is honest, \mathcal{S} aborts after the Vfy step in the protocol, regardless of its outcome. (Cf. game **G**₅.)

Additionally, \mathcal{S} forwards all other instructions from \mathcal{Z} to \mathcal{A} and reports all output of \mathcal{A} towards \mathcal{Z} . Instructions of corrupting a player are only obeyed if they are received before the protocol started, i.e., before \mathcal{S} received any **NewSession** query from $\mathcal{F}_{\text{iPAKE}}^M$.

^a An entity is any internally simulated ITM such as parties or the real-world adversary as well as ITMs outside \mathcal{S} such as the distinguisher \mathcal{Z} .

Fig. 31. The Simulator \mathcal{S} for $\text{fPAKE}_{\text{RSS}}$

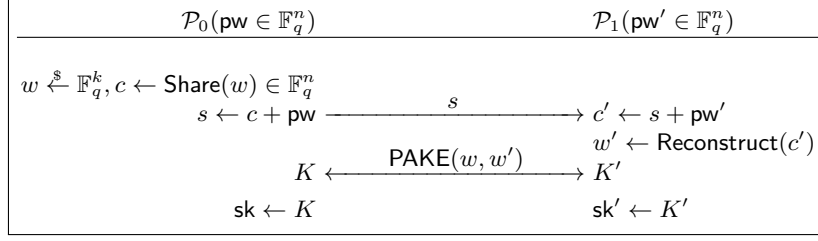


Fig. 32. A First Natural Construction (with code-offset fuzzy sketch and PAKE)

H A Natural (But Failed) Approach to fPAKE

A natural idea for building a fPAKE is the use of a fuzzy extractor [DRS04, Boy04], that allows to extract a common secret from two strings close enough, and to compose it with a regular PAKE. This approach was introduced in [BDK⁺05] (Section 4). Their protocol uses the code-offset construction of a fuzzy sketch [DRS04], a.k.a. fuzzy commitment [JW99], to implement a fuzzy-extractor as a two-party primitive. It is presented in Figure 32.

Theorem 15. *The construction from Figure 32 cannot securely realize $\mathcal{F}_{\text{fPAKE}}^P$.*

Proof. Consider the following attack by \mathcal{Z} . \mathcal{Z} sends a randomly chosen pw as input to an honest \mathcal{P}_0 and obtains a sketch s from \mathcal{A} . It then computes $c \leftarrow s - \text{pw}$ and outputs 1 if c is in the image of Share . In the real world, this happens with probability 1. Now assume there is a simulator \mathcal{S} outputting a simulated sketch \tilde{s} in the ideal world. Since \mathcal{S} does not get to learn pw unless it succeeds at a TestPw query, observe that this output may not depend on pw except with some small (but non-negligible) probability p , namely the probability of guessing a pass-string that makes $\mathcal{F}_{\text{fPAKE}}^P$ output pw . Thus, with probability $1 - p \approx 1$, $\tilde{c} := \tilde{s} - \text{pw}$ is randomly distributed in \mathbb{F}_q^n and lies in the image of Share only with probability $1/q^{mn-l}$. More formally, the probability that \mathcal{Z} outputs 1 in the ideal world is

$$\begin{aligned}
\Pr[\tilde{c} \in \text{Im}(\text{Share})] &= \Pr[\tilde{c} \in \text{Im}(\text{Share}) | \mathcal{S} \text{ depends on } \text{pw}] \cdot p \\
&\quad + \Pr[\tilde{c} \in \text{Im}(\text{Share}) | \mathcal{S} \text{ does not depend on } \text{pw}] \cdot (1 - p) \\
&\leq p + 1/q^{mn-l}(1 - p) \approx p.
\end{aligned}$$