A preliminary version of this paper appears in the proceedings of the *37th International Cryptology Conference (CRYPTO 2017)*, © IACR 2017. This is the full version.

PRF-ODH: Relations, Instantiations, and Impossibility Results

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June 2, 2017

Abstract. The pseudorandom-function oracle-Diffie–Hellman (PRF-ODH) assumption has been introduced recently to analyze a variety of DH-based key exchange protocols, including TLS 1.2 and the TLS 1.3 candidates, as well as the extended access control (EAC) protocol. Remarkably, the assumption comes in different flavors in these settings and none of them has been scrutinized comprehensively yet. In this paper here we therefore present a systematic study of the different PRF-ODH variants in the literature. In particular, we analyze their strengths relative to each other, carving out that the variants form a hierarchy. We further investigate the boundaries between instantiating the assumptions in the standard model and the random oracle model. While we show that even the strongest variant is achievable in the random oracle model under the strong Diffie–Hellman assumption, we provide a negative result showing that it is implausible to instantiate even the weaker variants in the standard model via algebraic black-box reductions to common cryptographic problems.

1 Introduction

Proposing new cryptographic assumptions is a valid strategy to analyze or design protocols which escape a formal treatment so far. Yet, the analysis of the protocol, usually carried out via a reduction to the new assumption, is only the first step. Only the evaluation of the new assumption completes the analysis and yields a meaningful security claim.

1.1 The PRF-ODH Assumption

In the context of key exchange protocols, a new assumption, called the pseudorandom-function oracle-Diffie–Hellman (PRF-ODH) assumption has recently been put forward by Jager et al. [JKSS12] for the analysis of TLS 1.2. It is a variant of the oracle-Diffie–Hellman assumption introduced by Abdalla et al. [ABR01] in the context of the encryption scheme DHIES. The PRF-ODH assumption basically says that the function value $\mathsf{PRF}(g^{uv}, x^\star)$ for a DH key g^{uv} looks random, even if given g^u and g^v and if seeing related values $\mathsf{PRF}(S^u, x)$ and/or $\mathsf{PRF}(T^v, x)$ for chosen values S, T, and x.

The PRF-ODH appears to be a natural assumption for any DH-based key exchange protocol, aiming at security against man-in-the-middle attacks (see Figure 1). In DH-based protocols both parties, the client and the server, exchange values g^u, g^v and locally compute the session key by applying a key derivation (or pseudorandom) function to the key g^{uv} and usually some parts of the transcript. The man-in-the-middle adversary can now try to attack the server's session key $\mathsf{PRF}(g^{uv}, \dots)$ by submitting a modified value S instead of g^v to the client, yielding a related key $\mathsf{PRF}(S^u, \dots)$ on the client's side. The PRF-ODH assumption guarantees now that the server's key is still fresh.

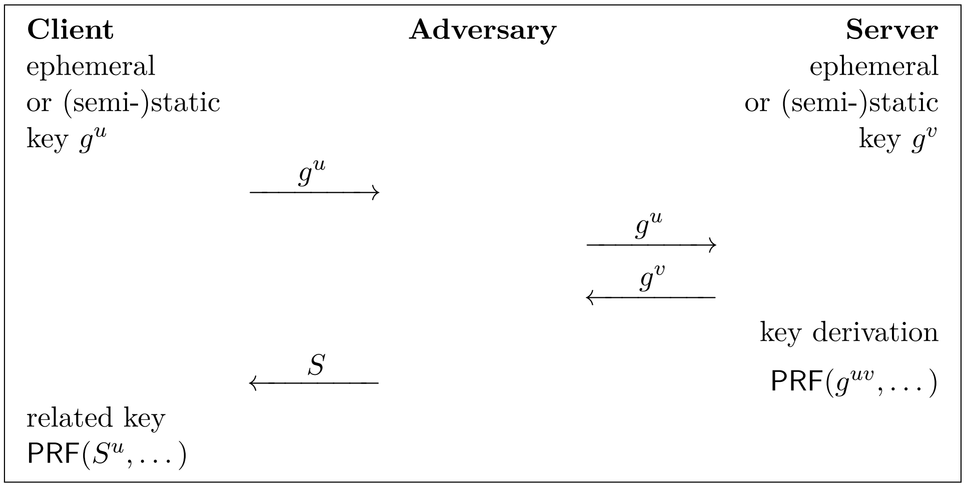

Figure 1: Origin of the PRF-ODH assumption: Man-in-the-middle attack on DH-based key exchange protocol.

Note that simple authentication of transmissions does not provide a remedy against the above problem. The adversary could act under a different, corrupt server identity towards the client, and only re-use the Diffie–Hellman data, authenticated under the corrupt server's key. Then the Diffie–Hellman keys in the executions would still be non-trivially related. This happens especially if keys are used in multiple sessions. Another problem is that some protocols may derive keys early, before applying signatures, e.g., such as for handshake encryption as well as in the post-handshake authentication mechanism in TLS 1.3 [Res17].

It therefore comes as no surprise that the PRF-ODH assumption has been used in different protocols for the security analysis, including the analysis of the TLS 1.2 [DR08] ephemeral and static Diffie–Hellman handshake modes [JKSS12, KPW13, BFK$^+$14], the TLS 1.3 [Res17] Diffie–Hellman-based and resumption handshake candidates [DFGS15a, DFGS15b, DFGS16] as well as 0-RTT handshake candidates [FG17], and a 0-RTT extension of the extended access control (EAC) protocol [BF17], for the original EAC protocol listed, for example, in Document 9303 of the International Civil Aviation Organization [Int15]. Notably, these scientific works use different versions of the PRF-ODH assumption, due to the different usages of the key shares g^u, g^v. These key shares can be ephemeral (for a single session), semi-static (for a small number of sessions), or static (for multiple sessions). Therefore, the man-in-the middle adversary may ask to see no related key for either key share, a single related key, or multiple related keys. For instance, while Jager et al. [JKSS12] required only security against a single query for one of the two key shares, Krawczyk et al. [KPW13] modify the original PRF-ODH assumption because they require security against multiple oracle queries against this key share. In [FG17] an extra query to the other key share has been added, and [BF17] require multiple queries to both key shares.

1.2 Evaluating the PRF-ODH Assumptions

Consequently, and to capture all of the above assumptions simultaneously, we generally speak of the lrPRF-ODH assumption, allowing the adversary no ($\mathsf{l}, \mathsf{r} = \mathsf{n}$), a single ($\mathsf{l}, \mathsf{r} = \mathsf{s}$), or multiple ($\mathsf{l}, \mathsf{r} = \mathsf{m}$) related key queries, for the "left" key g^u or the "right" key g^v. Such queries are handled by oracles ODH_u and ODH_v, returning the corresponding pseudorandom function value. This results in nine variants, for each combination $\mathsf{l}, \mathsf{r} \in \{\mathsf{n}, \mathsf{s}, \mathsf{m}\}$. We also discuss some more fine-grained distinctions, e.g., if the adversary learns both keys g^u, g^v before choosing the input $x^\star$ for the challenge value $\mathsf{PRF}(g^{uv}, x^\star)$, or if $x^\star$ can only depend on g^u.

To evaluate the strengths of the different types of lrPRF-ODH assumptions one can ask how the variants relate to each other. Another important aspect is the question whether, and if so, to which (well investigated) Diffie–Hellman problem it possibly relates to, e.g., the computational Diffie–Hellman (CDH), the decisional Diffie–Hellman (DDH), the strong Diffie–Hellman (StDH), or the even more general Gap-Diffie–Hellman (GapDH) problem. While the answer to this question may rely on the random oracle model, the final issue would be to check if (any version of) the assumption can be instantiated in the standard model.

Especially the question whether the PRF-ODH assumption (or which variant) can be instantiated in the standard model is of utmost interest. Some of the aforementioned works refer to the(ir) PRF-ODH assumption as a standard-model assumption, since there is no immediate reference to a random oracle. This would not only apply to the schemes analyzed with respect to the PRF-ODH assumption, but potentially also to other works where the Gap-DH or related assumptions in the random oracle have been used for the analysis, yet where the PRF-ODH assumption is a promising alternative for carrying out a proof. Examples include the QUIC protocol [FG14, LJBN15] and OPTLS [KW16] which forms the base for TLS 1.3.

1.3 Our Results

Figure 2 gives an overview over our results. We explain the details next.

Instantiations. Our first contribution is to discuss instantiation possibilities of the PRF-ODH variants. We stress that some of these results mainly confirm the expectation: the nnPRF-ODH assumption where no oracle queries are allowed can be based upon the decisional Diffie–Hellman assumption DDH, and the one-sided assumptions mnPRF-ODH and nmPRF-ODH where the adversary has (multiple) access to either oracle ODH_u or ODH_v can be based on the strong Diffie–Hellman assumption in the random oracle model. The strong DH assumption (StDH) demands that the adversary solves the computational problem of computing g^{uv} from g^u, g^v, but having access to a decisional oracle $\mathsf{DDH}(g^u, \cdot, \cdot)$ checking for DH tuples. Such checks are necessary to provide consistency when simulating the random oracle through lazy sampling, i.e., in the case that random values are only sampled on their first explicit usage. The proofs for mnPRF-ODH and nmPRF-ODH appear already implicitly in previous work about key exchange, e.g., [Kra05, Ust08, DF11, FG14, LJBN15, KW16, LXZ$^+$16], but where the reduction to the StDH problem in the random oracle model has been carried out by dragging along all the steps of the key exchange protocols.

Our final instantiation result for the strongest notion mmPRF-ODH holds in the random oracle model under the strong DH (StDH) assumption. Surprisingly, the proof is less straightforward than one would expect, since the availability of both oracles ODH_u and ODH_v imposes the need for further consistency checks between cross-over calls for the two oracles in the simulation. We show that such consistency checks can indeed be implemented assuming StDH, but causing a square-root loss in the security reduction to StDH. This loss is due to the fact that in an intermediate step we go through the square-DH problem SqDH (given g, g^v compute g^{v^2}) to which CDH reduces by making two calls to the square-DH problem adversary (see, e.g., [Kil01]), effectively squaring the success probability.

The instantiations are shown through the boxes with dotted surrounding lines in Figure 2. We also discuss briefly the relationship to related-key security for pseudorandom functions, where the adversary can ask to see values for transformed keys $\phi(K)$. While similar in spirit at first glance, it seems to us that the notions differ in technical details which makes it hard to relate them.

Relations. The instantiation results give a sort of general method to achieve any PRF-ODH notion, leaving open the possibility that one notion may be actually easier to achieve. This is even more relevant in light of the fact that previous works used different notions. In order to support a better comparison between the various notions we relate them in terms of strength of the assumption. Some of these relationships,

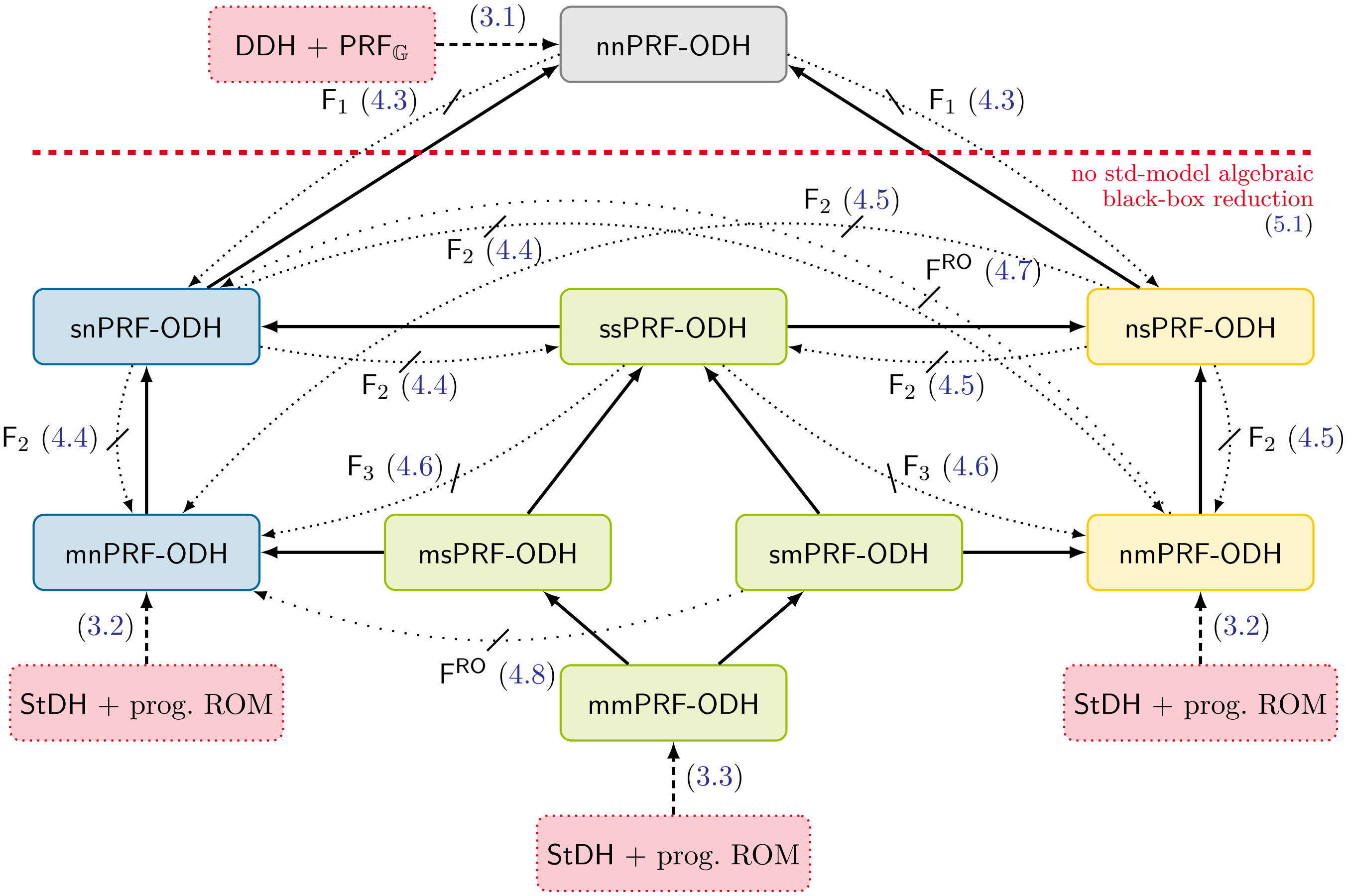

Figure 2: Relations between the different PRF-ODH variants (in solid-line rounded rectangles) from Definition 2.1 and other assumptions (in dotted-line rounded rectangles). Solid arrows indicate the trivial implications between PRF-ODH variants, dashed arrows indicate implications we establish. Struck-out, densely dotted arrows indicate separations in the standard model via the indicated function $\mathsf{F}_n \in \mathcal{F}$ (cf. Definition 4.1). Struck-out, sparsely dotted arrows indicated separations in the random-oracle model. The dashed horizontal line demarcates the boundary below which our impossibility result for standard-model algebraic black-box reductions from Section 5 holds. Numbers in parentheses indicate the respective propositions and theorems.

especially implications, are easy to establish. For example, since the adversary in the mmPRF-ODH game can always forgo using its ODH_v oracle, this immediately implies mnPRF-ODH security. All implications are marked by solid arrows in Figure 2.

As for separations we are able to rule out a number of implications unconditionally. By this we mean that we only make the minimal assumption that a secure instantiation exists, and then build one still satisfying this notion but not the stronger one. These separations are displayed in Figure 2 through dotted arrows.

We are also able to separate further notions conditionally, using random oracles and a plausible number-theoretic assumption. Namely, under these assumptions, the notion of snPRF-ODH (with a single call to ODH_u) is strictly stronger than the nmPRF-ODH notion where the adversary can ask the ODH_v oracle multiple times but does not get access to the ODH_u oracle. With a similar strategy we can also separate mnPRF-ODH with multiple ODH_u queries from smPRF-ODH, where the adversary can now make one extra call to ODH_u on top of the ODH_v queries.

The conditional separations are not symmetric in the sense that they apply to the other oracle as well.

The reason is that these results exploit that the adversary receives g^u before g^v, such that the converse does not simply follow. Besides these opposite cases there are also some other cases where we could not provide a separation, e.g., from mmPRF-ODH to msPRF-ODH. We give more insights within.

Impossibility result. The third important contribution is our impossibility result. We show that proving security of even the mild snPRF-ODH or nsPRF-ODH notions based on general cryptographic problems is hard. Besides the common assumption that the reduction uses the adversary only as a black box, we also assume that the reduction is algebraic. This means that whenever the reduction passes a group element A to the outside, it knows a representation $(\alpha_1, \alpha_2, \dots)$ such that $A = \prod g_i^{\alpha_i}$ for the reduction's input values $g_1, g_2, \dots$. This notion of algebraic reductions has been used in other separation works before, e.g., [BV98, PV05, GBL08]. Unlike generic reductions, algebraic reductions can take advantage of the representation of group elements.

In detail, we then show via a meta-reduction technique [GMR88, BV98, PV05], that one cannot prove security of the snPRF-ODH or nsPRF-ODH assumption via algebraic black-box reductions to a class of cryptographic problems. The problems we rule out are quite general, saying that the adversary receives some input, can interact multiple times with a challenger in an arbitrary way, and should then provide a solution. We remark that we also need to augment this problem by a Diffie–Hellman problem in order to give a reference point for the algebraicity of the reduction. Our result also requires that the decisional square-DH problem is hard, i.e., that g, g^v, g^{v^2} is indistinguishable from g, g^v, g^z for random v, z.[1]

In a sense, our negative result, displayed by the dashed horizontal line on top in Figure 2, is optimal in terms of the relation of PRF-ODH assumptions, as it rules out exactly the notions "one above" the nnPRF-ODH notion with a standard model instantiation. We still note that the restrictions on the reduction, and the additional assumption, may allow to bypass our result. This also means that our implications and separations between the different notions, established earlier, are not moot.

Implications for practical key derivation functions. Since the PRF-ODH assumptions have been used in connection with applied protocols like TLS, we finally address the question which security guarantees we get for practical key derivation functions used in such protocols. We are especially interested in HMAC [KBC97] on which the key derivation function HKDF [Kra10, KE10] is based upon. Our instantiation results in the random oracle so far treat the key derivation function as a monolithic random oracle, whereas key derivation functions like HMAC have an iterative structure. At the same time, our impossibility result tells us that giving a standard-model proof for HMAC, based on say collision-resistance of the compression function, may be elusive. We thus make the assumption that the compression function is a random oracle.

We show that HMAC provides the strong notion of mmPRF-ODH security, assuming StDH and that the compression function is a random oracle. We note that Coron et al. [CDMP05] show that a variant of HMAC is indifferentiable from a random oracle, and Krawczyk [Kra10] briefly remarks that the result would carry over to the actual HMAC construction. However, in HKDF the HMAC function is applied in a special mode in which the key part is hashed first, and it is therefore unclear if our result for the monolithic random oracle immediately applies. But based on the techniques used in the instantiation part we can give a direct proof of the security of (the general mode of) HMAC.

[1] While the computational version of the square-DH problem is known to be equivalent to the CDH problem, it is unclear if the decisional version follows from DDH.

2 PRF-ODH Definition

Different variants of the new PRF oracle-Diffie–Hellman ($\mathsf{PRF\text{-}ODH}$) assumption have been introduced and used in the literature in the context of key exchange protocols. In this section we first provide a generic $\mathsf{PRF\text{-}ODH}$ assumption definition capturing all different flavors and discuss its relation to previous occurrences [JKSS12, KPW13, DFGS15b, DFGS16, BF17, FG17].

Definition 2.1 (Generic $\mathsf{PRF\text{-}ODH}$ assumption)**.** *Let $\mathbb{G}$ be a cyclic group of order q with generator g. Let $\mathsf{PRF}\colon \mathbb{G} \times \{0,1\}^* \to \{0,1\}^\lambda$ be a pseudorandom function that takes a key $K \in \mathbb{G}$ and a label $x \in \{0,1\}^*$ as input and outputs a value $y \in \{0,1\}^\lambda$, i.e., $y \leftarrow \mathsf{PRF}(K, x)$.*

We define a generic security notion $\mathsf{lrPRF\text{-}ODH}$ which is parameterized by $\mathsf{l}, \mathsf{r} \in \{\mathsf{n}, \mathsf{s}, \mathsf{m}\}$ indicating how often the adversary is allowed to query a certain "left" resp. "right" oracle (ODH_u resp. ODH_v) where n indicates that no query is allowed, s that a single query is allowed, and m that multiple (polynomially many) queries are allowed to the respective side. Consider the following security game $\mathsf{Game}^{\mathsf{lrPRF\text{-}ODH}}_{\mathsf{PRF},\mathcal{A}}$ between a challenger $\mathcal{C}$ and a probabilistic polynomial-time (PPT) adversary $\mathcal{A}$.

1. *The challenger $\mathcal{C}$ samples $u \xleftarrow{\$} \mathbb{Z}_q$ and provides $\mathbb{G}$, g, and g^u to the adversary $\mathcal{A}$.*
2. *If $\mathsf{l} = \mathsf{m}$, $\mathcal{A}$ can issue arbitrarily many queries to the following oracle ODH_u.*

   ODH_u **oracle.** *On a query of the form (S, x), the challenger first checks if $S \notin \mathbb{G}$ and returns $\perp$ if this is the case. Otherwise, it computes $y \leftarrow \mathsf{PRF}(S^u, x)$ and returns y.*
3. *Eventually, $\mathcal{A}$ issues a challenge query $x^\star$. On this query, $\mathcal{C}$ samples $v \xleftarrow{\$} \mathbb{Z}_q$ and a bit $b \xleftarrow{\$} \{0,1\}$ uniformly at random. It then computes $y_0^\star = \mathsf{PRF}(g^{uv}, x^\star)$ and samples $y_1^\star \xleftarrow{\$} \{0,1\}^\lambda$ uniformly random. The challenger returns $(g^v, y_b^\star)$ to $\mathcal{A}$.*
4. *Next, $\mathcal{A}$ may issue (arbitrarily interleaved) queries to the following oracles ODH_u and ODH_v (depending on l and r).*

   ODH_u **oracle.** *The adversary $\mathcal{A}$ may ask no ($\mathsf{l} = \mathsf{n}$), a single ($\mathsf{l} = \mathsf{s}$), or arbitrarily many ($\mathsf{l} = \mathsf{m}$) queries to this oracle. On a query of the form (S, x), the challenger first checks if $S \notin \mathbb{G}$ or $(S, x) = (g^v, x^\star)$ and returns $\perp$ if this is the case. Otherwise, it computes $y \leftarrow \mathsf{PRF}(S^u, x)$ and returns y.*

   ODH_v **oracle.** *The adversary $\mathcal{A}$ may ask no ($\mathsf{r} = \mathsf{n}$), a single ($\mathsf{r} = \mathsf{s}$), or arbitrarily many ($\mathsf{r} = \mathsf{m}$) queries to this oracle. On a query of the form (T, x), the challenger first checks if $T \notin \mathbb{G}$ or $(T, x) = (g^u, x^\star)$ and returns $\perp$ if this is the case. Otherwise, it computes $y \leftarrow \mathsf{PRF}(T^v, x)$ and returns y.*
5. *At some point, $\mathcal{A}$ stops and outputs a guess $b' \in \{0,1\}$.*

We say that the adversary wins the $\mathsf{lrPRF\text{-}ODH}$ game if $b' = b$ and define the advantage function

$$\mathsf{Adv}^{\mathsf{lrPRF\text{-}ODH}}_{\mathsf{PRF},\mathcal{A}}(\lambda) := 2 \cdot \left(\Pr[b' = b] - \frac{1}{2}\right)$$

and, assuming a sequence of groups in dependency of the security parameter, we say that a pseudorandom function PRF with keys from $(\mathbb{G}_\lambda)_\lambda$ provides $\mathsf{lrPRF\text{-}ODH}$ security (for $\mathsf{l}, \mathsf{r} \in \{\mathsf{n}, \mathsf{s}, \mathsf{m}\}$) if for any $\mathcal{A}$ the advantage $\mathsf{Adv}^{\mathsf{lrPRF\text{-}ODH}}_{\mathsf{PRF},\mathcal{A}}(\lambda)$ is negligible in the security parameter λ.

In the following, if clear from the context, we will omit the group $\mathbb{G}$ and sometimes its generator g as explicit inputs to the adversary.

Relations to previous $\mathsf{PRF\text{-}ODH}$ assumptions. The above generic and parameterized $\mathsf{lrPRF\text{-}ODH}$ definition captures different variants of the PRF-ODH assumption present in the literature. The $\mathsf{PRF\text{-}ODH}$ formulation put forward by Jager et al. [JKSS12] is captured by ours in case the parameters are set to $\mathsf{l} = \mathsf{s}$ and $\mathsf{r} = \mathsf{n}$ meaning that only the "left" oracle (querying the DH share g^u) can be queried once. Note that Step 2 is only required if $\mathsf{l} = \mathsf{m}$, capturing that Jager et al. first request their challenge before issuing an oracle query. The same variant, $\mathsf{snPRF\text{-}ODH}$, was also used by Dowling et al. [DFGS16]. Krawczyk et al. [KPW13] modified the $\mathsf{PRF\text{-}ODH}$ formulation of Jager et al. since they require security against multiple ("left") oracle queries against the DH key share. Thus, their variant is captured by ours through setting the parameters to $\mathsf{l} = \mathsf{m}$ and $\mathsf{r} = \mathsf{n}$, and thus making use of Step 2. Recent works further introduced an additional query to the other DH key share, due to the fact that the keys are static or semi-static, respectively. In more detail, Fischlin and Günther [FG17] require an extra single ("right") oracle query while still requesting polynomial many queries to the "left" oracle. This is captured by our definition through setting the parameters to $\mathsf{l} = \mathsf{m}$ and $\mathsf{r} = \mathsf{s}$. Lastly, Brendel and Fischlin [BF17] require to query both key shares multiple times, which our definition captures as well by choosing the parameters as $\mathsf{l} = \mathsf{m}$ and $\mathsf{r} = \mathsf{m}$.

Design options. The above generic definition can be refined further, e.g., by enabling the challenger to provide the value g^v to the adversary at the outset in Step 1. This variant was used in the analysis of earlier TLS 1.3 draft handshakes by Dowling et al. [DFGS15b]. Such change would be accompanied by giving the adversary in Step 2 also access to the ODH_v oracle in case $\mathsf{r} = \mathsf{m}$. Another reasonable change could encompass enabling the adversary in multi-query variants (i.e., $\mathsf{l} = \mathsf{m}$ or $\mathsf{r} = \mathsf{m}$) to also issue *multiple* challenge queries in Step 3, for the same value g^v or even freshly chosen values g^{v_i} in each call. However, one can show via a standard hybrid argument that both notions (i.e., single challenge query and multiple challenge query) are polynomially equivalent.

In this work, we focus on the common structure of previously studied $\mathsf{PRF\text{-}ODH}$ notions [JKSS12, KPW13, DFGS16, BF17, FG17] which are captured by our generic definition above. Additionally, in Section 4 we briefly discuss the impact of such changes regarding the analysis of the relations between the different variants of the assumption.

3 Instantiating the PRF-ODH Assumption

We next turn to the question how to instantiate the $\mathsf{PRF\text{-}ODH}$ assumption. Concretely, we provide instantiations of the two notions that mark both ends of the strength spectrum of the PRF-ODH variants. First, we show that the weakest PRF-ODH variant, $\mathsf{nnPRF\text{-}ODH}$, can be instantiated in the standard model under well-established assumptions, namely the Decisional Diffie–Hellman (DDH) assumption and (ordinary) PRF security in a group $\mathbb{G}$. Second, we establish that, in the (programmable) random oracle model, both the strongest *one-sided* PRF-ODH variants, $\mathsf{mnPRF\text{-}ODH}$ and $\mathsf{nmPRF\text{-}ODH}$, as well as the most general $\mathsf{mmPRF\text{-}ODH}$ assumption can be instantiated from the strong Diffie–Hellman assumption (StDH). We define all these number-theoretic assumptions when discussing the security notions. Furthermore, we discuss the relation of the PRF-ODH notion to that of PRF security under related-key attacks.

3.1 Standard-Model Instantiation of $\mathsf{nnPRF\text{-}ODH}$

We begin with instantiating the $\mathsf{nnPRF\text{-}ODH}$ assumption in the standard model. For this we speak of a function $\mathsf{F}\colon \mathbb{G} \times \{0,1\}^* \to \{0,1\}^\lambda$ to be $\mathsf{PRF}_{\mathbb{G}}$-secure if no efficient adversary which, upon querying x, gets to see either the function value $\mathsf{F}(K, x)$ for a then chosen random key $K \xleftarrow{\$} \mathbb{G}$, or a random value, can distinguish the two cases. As in the other games before, the choice of answering genuinely or randomly is made at random, and we let $\mathsf{Adv}^{\mathsf{PRF}_{\mathbb{G}}}_{\mathsf{F},\mathcal{A}}$ denote the advantage of algorithm $\mathcal{A}$. Here, we normalize again the advantage by subtracting the guessing probability of $\frac{1}{2}$ and multiplying the result by a factor of 2. Note that the difference to the $\mathsf{nnPRF\text{-}ODH}$ assumption is that the adversary does not get to see a pair g^u, g^v from which the key is generated.

The underlying DDH assumption says that one cannot efficiently distinguish tuples (g, g^u, g^v, g^{uv}) from tuples (g, g^u, g^v, g^z) for random $u, v, z \in \mathbb{Z}_q$. More formally, for an adversary $\mathcal{B}$ we define $\mathsf{Adv}^{\mathsf{DDH}}_{\mathbb{G},\mathcal{B}}$ to be the probability of $\mathcal{B}$ predicting a random bit b, when given g, g^u, g^v, g^{uv} for $b = 0$ and g, g^u, g^v, g^z for $b = 1$, with the usual normalization as above. Alternatively, one may define $\mathsf{Adv}^{\mathsf{DDH}}_{\mathbb{G},\mathcal{B}}$ to be the advantage in the $\mathsf{nnPRF\text{-}ODH}$ game for the function $\mathsf{F}(K, x) = K$.

Theorem 3.1 ($\mathsf{DDH} + \mathsf{PRF}_{\mathbb{G}} \implies \mathsf{nnPRF\text{-}ODH}$). *If a function* $\mathsf{F} \colon \mathbb{G} \times \{0,1\}^* \to \{0,1\}^\lambda$ *is* $\mathsf{PRF}_{\mathbb{G}}$*-secure and the* DDH *assumption holds in* $\mathbb{G}$*, then* F *is also* $\mathsf{nnPRF\text{-}ODH}$*-secure. More precisely, for any efficient adversary* $\mathcal{A}$ *against the* $\mathsf{nnPRF\text{-}ODH}$ *security of* F*, there exist efficient algorithms* $\mathcal{B}_1$ *and* $\mathcal{B}_2$ *such that*

$$\mathsf{Adv}^{\mathsf{nnPRF\text{-}ODH}}_{\mathsf{F},\mathcal{A}} \leq 2 \cdot \mathsf{Adv}^{\mathsf{DDH}}_{\mathbb{G},\mathcal{B}_1} + 2 \cdot \mathsf{Adv}^{\mathsf{PRF}_{\mathbb{G}}}_{\mathsf{F},\mathcal{B}_2}.$$

We note that the factor 2 is the common loss due to the game-hopping technique, when switching from indistinguishability for two fixed games to choosing one of the games at random.

Proof. Let $\mathcal{A}$ be an efficient adversary against the $\mathsf{nnPRF\text{-}ODH}$ security of F. We show that $\mathcal{A}$'s advantage $\mathsf{Adv}^{\mathsf{nnPRF\text{-}ODH}}_{\mathsf{F},\mathcal{A}}$ in winning the $\mathsf{nnPRF\text{-}ODH}$ game is bounded by $\mathcal{B}_1$'s advantage $\mathsf{Adv}^{\mathsf{DDH}}_{\mathbb{G},\mathcal{B}_1}$ against DDH, and $\mathcal{B}_2$'s advantage $\mathsf{Adv}^{\mathsf{PRF}_{\mathbb{G}}}_{\mathsf{F},\mathcal{B}_2}$ against the $\mathsf{PRF}_{\mathbb{G}}$-security of F. This is done via game-hopping:

Game 0. The original $\mathsf{nnPRF\text{-}ODH}$ game.

Game 1. As the original game, but we replace the key g^{uv} used in computing the challenge value $y_0^\star$ by an independent random group element $g^z \in \mathbb{G}$. We claim that $\mathcal{A}$ cannot distinguish Game 0 from Game 1 efficiently with non-negligible advantage, since otherwise there exists an efficient adversary $\mathcal{B}_1$ that can solve DDH with non-negligible advantage. So, assume that $\mathcal{A}$ is able to distinguish the two games. Then $\mathcal{B}_1$ is constructed as follows: In the DDH game, $\mathcal{B}_1$ receives its challenge, say (g, g^u, g^v, g^z). In order to decide whether $g^z = g^{uv}$ or $g^z \xleftarrow{\$} \mathbb{G}$, algorithm $\mathcal{B}_1$ runs $\mathcal{A}$ as a subroutine on input g, g^u. Then $\mathcal{B}_1$ answers $\mathcal{A}$'s challenge query $x^\star$ with $(g^v, y_b^\star)$ where $y_0^\star = \mathsf{F}(g^z, x^\star)$ and $y_1^\star \xleftarrow{\$} \{0,1\}^\lambda$, and the choice $b \xleftarrow{\$} \{0,1\}$ is made by $\mathcal{B}_1$. Eventually, $\mathcal{A}$ outputs a bit b', and $\mathcal{B}$ returns 0 if and only if $b = b'$.

Note, that if $\mathcal{A}$ can efficiently distinguish the games, i.e., detect whether $g^z = g^{uv}$ (Game 0) or not (Game 1), $\mathcal{B}_1$ can also efficiently solve its DDH challenge with the same advantage (times a factor $\frac{1}{2}$ for moving from a random choice of either giving $\mathcal{B}_1$ the value g^{uv} or a random value g^z to fixed games Game 0 and Game 1). Thus, we can bound the advantage by

$$\mathsf{Adv}^{\mathsf{nnPRF\text{-}ODH}}_{\mathsf{F},\mathcal{A}} \leq \mathsf{Adv}^{\mathsf{Game1}}_{\mathsf{F},\mathcal{A}} + 2 \cdot \mathsf{Adv}^{\mathsf{DDH}}_{\mathbb{G},\mathcal{B}_1}.$$

Game 2. As the previous game, but this time we replace the challenge value $y_0^\star$ itself by a uniform random value, i.e., $y_0^\star \xleftarrow{\$} \{0,1\}^\lambda$. We show that if there exists an efficient adversary $\mathcal{A}$ that can distinguish Game 2 from Game 1, then there necessarily exists an efficient algorithm $\mathcal{B}_2$ that can break the $\mathsf{PRF}_{\mathbb{G}}$ security of F. We construct $\mathcal{B}_2$ as follows: To initiate the environment for $\mathcal{A}$, algorithm $\mathcal{B}_2$ chooses some arbitrary group element g^u and forwards it to $\mathcal{A}$. At some point, $\mathcal{A}$ asks the challenge query $x^\star$, which $\mathcal{B}_2$ relays to its own challenger, receiving the $\mathsf{PRF}_{\mathbb{G}}$-challenge $y_b^\star$. Algorithm $\mathcal{B}_2$ forwards $y^\star$ along with an arbitrarily chosen group element g^v to $\mathcal{A}$. Eventually, $\mathcal{A}$ stops and outputs a bit b'. Algorithm $\mathcal{B}_2$ outputs the same bit as $\mathcal{A}$.

Note that $\hat{b} = 0$ corresponds to Game 1, whereas $\hat{b} = 1$ corresponds to the environment of Game 2. Therefore, if $\mathcal{A}$ can efficiently distinguish between the two games, $\mathcal{B}_2$ can distinguish between $\mathsf{PRF}_{\mathbb{G}}$ values and independent random values with the same advantage. Hence, we can bound $\mathcal{A}$'s advantage by

$$\mathsf{Adv}^{\mathsf{Game1}}_{\mathsf{F},\mathcal{A}} \leq \mathsf{Adv}^{\mathsf{Game2}}_{\mathsf{F},\mathcal{A}} + 2 \cdot \mathsf{Adv}^{\mathsf{PRF}_{\mathbb{G}}}_{\mathsf{F},\mathcal{B}_2}.$$

Since both $y_0^\star$ and $y_1^\star$ are now drawn independently and at random from $\{0,1\}^\lambda$, $\mathcal{A}$ cannot do better than guessing, i.e., $\mathsf{Adv}^{\mathsf{Game2}}_{\mathsf{F},\mathcal{A}} = 0$, which completes the proof. □

3.2 Random-Oracle Instantiation of mnPRF-ODH and nmPRF-ODH

Abdalla et al. [ABR01] proved that the oracle DH assumption ODH is implied by the strong Diffie–Hellman assumption in the random oracle model. Here, we show that our strongest *one-sided* PRF-ODH variants, mnPRF-ODH and nmPRF-ODH, can be instantiated under the strong Diffie–Hellman assumption StDH . The assumption says that, given g, g^u, g^v and access to a decisional DH oracle for fixed value g^u, i.e., $\mathsf{DDH}(g^u, \cdot, \cdot)$, it is infeasible to compute g^{uv}. Observe that this assumption is implied by the GapDH assumption, where the adversary can choose the first group element freely, too. Let $\mathsf{Adv}^{\mathsf{StDH}}_{\mathbb{G},\mathcal{B}}$ denote the probability of algorithm $\mathcal{B}^{\mathsf{DDH}(g^u,\cdot,\cdot)}(g, g^u, g^v)$ outputting g^{uv}.

Theorem 3.2. *In the random oracle model,* StDH *implies* mnPRF-ODH *security and* nmPRF-ODH *security of* $\mathsf{F}(K, x) = \mathsf{RO}(K, x)$ *for random oracle* RO. *More precisely, for any efficient adversary* $\mathcal{A}$ *against the* mnPRF-ODH *or* nmPRF-ODH *security of* F, *there exists an efficient algorithms* $\mathcal{B}$ *such that*

$$\mathsf{Adv}^{\mathsf{mnPRF\text{-}ODH}}_{\mathsf{F},\mathcal{A}} \leq \mathsf{Adv}^{\mathsf{StDH}}_{\mathbb{G},\mathcal{B}} \qquad \textit{and} \qquad \mathsf{Adv}^{\mathsf{nmPRF\text{-}ODH}}_{\mathsf{F},\mathcal{A}} \leq \mathsf{Adv}^{\mathsf{StDH}}_{\mathbb{G},\mathcal{B}}.$$

We prove the following theorem for the case of mnPRF-ODH only; the case nmPRF-ODH follows analogously, noting that by symmetry of the inputs g^u, g^v we can assume that the adversary $\mathcal{B}$ gets access to a $\mathsf{DDH}(g^v, \cdot, \cdot)$ oracle.

Proof. The proof is by reduction. We describe the construction of the reduction $\mathcal{B}$ from mnPRF-ODH to StDH. First of all, $\mathcal{B}$ obtains group elements g, g^u, g^v in the StDH game. To initiate the mnPRF-ODH game environment for $\mathcal{A}$, the StDH-adversary $\mathcal{B}$ forwards g and g^u to $\mathcal{A}$ as input. $\mathcal{A}$ now has access to the oracles RO and ODH_u, i.e., $\mathcal{A}$ may send queries of the form (S, x) to the ODH_u oracle, with $S \in \mathbb{G}$ and x being some bit string. To provide an appropriate simulation it must be ensured that, if $\mathcal{A}$ first queries some (S, x) to ODH_u and then (S^u, x) to the random oracle RO, the answer of RO is consistent with the simulation of ODH_u, and vice versa. This can be achieved if $\mathcal{B}$ can program the RO and has access to a $\mathsf{DDH}(g^u, \cdot, \cdot)$ oracle.[2] $\mathcal{B}$ simulates the two oracles as follows:

Simulation of RO. The answers of the random oracle RO need to be consistent, i.e., if a query is asked repeatedly, RO returns the same answer. This can be ensured by standard bookkeeping techniques. If a previously unseen query (K, x) is received, $\mathcal{B}$ must consider the case that $\mathcal{A}$ has already queried (S, x) with $K = S^u$ to ODH_u. Thus, when receiving a call (K, x) to the RO, algorithm $\mathcal{B}$ queries its $\mathsf{DDH}(g^u, \cdot, \cdot)$ oracle on (S, K) for any group element S that has been queried with x to ODH_u. If the DDH oracle returns 1 on any such input then $\mathcal{B}$ answers consistently with the corresponding answer from earlier. Else, $\mathcal{B}$ assigns a fresh value y to (K, x) and returns y to $\mathcal{A}$.

[2] We remark that this is the reason why, in the given scenario, the computational DH assumption is not sufficient.

Simulation of ODH_u. Analogously to the simulation of the random oracle, $\mathcal{B}$ checks each newly received request by $\mathcal{A}$ against all previous query-response pairs of ODH_u and answers consistently in case of repetition. If a previously unseen query (S, x) is received by ODH_u, $\mathcal{B}$ must further check whether the related value (S^u, x) has been queried to RO before. Similar to the reverse case, $\mathcal{B}$ uses its $\mathsf{DDH}(g^u, \cdot, \cdot)$ oracle on (S, K) on all previous RO queries (K, x) to detect this. If $\mathsf{DDH}(g^u, S, K) = 1$ for some K, the simulation of ODH_u answers with the respective output of RO. Otherwise, a response y is drawn uniformly at random from $\{0,1\}^\lambda$ and returned to $\mathcal{A}$ (and the tuple (S, x, y) stored for future reference).

At some point, $\mathcal{A}$ issues a challenge query $x^\star$ to its challenger. Algorithm $\mathcal{B}$ answers this query with g^v and some value $y^\star$, drawn uniformly at random from $\{0,1\}^\lambda$. Adversary $\mathcal{A}$ can now query ODH_u and RO further, with the limitation that it may not query the pair $(g^v, x^\star)$ to ODH_u. These queries are simulated as before. Eventually, $\mathcal{A}$ stops and outputs a guess bit b'. Then, $\mathcal{B}$ queries $\mathsf{DDH}(g^u, g^v, K)$ on all queries $(K, x^\star)$ of $\mathcal{A}$ to the RO. If $\mathsf{DDH}(g^u, g^v, K) = 1$ for some RO-query $(K, x^\star)$, then $\mathcal{B}$ outputs K in the StDH game.

If the efficient adversary $\mathcal{A}$ wins the $\mathsf{mnPRF\text{-}ODH}$ game with non-negligible advantage, then $\mathcal{B}$ also outputs the correct value g^{uv} with non-negligible probability. To see this, we note that $\mathcal{A}$ can only win the $\mathsf{mnPRF\text{-}ODH}$ game in the random oracle model with non-negligible advantage if g^{uv} appears in one of its RO queries. Assume that this is not the case. Then $\mathcal{A}$ expects $y^\star$ to be either $y_0 = \mathsf{RO}(g^{uv}, x^\star)$ or $y_1 \xleftarrow{\$} \{0,1\}^\lambda$. By the nature of random oracles, y_0 and y_1 are indistinguishable for $\mathcal{A}$ since both are drawn uniformly at random from $\{0,1\}^\lambda$. This holds even if $\mathcal{A}$ could correctly determine the value g^{uv}, since it cannot compute y_0 without querying the random oracle to compare with the received challenge.

Thus, $\mathcal{A}$ must necessarily query g^{uv} to the random oracle in order to distinguish y_0 and y_1. If this happens in the actual attack (with a "genuine" random oracle) with non-negligible probability, then it must also happen in the simulation above with non-negligible probability, because $\mathcal{B}$ provides a perfectly sound simulation of a random oracle. But if $\mathcal{A}$ makes such a query (in the simulation) then $\mathcal{B}$ finds the DH value in the list of queries and correctly outputs it. Furthermore, $\mathcal{B}$ is efficient, since $\mathcal{A}$ is efficient and asks at most polynomially many (in the security parameter) queries to each oracle. □

3.3 Random-Oracle Instantiation of mmPRF-ODH

We next look at the case that the adversary can make queries to both oracles, ODH_u and ODH_v. Interestingly, this does not follow straightforwardly from the StDH assumption as above. The reason is that, there, we have used the DDH-oracle with fixed element g^u to check for consistency of ODH_u queries with random oracle queries. In the most general $\mathsf{mmPRF\text{-}ODH}$ case, however, we would also need to check consistency across ODH_u and ODH_v queries. In particular, a simulator would need to be able to check for queries (S, x) to ODH_u and (T, x) to ODH_v if they result in the same key $S^u = K = T^v$, but the simulator is given only S, T, g, g^u, and g^v. Such a test cannot be immediately performed with the $\mathsf{DDH}(g^u, \cdot, \cdot)$ oracle as in the StDH case, and not even with the more liberal $\mathsf{DDH}(\cdot, \cdot, \cdot)$ oracle as in the GapDH case.

Suppose that we take the StDH problem and augment it by another oracle which allows to check for "claws" S, T with $S^u = T^v$. Call this the claw-verifying oracle Claw and the problem the $\mathsf{ClawStDH}$ problem. For pairing-friendly groups $\mathbb{G}$ we get this oracle for free via the bilinear map e as $\mathsf{Claw}(S, T) = [e(g^u, S) = e(g^v, T)?]$. Next, we show that for general groups the claw-verifying oracle can be implemented in the StDH game, too, but at the cost of a loose security reduction to StDH.

The idea of representing the oracle Claw is as follows. Suppose that, in addition to g, g^u and g^v we would also receive the value $g^{u/v}$ (where we assume here and in the following that $v \neq 0$, since the case $v = 0$ is trivial to deal with). Then we can run the check for claws via the stronger DDH oracle by calling $\mathsf{DDH}(g^{u/v}, S, T)$, checking that $S^{u/v} = T$ and therefore $S^u = T^v$. The question remains if the computational problem of computing g^{uv} given $g^{u/v}$ (in the presence of a DDH oracle) becomes significantly easier, and if we can relax the requirement to a $\mathsf{DDH}(g^u, \cdot, \cdot)$ oracle. Switching to the square DH problem in an intermediate step, we show that this is not the case, although the intermediate step causes a loose security relationship.

Assume that we have an algorithm $\mathcal{A}$ which (given oracle access to $\mathsf{DDH}(g^u,\cdot,\cdot)$, $\mathsf{DDH}(g^v,\cdot,\cdot)$, and the claw-verifying oracle Claw) on input (g,g^u,g^v) is able to compute g^{uv}. Then we show that we can use this algorithm to build an algorithm $\mathcal{B}$ for the square-DH problem (given g,g^v compute g^{v^2}) relative to a $\mathsf{DDH}(g^v,\cdot,\cdot)$ oracle. For this, algorithm $\mathcal{B}$ for input g,g^v picks $r \xleftarrow{\$} \mathbb{Z}_q$ and sets $g^u = (g^v)^r$. With this choice, $g^{u/v} = g^r$ can be easily computed with the knowledge of r, allowing to implement the claw-verifying oracle for free. Similarly, we have $\mathsf{DDH}(g^u,\cdot,\cdot) = \mathsf{DDH}(g^v,(\cdot)^r,\cdot)$, giving us the "mirrored" oracle for free. Algorithm $\mathcal{B}$ now runs $\mathcal{A}$ on input (g,g^u,g^v) and answers all oracle requests of $\mathcal{A}$ during the computation with the help of its $\mathsf{DDH}(g^v,\cdot,\cdot)$ oracle. Suppose that the adversary $\mathcal{A}$ eventually outputs K. Then, $\mathcal{B}$ returns $K^{1/r}$ which equals g^{v^2} for a correct answer $K = g^{uv} = g^{rv^2}$ of $\mathcal{A}$.

Next, we show that from a solver for the square-DH problem (with $\mathsf{DDH}(g^v,\cdot,\cdot)$ oracle) we can build a solver for the StDH problem. Going from the square-DH problem to the CDH problem is already known. Interestingly, though, the common strategies in the literature [MW96, BDZ03, Gal12] require three calls to the square-DH solver, basically to compute the square $g^{(u+v)^2} = g^{u^2+2uv+v^2}$ and then to divide out g^{u^2} and g^{v^2}. Fortunately, two calls are sufficient, see for example [Kil01], yielding a tighter security bound.

So suppose we have a square-DH algorithm (with oracle $\mathsf{DDH}(g^v,\cdot,\cdot)$) then we call this algorithm once on g,g^{u+v} and once on $g,g^{r(u-v)}$ for randomizer $r \xleftarrow{\$} \mathbb{Z}_q$. Since both inputs are random and independent, we get two valid answers $g^{u^2+2uv+v^2}$ and $g^{r^2(u^2-2uv+v^2)}$ with the product of the square-DH algorithm's success probability. Note that these two executions at most double the number of oracle queries to the DDH oracle. Dividing out the exponent r^2 from the second term by raising it to the power $1/r^2$, and then dividing the two group elements we obtain g^{4uv} from which we can easily compute g^{uv}.

Overall, we can show that solving the problem in presence of the decisional oracles for g^u and g^v, and an additional claw-verifying oracle, is implied by the StDH assumption, albeit with a security loss. More precisely, for any efficient adversary $\mathcal{A}$ against $\mathsf{ClawStDH}$ we get an efficient adversary $\mathcal{B}$ (making at most twice as many calls to its StDH oracle as $\mathcal{A}$) such that

$$\mathsf{Adv}^{\mathsf{ClawStDH}}_{\mathbb{G},\mathcal{A}} \leq \sqrt{\mathsf{Adv}^{\mathsf{StDH}}_{\mathbb{G},\mathcal{B}}}.$$

We can now give our security proof for $\mathsf{mmPRF\text{-}ODH}$, implying also security of $\mathsf{msPRF\text{-}ODH}$ and $\mathsf{smPRF\text{-}ODH}$, of course:

Theorem 3.3. *In the random oracle model,* $\mathsf{ClawStDH}$ *(resp.* StDH*) implies*$\mathsf{mmPRF\text{-}ODH}$ *security of* $\mathsf{F}(K,x) = \mathsf{RO}(K,x)$ *for random oracle* RO*. More precisely, for any efficient adversary* $\mathcal{A}$ *against the* $\mathsf{mmPRF\text{-}ODH}$ *security of* F*, there exist efficient algorithms* $\mathcal{B}_1,\mathcal{B}_2$ *such that*

$$\mathsf{Adv}^{\mathsf{mmPRF\text{-}ODH}}_{\mathsf{F},\mathcal{A}} \leq \mathsf{Adv}^{\mathsf{ClawStDH}}_{\mathbb{G},\mathcal{B}_1} \leq \sqrt{\mathsf{Adv}^{\mathsf{StDH}}_{\mathbb{G},\mathcal{B}_2}}$$

Proof. The proof is almost identical to the one for $\mathsf{mnPRF\text{-}ODH}$, only that we here simulate the other oracle ODH_v as the oracle ODH_u, and for each query to either of the oracles also check via the help of Claw consistency between ODH_u and ODH_v evaluations. This provides a sound simulation of the random oracle. It follows as before that the adversary $\mathcal{A}$ can only distinguish genuine $y^\star$ from random ones if it queries the random oracle about g^{uv} (in the sound simulation), in which case $\mathcal{B}_1$ finds this value in the list of queries. □

3.4 On the Relation Between PRF-ODH and Security Against Related-key Attacks

The PRF-ODH assumption demands the output of a PRF to be indistinguishable from random even when given access to PRF evaluations under a related (group-element) key, sharing (at least) one exponent of the challenge key. On a high level, this setting resembles the concept of *related-key attack (RKA) security* for pseudorandom functions as introduced by Bellare and Kohno [BK03]. This raises the question if the PRF-ODH assumption can be instantiated from RKA-secure PRFs (or vice versa).

Related-key attack security of a PRF $f\colon \mathcal{K} \times \mathcal{D} \to \mathcal{R}$ with respect to a set Φ of related-key-deriving (RKD) functions is defined as the indistinguishability of two oracles $F_{(\cdot,K)}(\cdot)$ and $G_{(\cdot,K)}(\cdot)$ for a randomly chosen key $K \xleftarrow{\$} \mathcal{K}$. The distinguishing adversary $\mathcal{A}$ may query the oracles on inputs $(\phi, x) \in \Phi \times \mathcal{D}$ on which the oracles respond as $F_{(\phi,K)}(x) := f(\phi(K), x)$ and $G_{(\phi,K)}(x) := g(\phi(K), x)$ for a function g drawn uniformly at random from the set $\mathcal{FF}(\mathcal{K}, \mathcal{D}, \mathcal{R})$ of all functions $\mathcal{K} \times \mathcal{D} \to \mathcal{R}$. Formally, the advantage of $\mathcal{A}$ against the RKA-PRF security of f with respect to set Φ is defined as

$$\begin{aligned}\mathsf{Adv}^{\mathsf{RKA\text{-}PRF},\Phi}_{f,\mathcal{A}} := &\Pr\left[\mathcal{A}^{F_{(\cdot,K)}(\cdot)} = 1 \mid K \xleftarrow{\$} \mathcal{K}\right] \\ &- \Pr\left[\mathcal{A}^{G_{(\cdot,K)}(\cdot)} = 1 \mid K \xleftarrow{\$} \mathcal{K}, g \xleftarrow{\$} \mathcal{FF}(\mathcal{K}, \mathcal{D}, \mathcal{R})\right].\end{aligned}$$

Intuitively, one should now be able to relate $\mathsf{RKA\text{-}PRF}$ security to $\mathsf{PRF\text{-}ODH}$ security by considering two correlated sets of RKD functions corresponding to the PRF-ODH oracles ODH_u and ODH_v with respect to a group $\mathbb{G}$ with generator g and two exponents $u, v \in \mathbb{Z}_q$:

$$\begin{aligned}\Phi_{\mathsf{ODH}_u} &:= \{\phi_{\mathsf{ODH}_u,S} \mid S \in \mathbb{G} \setminus \{g^v\}\} &\quad \text{where } \phi_{\mathsf{ODH}_u,S}(K) := (K^{1/v})^{\log_g(S)}, \\ \Phi_{\mathsf{ODH}_v} &:= \{\phi_{\mathsf{ODH}_v,T} \mid T \in \mathbb{G} \setminus \{g^u\}\} &\quad \text{where } \phi_{\mathsf{ODH}_v,T}(K) := (K^{1/u})^{\log_g(T)}.\end{aligned}$$

Insurmountable hurdles however seem to remain when trying to relate PRF-ODH notions and RKA-PRF security (for according sets Φ) via implications. In the one direction, the adversary in the PRF-ODH setting is provided with the DH shares g^u and g^v forming the (challenge) PRF key while such side information on the key is not given in the RKA-PRF setting. Hence, in a reduction of PRF-ODH security to some RKA-PRF notion, even for an appropriate RKD function set a simulation always lacks means to provide the PRF-ODH adversary with these shares. In the other direction, the RKA-PRF challenge can be issued on any related key $\phi(K)$ for an admissible RKD function ϕ while the PRF-ODH challenge is, for the case of the real PRF response, always computed on the key g^{uv}. A reduction would hence need to map the RKA-PRF challenge for an arbitrary, related key onto the fixed PRF-ODH challenge key.

Though on a high level capturing a relatively similar idea, the relation between PRF-ODH and RKA-PRF security hence remains an open question.

4 PRF-ODH Relations

In this section we study the relations of different $\mathsf{PRF\text{-}ODH}$ variants spanned by our generic Definition 2.1. The relationships are also illustrated in Figure 2.

Let us start with observing the trivial implications (indicated by solid arrows in Figure 2) which are induced by restricting the adversary's capabilities in our definition. That is, by restricting the access to one of the oracles ODH_u and ODH_v (from multiple queries to a single query or from a single query to no query) for a notion from Definition 2.1, we obtain a trivially weaker variant. The more interesting question is which of these implications are strict, i.e., for which of two $\mathsf{PRF\text{-}ODH}$ variant pairs one notion is *strictly* stronger than the other. For a majority of these cases we can give separations which only require the assumption that the underlying primitive exists at all, for some separations we rely on the random oracle model (and a plausible number-theoretic assumption).

4.1 Separations in the Standard Model

For our standard model separations we introduce the following family of functions $\mathcal{F}$.

Definition 4.1 (Separating function family $\mathcal{F}$)**.** *Let* $\mathsf{G}\colon \mathbb{G} \times \{0,1\}^* \to \{0,1\}^\lambda$*. We define the family of functions* $\mathcal{F} = \{\mathsf{F}_n\}_{n\in\mathbb{N}}$ *with* $\mathsf{F}_n\colon \mathbb{G} \times \{0,1\}^* \to \{0,1\}^\lambda$ *as follows:*

$$\mathsf{F}_n(K, x) := \begin{cases} \mathsf{G}(K,1) \oplus \ldots \oplus \mathsf{G}(K,n) & \textit{if } x = 0 \\ \mathsf{G}(K,x) & \textit{otherwise.} \end{cases}$$

As a warm-up, let us first consider the (in)security of functions $\mathsf{F}_n \in \mathcal{F}$ in the standard PRF setting. It is easy to see that no function $\mathsf{F}_n \in \mathcal{F}$ can satisfy the (regular) security notion for pseudorandom functions: for any function F_n, querying the PRF oracle on $x_0 = 0, \ldots, x_n = n$ yields responses $y_0, \ldots, y_n$ for which the combined XOR value $y = y_0 \oplus \ldots \oplus y_n$, in case the oracle computes the real function F_n, is always 0 whereas otherwise it is 0 only with probability $2^{-\lambda}$. However, in a restricted setting where the PRF adversary $\mathcal{A}$ is allowed to query the oracle only a limited number of times (at most n queries for function F_n), we can indeed establish the following, restricted PRF security for functions $\mathsf{F}_n \in \mathcal{F}$.

Proposition 4.2 ($\mathcal{F}$ is restricted-PRF-secure)**.** *If* G *is an (ordinary) secure pseudorandom function, then each* $\mathsf{F}_n \in \mathcal{F}$ *from Definition 4.1 is an* n*-restricted secure pseudorandom function in the sense that it provides PRF security against any adversary that is allowed to query the PRF oracle at most* n *times.*

Proof (informal). Fix a function $\mathsf{F}_n \in \mathcal{F}$. First, we replace G in the definition of F_n by a truly random function G'. The introduced advantage difference for adversary $\mathcal{A}$ by this step can be bounded by the advantage of an adversary $\mathcal{B}$ against the PRF security of G, simulating the (restricted) PRF game for $\mathcal{A}$ using its own PRF oracle for G.

After this change, the output values of F_n on inputs $x > 1$ are independent random values and the output on $x = 0$ is the XOR of the outputs on $x = 1, \ldots, n$. In contrast, for a truly random function, the outputs on all inputs (incl. $x = 0$) are independent and random. However, any adversary $\mathcal{A}$ that is allowed to query the PRF oracle on at most n inputs cannot distinguish these two cases, bounding its success probability at this point by 0. □

Let us now turn to the more involved PRF-ODH setting. Equipped with the function family $\mathcal{F}$, we can establish separations between various PRF-ODH variants, as illustrated in Figure 2. The key insight for these separations is similar to the one in the standard PRF setting: an adversary with a limited number of n queries (including the challenge query in the PRF-ODH setting) cannot distinguish (a challenge under) F_n from (a challenge under) a truly random function. As subsequent propositions establish, this allows us to separate the notion nnPRF-ODH (with only one challenge query) from snPRF-ODH and nsPRF-ODH (with two queries, the challenge and one to an ODH oracle) via function F_1. Furthermore, the notions snPRF-ODH and nsPRF-ODH (with two queries) are separated from mnPRF-ODH, ssPRF-ODH, and nmPRF-ODH (with three or polynomially many queries) via F_2. Finally, we establish that the notion ssPRF-ODH (three queries) can be separated from mnPRF-ODH and nmPRF-ODH (multiple queries) using function F_3. Note that functions $\mathsf{F}_n \in \mathcal{F}$ cannot provide a separation between two notions that both allow polynomially many queries (e.g., mnPRF-ODH and msPRF-ODH). To keep the propositions compact, the given separations constitute the minimal spanning set; recall that if a notion A implies another notion B, separating a notion C from B also separates C from A.

We begin with separating nnPRF-ODH from snPRF-ODH and nsPRF-ODH security.

Proposition 4.3 (nnPRF-ODH $\not\Longrightarrow$ snPRF-ODH, nsPRF-ODH)**.** *If* G *from Definition 4.1 is* nnPRF-ODH-*secure, then* $\mathsf{F}_1 \in \mathcal{F}$ *is* nnPRF-ODH-*secure, but neither* snPRF-ODH- *nor* snPRF-ODH-*secure. More precisely, for any efficient adversary* $\mathcal{A}$ *against the* nnPRF-ODH *security of* F_1*, there exists an efficient algorithm* $\mathcal{B}$ *such that*

$$\mathsf{Adv}^{\mathsf{nnPRF\text{-}ODH}}_{\mathsf{F}_1,\mathcal{A}} \leq \mathsf{Adv}^{\mathsf{nnPRF\text{-}ODH}}_{\mathsf{G},\mathcal{B}},$$

but there exist algorithms $\mathcal{A}_1$*,* $\mathcal{A}_2$ *with non-negligible advantage* $\mathsf{Adv}^{\mathsf{snPRF\text{-}ODH}}_{\mathsf{F}_1,\mathcal{A}_1} = \mathsf{Adv}^{\mathsf{nsPRF\text{-}ODH}}_{\mathsf{F}_1,\mathcal{A}_2} = 1 - 2^{-\lambda}$.

Proof. First, observe the following $\mathsf{snPRF\text{-}ODH}$-adversary $\mathcal{A}_1$ and $\mathsf{nsPRF\text{-}ODH}$-adversary $\mathcal{A}_2$ are successful (except with negligible probability). Both first challenge F_1 on $x^\star = 0$ (obtaining as $y^\star$ either $y_0^\star = \mathsf{G}(g^{uv}, 1)$ or $y_1^\star \xleftarrow{\$} \{0,1\}^\lambda$), then query $(g^v, 1)$ resp. $(g^u, 1)$ to their ODH_u resp. ODH_v oracle, obtaining a value $y = \mathsf{G}(g^{uv}, 1)$. They distinguish the challenge by outputting 0 if $y^\star = y$ and 1 otherwise and win except if coincidentally $y_1^\star = y$, which happens with probability $2^{-\lambda}$.

To see that F_1 is $\mathsf{nnPRF\text{-}ODH}$-secure if G is, consider an algorithm $\mathcal{B}$ that simply relays its obtained value g^u to $\mathcal{A}$ and the challenge query of $\mathcal{A}$ to its challenger unmodified if $x^\star \neq 0$, but for $x^\star = 0$ asks its challenge query on 1. Forwarding the response and outputting the same bit b' as $\mathcal{A}$ outputs, $\mathcal{B}$ provides a correct simulation for $\mathcal{A}$ and, moreover, wins if $\mathcal{A}$ does. □

We continue with the separation of $\mathsf{snPRF\text{-}ODH}$ from $\mathsf{mnPRF\text{-}ODH}$, $\mathsf{ssPRF\text{-}ODH}$, and $\mathsf{nmPRF\text{-}ODH}$ security.

Proposition 4.4 ($\mathsf{snPRF\text{-}ODH} \not\Longrightarrow \mathsf{mnPRF\text{-}ODH}, \mathsf{ssPRF\text{-}ODH}, \mathsf{nmPRF\text{-}ODH}$)**.** *If* G *from Definition 4.1 is* $\mathsf{mnPRF\text{-}ODH}$*-secure, then* $\mathsf{F}_2 \in \mathcal{F}$ *is* $\mathsf{snPRF\text{-}ODH}$*-secure, but neither* $\mathsf{mnPRF\text{-}ODH}$*-, nor* $\mathsf{ssPRF\text{-}ODH}$*-, nor* $\mathsf{nmPRF\text{-}ODH}$*-secure. More precisely, for any efficient adversary* $\mathcal{A}$ *against the* $\mathsf{snPRF\text{-}ODH}$ *security of* F_2*, there exist efficient algorithms* $\mathcal{B}_1, \ldots, \mathcal{B}_4$ *such that*

$$\mathsf{Adv}^{\mathsf{snPRF\text{-}ODH}}_{\mathsf{F}_2,\mathcal{A}} \leq \mathsf{Adv}^{\mathsf{mnPRF\text{-}ODH}}_{\mathsf{G},\mathcal{B}_1} + 4 \cdot \mathsf{Adv}^{\mathsf{mnPRF\text{-}ODH}}_{\mathsf{G},\mathcal{B}_2} + 4 \cdot \mathsf{Adv}^{\mathsf{mnPRF\text{-}ODH}}_{\mathsf{G},\mathcal{B}_3} + \mathsf{Adv}^{\mathsf{mnPRF\text{-}ODH}}_{\mathsf{G},\mathcal{B}_4},$$

but there exist algorithms $\mathcal{A}_1, \ldots, \mathcal{A}_3$ *with non-negligible advantage* $\mathsf{Adv}^{\mathsf{mnPRF\text{-}ODH}}_{\mathsf{F}_2,\mathcal{A}_1} = \mathsf{Adv}^{\mathsf{ssPRF\text{-}ODH}}_{\mathsf{F}_2,\mathcal{A}_2} = \mathsf{Adv}^{\mathsf{nmPRF\text{-}ODH}}_{\mathsf{F}_2,\mathcal{A}_3} = 1 - 2^{-\lambda}$*.*

Proof. It is immediate to see that F_2 is not $\mathsf{mnPRF\text{-}ODH}$-, $\mathsf{ssPRF\text{-}ODH}$-, or $\mathsf{nmPRF\text{-}ODH}$-secure. For this, consider respective adversaries $\mathcal{A}_1$, $\mathcal{A}_2$, $\mathcal{A}_3$ which first query the challenge $x^\star = 0$, obtaining, besides, g, g^v, a value $y^\star$ which equals either $y_0^\star = \mathsf{F}_2(g^{uv}, 0) = \mathsf{G}(g^{uv}, 1) \oplus \mathsf{G}(g^{uv}, 2)$ or $y_1^\star \xleftarrow{\$} \{0,1\}^*$. Then, $\mathcal{A}_1$ issues two queries $(g^v, 1)$ and $(g^v, 2)$ to its ODH_u oracle while $\mathcal{A}_2$ issues $(g^v, 1)$ to its ODH_u oracle and $(g^u, 2)$ to its ODH_v oracle, and $\mathcal{A}_3$ issues queries $(g^u, 1)$ and $(g^u, 2)$ to its ODH_v oracle, all of them obtaining values $y' = \mathsf{G}(g^{uv}, 1)$ and $y'' = \mathsf{G}(g^{uv}, 2)$. The adversaries then check whether $y^\star = y' \oplus y''$. If so, they output 0, else 1. Hence, $\mathcal{A}_1$, $\mathcal{A}_2$, and $\mathcal{A}_3$ always win the $\mathsf{mnPRF\text{-}ODH}$, $\mathsf{ssPRF\text{-}ODH}$, resp. $\mathsf{nmPRF\text{-}ODH}$ game for F_2, except when $y_1^\star = y' \oplus y''$, which happens with negligible probability $2^{-\lambda}$.

We can now focus on showing that F_2 is $\mathsf{snPRF\text{-}ODH}$-secure if G is $\mathsf{mnPRF\text{-}ODH}$-secure. For this, we separately consider the two distinct cases that $\mathcal{A}$ issues a challenge query for $x^\star > n$ and the case that $x^\star \leq n$.

In case $x^\star > n$, we can bound $\mathcal{A}$'s advantage, denoted as $\mathsf{Adv}^{\mathsf{snPRF\text{-}ODH}, x^\star > n}_{\mathsf{F}_2,\mathcal{A}}$, by the advantage of an adversary $\mathcal{B}_1$ against the $\mathsf{mnPRF\text{-}ODH}$ security of G as follows. Algorithm $\mathcal{B}_1$ provides its initially obtained values g, g^u as the initial input to $\mathcal{A}$. When $\mathcal{A}$ queries $x^\star$, $\mathcal{B}_1$ relays the challenge to its own challenge oracle, and forwards the obtained response $(g^v, y^\star)$ back to $\mathcal{A}$. When $\mathcal{A}$ issues its (sole) ODH_u query (S, x), $\mathcal{B}_1$ simply relays query and response in case $x \neq 0$. In case $x = 0$, $\mathcal{B}_1$ instead queries its ODH_u oracle on $(S, 1)$ and $(S, 2)$ and returns the combined XOR to $\mathcal{A}$. Finally, when $\mathcal{A}$ outputs its guess b', $\mathcal{B}_1$ stops and outputs the same guess b'. Observe that $\mathcal{B}_1$ is efficient (issuing 3 queries where $\mathcal{A}$ issues 2 queries) and provides a sound simulation for $\mathcal{A}$. Note in particular that, as $x^\star > n$, all of $\mathcal{B}_1$'s ODH_u queries (in particular on $x \leq n$) are permissible. As the challenge bit b coincides for $\mathcal{B}_1$'s $\mathsf{mnPRF\text{-}ODH}$ game and the simulated $\mathsf{snPRF\text{-}ODH}$ game for $\mathcal{A}$, $\mathcal{B}_1$ wins if $\mathcal{A}$ does, hence $\mathsf{Adv}^{\mathsf{snPRF\text{-}ODH}, x^\star > n}_{\mathsf{F}_2,\mathcal{A}} \leq \mathsf{Adv}^{\mathsf{mnPRF\text{-}ODH}}_{\mathsf{G},\mathcal{B}_1}$.

In case $x^\star \leq n$, we first modify the $\mathsf{snPRF\text{-}ODH}$ game for $\mathcal{A}$ by changing the function F_2 to a function F_2' which, for a truly random function R, the $\mathsf{snPRF\text{-}ODH}$ challenge group elements g^u, g^v, and some random, but fixed $i \xleftarrow{\$} \{1,2\}$ and $j \in \{1,2\}$ s.t. $j \neq i$, is defined as follows:

$$\mathsf{F}_2'(K,x) := \begin{cases} \mathsf{F}_2(K,x) & \text{if } K \neq g^{uv} \\ R(i) \oplus \mathsf{G}(K,j) & \text{if } K = g^{uv} \text{ and } x = 0 \\ R(i) & \text{if } K = g^{uv} \text{ and } x = i \\ \mathsf{G}(K,x) & \text{if } K = g^{uv} \text{ and } x \notin \{0,i\} \end{cases}$$

The difference in $\mathcal{A}$'s advantage introduced by this step can be bounded by the advantage of an adversary $\mathcal{B}_2$ against the $\mathsf{mnPRF\text{-}ODH}$ security of G as follows. Algorithm $\mathcal{B}_2$ obtains g^u and begins by picking $b^* \in \{0,1\}$ and $i \xleftarrow{\$} \{1,2\}$ at random. It asks $x^{\star\prime} = i$ as its challenge and obtains $(g^v, y^{\star\prime})$ where $y^{\star\prime}$ equals either ${y^{\star}}'_0 = \mathsf{G}(g^{uv}, i)$ or ${y^{\star}}'_1 \xleftarrow{\$} \{0,1\}^\lambda$. It then provides $\mathcal{A}$ with g^u and responds to $\mathcal{A}$'s queries as follows, dependent on $\mathcal{A}$'s challenge query:

1. $\mathcal{A}$ asks the challenge $x^\star = 0$. In this case, $\mathcal{B}_2$ queries (g^v, j), obtains y as response and computes $y_0^\star = y^{\star\prime} \oplus y$ and $y_1^\star \xleftarrow{\$} \{0,1\}^\lambda$. It returns $y_{b^*}^\star$ to $\mathcal{A}$.[3] For $\mathcal{A}$'s ODH_u query, we let $\mathcal{B}_2$ respond with $y^{\star\prime}$ if $\mathcal{A}$ queries (g^v, i) and have $\mathcal{B}_2$ relay the query to its own ODH_u oracle otherwise.
2. $\mathcal{A}$ asks a challenge $x^\star \neq 0$. We let $\mathcal{B}_2$ abort if $x^\star = i$ (which happens with probability at most $\frac{1}{2}$). Otherwise we know $x^\star = j$ and let $\mathcal{B}_2$ query $(g^v, x^\star)$ to its own ODH_u oracle, obtaining y. It sets $y_0^\star = y$ and $y_1^\star \xleftarrow{\$} \{0,1\}^\lambda$ and returns $y_{b^*}^\star$ to $\mathcal{A}$.[4] For $\mathcal{A}$'s ODH_u query, $\mathcal{B}_2$ simply relays queries (S, x) with $S \neq g^v$ or $x \notin \{0, i\}$. In case $\mathcal{A}$ queries $(g^v, 0)$, $\mathcal{B}_2$ asks its own ODH_u query on (g^v, j), obtaining y', and returns $y^{\star\prime} \oplus y'$ to $\mathcal{A}$. In case $\mathcal{A}$ queries (g^v, i), $\mathcal{B}_2$ responds with $y^{\star\prime}$.

When $\mathcal{A}$ stops and outputs its guess b', $\mathcal{B}_2$ stops as well and outputs 0 if $b^* = b'$ and 1 otherwise.

Let us see why $\mathcal{B}_2$ correctly simulates the $\mathsf{snPRF\text{-}ODH}$ game either for function F_2 (in case $b = 0$ in $\mathcal{B}_2$'s $\mathsf{mnPRF\text{-}ODH}$ game) or for function F_2' (in case $b = 1$). Through using $y^{\star\prime}$, if $b = 0$ (and hence $y^{\star\prime} = {y^{\star}}'_0 = \mathsf{G}(g^{uv}, i)$) $\mathcal{B}_2$ computes $\mathsf{F}_2(g^{uv}, 0) = \mathsf{G}(g^{uv}, i) \oplus \mathsf{G}(g^{uv}, j)$ in the challenge and ODH_u responses for $x^\star = 0$ resp. $x = 0$, otherwise it computes $\mathsf{F}_2'(g^{uv}, 0) = R(i) \oplus \mathsf{G}(g^{uv}, j)$. Similarly, $\mathcal{B}_2$ responds to ODH_u queries on (g^v, i) with either $\mathsf{G}(g^{uv}, i)$ or $R(i)$, depending on $y^{\star\prime}$. Furthermore, by picking $b^* \xleftarrow{\$} \{0,1\}$ on its own and aborting on $x^\star = i$, algorithm $\mathcal{B}_2$ also correctly provides $\mathcal{A}$ with a real-or-random challenge response in both cases.

We can determine the advantage of $\mathcal{B}_2$ against the $\mathsf{mnPRF\text{-}ODH}$ security of G as follows (informally denoting by, e.g., "$\mathcal{B}_2 \to 0$" the event that $\mathcal{B}_2$ outputs 0).

$$\begin{aligned} \mathsf{Adv}^{\mathsf{mnPRF\text{-}ODH}}_{\mathsf{G},\mathcal{B}_2} &= \left|\Pr[\mathcal{B}_2 \to b] - \tfrac{1}{2}\right| = \left|\tfrac{1}{2} \cdot (\Pr[\mathcal{B}_2 \to 0 \mid b = 0] + \Pr[\mathcal{B}_2 \to 1 \mid b = 1]) - \tfrac{1}{2}\right| \\ &= \tfrac{1}{2} \cdot |\Pr[\mathcal{B}_2 \to 0 \mid b = 0] - \Pr[\mathcal{B}_2 \to 0 \mid b = 1]| \\ &= \tfrac{1}{4} \cdot |\Pr[\mathcal{B}_2 \to 0 \mid b = 0 \wedge \neg abort] - \Pr[\mathcal{B}_2 \to 0 \mid b = 1 \wedge \neg abort]| \\ &\geq \tfrac{1}{4} \cdot \left|\mathsf{Adv}^{\mathsf{snPRF\text{-}ODH},x^\star \leq n}_{\mathsf{F}_2,\mathcal{A}} - \tfrac{1}{2} - \mathsf{Adv}^{\mathsf{snPRF\text{-}ODH},x^\star \leq n}_{\mathsf{F}_2',\mathcal{A}} + \tfrac{1}{2}\right| \\ &= \tfrac{1}{4} \cdot \left|\mathsf{Adv}^{\mathsf{snPRF\text{-}ODH},x^\star \leq n}_{\mathsf{F}_2,\mathcal{A}} - \mathsf{Adv}^{\mathsf{snPRF\text{-}ODH},x^\star \leq n}_{\mathsf{F}_2',\mathcal{A}}\right| \end{aligned}$$

Hence, the advantage difference for $\mathcal{A}$ introduced by switching from F_2 to F_2' is bounded as

$$\left|\mathsf{Adv}^{\mathsf{snPRF\text{-}ODH},x^\star \leq n}_{\mathsf{F}_2,\mathcal{A}} - \mathsf{Adv}^{\mathsf{snPRF\text{-}ODH},x^\star \leq n}_{\mathsf{F}_2',\mathcal{A}}\right| \leq 4 \cdot \mathsf{Adv}^{\mathsf{mnPRF\text{-}ODH}}_{\mathsf{G},\mathcal{B}_2}.$$

[3] Here, including $y^{\star\prime}$ (which is either ${y^{\star}}'_0$ or ${y^{\star}}'_1$) makes the response represent either F_2 or F_2'. Further, the bit b^* chosen by $\mathcal{B}_2$ renders the response real-or-random in either case.

[4] This makes the response to $\mathcal{A}$ real-or-random.

In a second step, we similarly modify the $\mathsf{snPRF\text{-}ODH}$ game for $\mathcal{A}$ by switching from F_2' to a function F_2'' which now also replaces values $\mathsf{G}(K, j)$ by the output of a random function:

$$\mathsf{F}_2''(K, x) := \begin{cases} \mathsf{F}_2(K, x) & \text{if } K \neq g^{uv} \\ R(1) \oplus R(2) & \text{if } K = g^{uv} \text{ and } x = 0 \\ R(x) & \text{if } K = g^{uv} \text{ and } x \in \{1, 2\} \\ \mathsf{G}(K, x) & \text{if } K = g^{uv} \text{ and } x > 2 \end{cases}$$

As before, this change can be bounded by the advantage of an algorithm $\mathcal{B}_3$ against the $\mathsf{mnPRF\text{-}ODH}$ security of g. Much like the reduction $\mathcal{B}_2$, $\mathcal{B}_3$ encodes its challenge on $x^{\star\prime} = i \xleftarrow{\$} \{1, 2\}$ as the value representing $\mathsf{G}(g^{uv}, i)$ in F_2' resp. $R(i)$ in F_2''. For inputs $x = j$, $i \neq j \in \{1, 2\}$, $\mathcal{B}_3$ samples a random value $R(j) \xleftarrow{\$} \{0, 1\}^\lambda$ on its own. Following the same analysis as for the switch from F_2 to F_2', we can bound the advantage difference introduced by this change as

$$\left| \mathsf{Adv}^{\mathsf{snPRF\text{-}ODH}, x^\star \leq n}_{\mathsf{F}_2', \mathcal{A}} - \mathsf{Adv}^{\mathsf{snPRF\text{-}ODH}, x^\star \leq n}_{\mathsf{F}_2'', \mathcal{A}} \right| \leq 4 \cdot \mathsf{Adv}^{\mathsf{mnPRF\text{-}ODH}}_{\mathsf{G}, \mathcal{B}_3}.$$

Finally, in the $\mathsf{snPRF\text{-}ODH}$ game for F_2'', the values $\mathsf{F}_2''(g^{uv}, 1)$ and $\mathsf{F}_2''(g^{uv}, 2)$ are independent, uniformly random values and $\mathsf{F}_2''(g^{uv}, 0)$ their XOR combination. The advantage of $\mathcal{A}$ in this game can hence immediately be reduced to the advantage $\mathcal{B}_4$ in a $\mathsf{mnPRF\text{-}ODH}$ game against G.[5] For this, $\mathcal{B}_4$ replies to all (challenge and ODH_u) queries on (g^{uv}, x) for $x \in \{0, 1, 2\}$ on its own with $R(x)$ (for $x \in \{1, 2\}$) resp. $R(1) \oplus R(2)$ (for $x = 0$), picking random values $R(1), R(2) \xleftarrow{\$} \{0, 1\}^\lambda$ itself. Note that, as $\mathcal{A}$ can pose at most two queries (one challenge and one ODH_u query), these responses are consistent even though $\mathcal{B}_4$ does respond with distinct real ($R(x)$) or random values. Beyond that, $\mathcal{B}_4$ can simply relay all other queries to its own (challenge and ODH_u) oracles.

This finally determines the advantage of $\mathcal{A}$ against F_2'' as

$$\mathsf{Adv}^{\mathsf{snPRF\text{-}ODH}, x^\star \leq n}_{\mathsf{F}_2'', \mathcal{A}} = \mathsf{Adv}^{\mathsf{mnPRF\text{-}ODH}}_{\mathsf{G}, \mathcal{B}_4}.$$

Combining the above advantage bounds yields the overall bound. □

In a similar way as for Proposition 4.4 we now also separate $\mathsf{nsPRF\text{-}ODH}$ security from $\mathsf{mnPRF\text{-}ODH}$, $\mathsf{ssPRF\text{-}ODH}$, and $\mathsf{nmPRF\text{-}ODH}$ security.

Proposition 4.5 ($\mathsf{nsPRF\text{-}ODH} \not\Longrightarrow \mathsf{mnPRF\text{-}ODH}, \mathsf{ssPRF\text{-}ODH}, \mathsf{nmPRF\text{-}ODH}$)**.** *If* G *from Definition 4.1 is* $\mathsf{nmPRF\text{-}ODH}$*-secure, then* $\mathsf{F}_2 \in \mathcal{F}$ *is* $\mathsf{nsPRF\text{-}ODH}$*-secure, but neither* $\mathsf{mnPRF\text{-}ODH}$*-, nor* $\mathsf{ssPRF\text{-}ODH}$*-, nor* $\mathsf{nmPRF\text{-}ODH}$*-secure. More precisely, for any efficient adversary* $\mathcal{A}$ *against the* $\mathsf{nsPRF\text{-}ODH}$ *security of* F_2*, there exist efficient algorithms* $\mathcal{B}_1$*, …,* $\mathcal{B}_4$ *such that*

$$\mathsf{Adv}^{\mathsf{nsPRF\text{-}ODH}}_{\mathsf{F}_2, \mathcal{A}} \leq \mathsf{Adv}^{\mathsf{nmPRF\text{-}ODH}}_{\mathsf{G}, \mathcal{B}_1} + 4 \cdot \mathsf{Adv}^{\mathsf{nmPRF\text{-}ODH}}_{\mathsf{G}, \mathcal{B}_2} + 4 \cdot \mathsf{Adv}^{\mathsf{nmPRF\text{-}ODH}}_{\mathsf{G}, \mathcal{B}_3} + \mathsf{Adv}^{\mathsf{nmPRF\text{-}ODH}}_{\mathsf{G}, \mathcal{B}_4},$$

but there exist algorithms $\mathcal{A}_1$*, …,* $\mathcal{A}_3$ *with non-negligible advantage* $\mathsf{Adv}^{\mathsf{mnPRF\text{-}ODH}}_{\mathsf{F}_2, \mathcal{A}_1} = \mathsf{Adv}^{\mathsf{ssPRF\text{-}ODH}}_{\mathsf{F}_2, \mathcal{A}_2} = \mathsf{Adv}^{\mathsf{nmPRF\text{-}ODH}}_{\mathsf{F}_2, \mathcal{A}_3} = 1 - 2^{-\lambda}$*.*

Proof. In Proposition 4.4 we have already established that F_2 provides no $\mathsf{nmPRF\text{-}ODH}$, $\mathsf{ssPRF\text{-}ODH}$ and $\mathsf{mnPRF\text{-}ODH}$ security given the respective algorithms $\mathcal{A}_1$, $\mathcal{A}_2$, and $\mathcal{A}_3$. For the proof that F_2 is $\mathsf{nsPRF\text{-}ODH}$-secure, observe that applying the mirrored proof that F_2 is $\mathsf{snPRF\text{-}ODH}$-secure from Proposition 4.4 establishes the desired bound. I.e., via algorithms $\mathcal{B}_1$, …, $\mathcal{B}_4$ with multiple *right-side* (ODH_v) oracle queries we can reduce the security to the $\mathsf{nmPRF\text{-}ODH}$ security of G. □

[5] We remark that in this reduction, $\mathcal{B}_4$ actually only needs to issue a single ODH_v query to its oracle for the simulation, hence the reduction would also apply to the $\mathsf{snPRF\text{-}ODH}$ security of G.

We now consider the separation of $\mathsf{ssPRF\text{-}ODH}$ from $\mathsf{mnPRF\text{-}ODH}$ and $\mathsf{nmPRF\text{-}ODH}$.

Proposition 4.6 ($\mathsf{ssPRF\text{-}ODH} \not\Rightarrow \mathsf{mnPRF\text{-}ODH}, \mathsf{nmPRF\text{-}ODH}$)**.** *If* G *from Definition 4.1 is* $\mathsf{msPRF\text{-}ODH}$*-secure, then* $\mathsf{F}_3 \in \mathcal{F}$ *is* $\mathsf{ssPRF\text{-}ODH}$*-secure, but neither* $\mathsf{mnPRF\text{-}ODH}$*- nor* $\mathsf{nmPRF\text{-}ODH}$*-secure. More precisely, for any efficient adversary* $\mathcal{A}$ *against the* $\mathsf{ssPRF\text{-}ODH}$ *security of* F_3*, there exist efficient algorithms* $\mathcal{B}_1, \ldots, \mathcal{B}_5$ *such that*

$$\mathsf{Adv}^{\mathsf{ssPRF\text{-}ODH}}_{\mathsf{F}_3,\mathcal{A}} \leq \mathsf{Adv}^{\mathsf{msPRF\text{-}ODH}}_{\mathsf{G},\mathcal{B}_1} + 3 \cdot \mathsf{Adv}^{\mathsf{msPRF\text{-}ODH}}_{\mathsf{G},\mathcal{B}_2} + 3 \cdot \mathsf{Adv}^{\mathsf{msPRF\text{-}ODH}}_{\mathsf{G},\mathcal{B}_3} + 3 \cdot \mathsf{Adv}^{\mathsf{msPRF\text{-}ODH}}_{\mathsf{G},\mathcal{B}_4} + \mathsf{Adv}^{\mathsf{msPRF\text{-}ODH}}_{\mathsf{G},\mathcal{B}_5},$$

but there exist algorithms $\mathcal{A}_1$*,* $\mathcal{A}_2$ *with non-negligible advantage* $\mathsf{Adv}^{\mathsf{mnPRF\text{-}ODH}}_{\mathsf{F}_3,\mathcal{A}_1} = \mathsf{Adv}^{\mathsf{nmPRF\text{-}ODH}}_{\mathsf{F}_3,\mathcal{A}_2} = 1 - 2^{-\lambda}$.

Proof. As for the previous separations, it is apparent that F_3 cannot provide security in the sense of $\mathsf{mnPRF\text{-}ODH}$ or $\mathsf{nmPRF\text{-}ODH}$. An adversary querying the three values $\mathsf{F}_3(g^{uv}, i)$, $i \in \{1, 2, 3\}$, through its ODH_u resp. ODH_v oracle can distinguish the real value $\mathsf{F}_3(g^{uv}, 0)$ from a random value, except with negligible probability $2^{-\lambda}$.

For proving that F_3 is $\mathsf{ssPRF\text{-}ODH}$-secure if G is $\mathsf{msPRF\text{-}ODH}$-secure, we apply the same proof strategy applied in the proof of Proposition 4.4 for showing $\mathsf{snPRF\text{-}ODH}$ security of F_2 based on the $\mathsf{mnPRF\text{-}ODH}$ security of G, but with two modifications. The first difference we have to take into account is that an $\mathsf{ssPRF\text{-}ODH}$ adversary $\mathcal{A}$ may (beyond its challenge and one ODH_u query) also issue a (single) ODH_v query for a value $T \neq g^u$. While such a query does not interfere with our proof steps, we cannot relay such a query to an ODH_u oracle and hence need that G is $\mathsf{msPRF\text{-}ODH}$-secure, providing the reductions with a (single) ODH_v oracle call. The other change is that we replace values $\mathsf{G}(g^{uv}, i)$ by the outputs $R(i)$ of a random function R for three ($i \in \{1, 2, 3\}$) instead of two labels and in each step abort if $\mathcal{A}$ asks a challenge $x^\star = i$ for our guessed i, which now happens with probability at most $\frac{1}{3}$. Accounting for the modified factor, this yields the three intermediary advantage bounds of $3 \cdot \mathsf{Adv}^{\mathsf{msPRF\text{-}ODH}}_{\mathsf{G},\mathcal{B}_j}$ for $j \in \{1, 2, 3\}$. □

4.2 Separations in the Random Oracle Model

In the following we use the following problem of computing non-trivial v-th roots in $\mathbb{G}$ for implicitly given v. That is, consider an algorithm $\mathcal{A}$ which outputs some group element $x \in \mathbb{G}$ with $x \neq 1$ (and some state information), then receives g^v for random $v \xleftarrow{\$} \mathbb{Z}_q$, and finally outputs y given g^v and the state information, such that $y^v = x$. Denote by $\mathsf{Adv}^{\mathsf{iiDH}}_{\mathbb{G},\mathcal{A}}$ the probability that $\mathcal{A}$ succeeds in this *interactive inversion* DH problem.

Note that the problem would be trivial if $x = 1$ was allowed (in which case $y = 1$ would provide a solution), or if x can be chosen after having seen g^v (in which case $x = g^v$ and $y = g$ would trivially work). Excluding these trivial cases, in terms of generic or algebraic hardness the problem is equivalent to the CDH problem. Namely, assume $\mathcal{A}$ "knows" $\alpha \in \mathbb{Z}_q$ such that $x = g^\alpha$. Since x is chosen before seeing g^v the adversary can only compute it as a power of g and, in addition, $x \neq 1$ implies $\alpha \neq 0$. Therefore, for any valid solution y the value $y^{1/\alpha}$ would be a v-th root of g, because $(y^{1/\alpha})^v = x^{1/\alpha} = g$. This problem of computing $g^{1/v}$ from g, g^v, however, is known as the inversion-DH (iDH) problem; it is equivalent to the CDH problem with a loose reduction [BDZ03].

For our separation result we still need a slightly stronger version here where, in the second phase, the adversary also gets access to a decision oracle which, on input two group elements $A, B \in \mathbb{G}$ outputs 1 if and only if $A^v = B$. We call this the strong interactive-inversion DH problem and denote it by siiDH. Note that for example for a pairing-based group such an oracle is given for free, while computing a v-th root of g (or, equivalently, solving the DH problem may still be hard).

Proposition 4.7 ($\mathsf{nmPRF\text{-}ODH} \not\Rightarrow \mathsf{snPRF\text{-}ODH}$)**.** *In the random oracle model, and assuming* StDH *and* siiDH*, there exists a function* F^{RO} *which is* $\mathsf{nmPRF\text{-}ODH}$*-secure but not* $\mathsf{snPRF\text{-}ODH}$*-secure. More* *precisely, for any efficient adversary $\mathcal{A}^{\mathsf{RO}}$ against the* nmPRF-ODH *security of* F^{RO}*, there exist efficient algorithms* $\mathcal{B}_1, \mathcal{B}_2$*, such that*

$$\mathsf{Adv}^{\mathsf{nmPRF\text{-}ODH}}_{\mathsf{F}^{\mathsf{RO}},\mathcal{A}^{\mathsf{RO}}} \leq \mathsf{Adv}^{\mathsf{StDH}}_{\mathbb{G},\mathcal{B}_1} + h \cdot \mathsf{Adv}^{\mathsf{siiDH}}_{\mathbb{G},\mathcal{B}_2} + \frac{h}{q} + 2^{-\lambda},$$

for the at most h *queries to the random oracle, but there exists an algorithm* $\mathcal{A}^{\mathsf{RO}}$ *with non-negligible advantage* $\mathsf{Adv}^{\mathsf{snPRF\text{-}ODH}}_{\mathsf{F}^{\mathsf{RO}},\mathcal{A}} \geq 1 - 2^{-\lambda+1}$.

Proof. Consider the function $\mathsf{F}^{\mathsf{RO}} : \mathbb{G} \times (\mathbb{G} \times \{0,1\}^{\lambda}) \to \{0,1\}^{\lambda}$ defined as:

$$\mathsf{F}^{\mathsf{RO}}(K,(x,y)) = \begin{cases} y & \text{if } \mathsf{RO}(Kx^{-1},(x,0^{\lambda})) = y \text{ and } x \neq 1 \\ \mathsf{RO}(K,(x,y)) & \text{else} \end{cases}$$

We first argue that this function is easy to attack in the snPRF-ODH sense, when given an ODH_u oracle which can be queried once after having learned the challenge values. To this end let $\mathcal{A}^{\mathsf{RO}}$, on input g, g^u, use $x^\star = (g^u, 0^\lambda)$ in the challenge query. This yields $(g^v, y^\star)$. If $g^u = 1$ then our adversary can compute the key $g^{uv} = g^v$ and verify $y^\star$ directly via RO, outputting 0 if the challenge value matches the computed value. Otherwise, in the single query to the ODH_u oracle let the adversary forward $(g^{v+1}, (g^u, y^\star))$ to get the value y'. Output 0 if and only if this answer y' matches $y^\star$.

Note that if $y^\star$ has been generated as the output $\mathsf{PRF}(g^{uv}, (g^u, 0^\lambda))$ then our adversary always returns 0. For $g^u = 1$ this is straightforward, and in the other case the key in the query to ODH_u equals g^{uv+v} such that the exception is satisfied. If $y^\star$ is random, then we have a match with probability at most $2^{-\lambda+1}$: either the random $y^\star$ accidentally matches the actual pseudorandom value, or the random oracle output $\mathsf{RO}(g^{uv+u}, (g^u, y^\star))$ coincides with $y^\star$. Both events happen with probability at most $2^{-\lambda}$.

For the security in the nmPRF-ODH sense recall that now the adversary has "only" access to an ODH_v oracle for multiple queries, after having received the challenge value $y^\star$ for challenge query $x^\star$. We first argue that the adversary can never ask the random oracle about the DH key g^{uv} used during the attack. This step will be discussed more thoroughly for the more general case in the instantiation proof of nmPRF-ODH under the StDH assumption, so we skip it here. We emphasize that this also provides a sound simulation of the random oracle answers without explicit knowledge of u, v in the attack.

Next, note that we can assume that the adversary with its challenge query $x^\star = (x', y')$ does not trigger the exceptional event of having the function output $y' = \mathsf{RO}(g^{uv}(x')^{-1}, x', 0^\lambda)$, since v is only picked after the adversary has submitted x'. Hence, the probability that any previous random oracle query was about this key, is at most $\frac{h}{q}$. Else, y' matches this previously not queried random oracle value only with probability $2^{-\lambda}$. Hence, the adversary receives the value $y^\star$ as either $\mathsf{RO}(g^{uv}, (x', y'))$ or as a random value.

Since queries $(g^u, x^\star)$ to the ODH_v oracle are prohibited and the adversary never asks the random oracle about the key g^{uv}, the only possibility to distinguish a random challenge $y^\star$ from an actual random oracle answer, is to trigger the exceptional case in the function evaluation for the challenge data. So assume that the adversary queries its ODH_v oracle about $(T, (x, y))$ such that $T^v x^{-1} = g^{uv}$, $x^\star = (x, 0^\lambda)$ for $x \neq 1$, and y matches $\mathsf{RO}(T^v x^{-1}, x, 0^\lambda)$. In this case we have $x = (Tg^{-u})^v$ and therefore Tg^{-u} is a v-th root of $x \neq 1$. This can now be straightforwardly turned into an attack against the (strong) interactive inversion assumption.

The reduction to the siiDH problem is as follows. Simulate the attack in the random oracle model by picking u and by using the decisional oracle to answer consistently (as in the reduction to the StDH problem).[6] Use x from the adversary's challenge as the element in the interactive game, receiving then

[6] This, however, requires some more careful argument here in the simulation of the random oracle without knowledge

g^v and guessing the right ODH_v query $(T,(x,y))$ of the adversary and outputting Tg^{-u} as the v-th root. This breaks the siiDH assumption, with a loss q_v in the number of queries to the ODH_v oracle.

Given that the adversary never makes a "bad" query to the ODH_v oracle it cannot distinguish the two cases at all. □

The idea can now be transferred to the case that we still allow one oracle query to ODH_u, basically by "secret sharing" the reply in the exceptional case among two queries:

Proposition 4.8 ($\mathsf{smPRF\text{-}ODH} \not\Longrightarrow \mathsf{mnPRF\text{-}ODH}$)**.** *In the random oracle model, and assuming* StDH *and* siiDH*, there exists a function* F^{RO} *which is* $\mathsf{smPRF\text{-}ODH}$*-secure but not* $\mathsf{mnPRF\text{-}ODH}$*-secure. More precisely, for any efficient adversary* $\mathcal{A}^{\mathsf{RO}}$ *against the* $\mathsf{smPRF\text{-}ODH}$ *security of* F^{RO}*, there exist efficient algorithms* $\mathcal{B}_1, \mathcal{B}_2$*, such that*

$$\mathsf{Adv}^{\mathsf{smPRF\text{-}ODH}}_{\mathsf{F}^{\mathsf{RO}},\mathcal{A}^{\mathsf{RO}}} \leq \sqrt{\mathsf{Adv}^{\mathsf{StDH}}_{\mathbb{G},\mathcal{B}_1}} + h \cdot \mathsf{Adv}^{\mathsf{siiDH}}_{\mathbb{G},\mathcal{B}_2} + \frac{h}{q} + 2^{-\lambda},$$

for the at most h *queries to the random oracle, but there exists an algorithm* $\mathcal{A}^{\mathsf{RO}}$ *with non-negligible advantage* $\mathsf{Adv}^{\mathsf{mnPRF\text{-}ODH}}_{\mathsf{F}^{\mathsf{RO}},\mathcal{A}^{\mathsf{RO}}} \geq 1 - 2^{-\lambda+1}$*.*

In fact, in the negative result for $\mathsf{mnPRF\text{-}ODH}$ the adversary only needs to ask two queries to the ODH_u oracle *after* receiving the challenge query. Since the function is still secure for a single ODH_u query, this is optimal in this regard.

Proof. Consider the function $\mathsf{F}^{\mathsf{RO}} : \mathbb{G} \times (\mathbb{G} \times \{0,1\}^\lambda) \to \{0,1\}^\lambda$ defined as:

$$\mathsf{F}^{\mathsf{RO}}(K,(x,y)) = \begin{cases} y \oplus \mathsf{RO}(K,K,0^\lambda) & \text{if } \mathsf{RO}(Kx^{-1},(x,0^\lambda)) = y \text{ and } x \neq 1 \\ \mathsf{RO}(K,K,0^\lambda) & \text{if } \mathsf{RO}(Kx^{-1},(x,0^\lambda)) = 1^\lambda \oplus y \text{ and } x \neq 1 \\ \mathsf{RO}(K,(x,y)) & \text{else} \end{cases}$$

We first show again that this function is easy to attack in the $\mathsf{mnPRF\text{-}ODH}$ case. The adversary is very similar to the one in Proposition 4.7. Namely, our adversary asks the challenge query $x^\star = (g^u, 0^\lambda)$ to receive $y^\star$, then asks its ODH_u oracle about $(g^{v+1}, g^u, y^\star)$ and $(g^{v+1}, g^u, y^\star \oplus 1^\lambda)$, and adds the answers and compares them to $y^\star$. If and only if the values match it outputs 0. In case $g^u = 1$ it can again compute the reply directly. Overall, it always outputs 0 if $y^\star$ is the function value, and only with probability at most $2^{-\lambda+1}$ if it is random.

Arguing again that an algorithm with multi-query access to ODH_v and single-query access to ODH_u cannot succeed with non-negligible probability is also very close to the previous case in Proposition 4.7. For this we first note that we can substitute the answer for the exceptional cases of the single ODH_u query by simply returning an independent random answer. This is a valid simulation strategy as the adversary cannot make a further ODH_u query, nor a direct random oracle query about $\mathsf{RO}(g^{uv}, g^{uv}, 0^\lambda)$ because this would violate the StDH assumption (as before). Also, the adversary cannot trigger the exceptional evaluation(s) in the ODH_v oracle, since this would violate the siiDH assumption (as before). Furthermore, the adversary cannot make the "regular" function evaluation in an ODH_v call evaluate on $\mathsf{RO}(g^{uv}, g^{uv}, 0^\lambda)$ because for this it would need to call about $x = (g^{uv}, 0^\lambda)$, embedding the DH key already. (Formally we can assume that the adversary always calls the random oracle about $(x', x', 0^\lambda)$ before making a query $(x', 0^\lambda)$

of g^v. Since g^v is not given from the beginning we cannot simulate in a sound way the adversary's random oracle queries before making the challenge query (in which case we output x and only then receive g^v in the interactive-inversion problem). However, the probability that the adversary has made such a query earlier is at most $\frac{h}{q}$, since v is chosen at random in $\mathbb{Z}_q$, and we have already accounted for such cases for the event that the adversary has not triggered the exceptional event in the challenge query.

to the ODH_v oracle, such that this would again violate the StDH assumption with a loose reduction.) Hence, providing a random answer yields a consistent simulation from the adversary's point of view.

Moreover, the adversary cannot make a call $(g^v, x^\star)$ to the ODH_u oracle such that any call either returns a random oracle value at a different point, or a now independently generated answer. This means that the adversary does not have the possibility to distinguish random challenge values from genuine ones anymore. □

4.3 Discussion

Let us close this section with some remarks about the separations.

Remark. Our separating function family (cf. Definition 4.1) establishes quite a number of separations, but cannot be used in order to separate the remaining variants. This is due to the fact that our function family cannot separate between notions that both allow polynomial many queries as for example nmPRF-ODH and smPRF-ODH. Thus, we have turned to the random oracle model to establish further separations. Using this model is alleviated by the result about the implausibility of instantiating the PRF-ODH assumption in the standard model.

In the random oracle model we have shown that it is crucial if the adversary has access to the ODH_u oracle or not (or how many times). This uses some asymmetry in the two oracles, namely, that g^u is given before the challenge query, and g^v only after. Our separations take advantage of this difference, visualized via the interactive-inversion DH problem which is only hard if $x^\star$ is chosen before receiving g^v.

It is currently open if the other notions are separable. Beyond the asymmetry that g^u is already available before the challenge, it is unclear how to "encode" other distinctive information into the input to the "memoryless" PRF which one oracle can exploit but the other one cannot.

Remark. In case our generic PRF-ODH assumption (cf. Definition 2.1) would provide the adversary additionally with the share g^v in the initialization phase (cf. step 1) then Figure 2 would symmetrically "collapse" along the vertical axis in the middle. In other words, this would result in equivalences of the notions snPRF-ODH and nsPRF-ODH, mnPRF-ODH and nmPRF-ODH, as well as msPRF-ODH and smPRF-ODH. Note that this is not a contradiction to our separation results among those notions, as they only work if (and exploit that) g^v is not given in advance.

5 On the Impossibility of Instantiating PRF-ODH in the Standard Model

In this section we show that there is no *algebraic* black-box reduction $\mathcal{R}$ which reduces the snPRF-ODH assumption (and analogously the nsPRF-ODH assumption) to a class of hard cryptographic problems, called DDH-augmented abstract problems. With these problems one captures reductions to the DDH problem or to some general, abstract problem like collision resistance of hash functions.

5.1 Overview

The idea is to use the meta-reduction technique. Assume that we have an algebraic reduction $\mathcal{R}$ from the snPRF-ODH assumption which turns any black-box adversary into a solver for a DDH-augmented problem. Then we in particular consider an inefficient adversary $\mathcal{A}_\infty$ which successfully breaks the snPRF-ODH assumption with constant probability. The reduction, with black-box access to $\mathcal{A}_\infty$, must then solve the DDH-augmented problem. For this it can then either not take any advantage of the infinite power of $\mathcal{A}_\infty$—in which case we can already break the DDH-augmented problem—or it tries to elicit some useful information from $\mathcal{A}_\infty$. In the latter case we build our meta-reduction by simulating $\mathcal{A}_\infty$ *efficiently.* This is accomplished by exploiting the algebraic property of the reduction and "peeking" at the internals of the reduction's group element choices. Our meta-reduction will then solve the decisional square-DH problem, saying that (g, g^a, g^{a^2}) is indistinguishable from (g, g^a, g^b) from random a, b.

Our impossibility result works for pseudorandom functions PRF, which take as input arbitrary bit strings and maps them to λ bits. We stick with this convention here, but remark that our negative result also holds if the input length is 1 only, and the output length is super-logarithmic in λ. Similarly, we assume that PRF is a $\mathsf{nnPRF\text{-}ODH}$, although it suffices for our negative result that the function PRF for a random group element (and some fixed input, say 1) is pseudorandom, i.e., that $\mathsf{PRF}(X, 1)$ is indistinguishable from random for a uniformly chosen group element $X \xleftarrow{\$} \mathbb{G}$ (without giving any "Diffie–Hellman decomposition" of X).

Theorem 5.1. *Assume that there is an efficient algebraic black-box reduction $\mathcal{R}$ from the* $\mathsf{snPRF\text{-}ODH}$ *(or* $\mathsf{nsPRF\text{-}ODH}$*) assumption to a DDH-augmented problem. Then either the DDH-augmented problem is not hard, or the decisional square-DH problem is not hard.*

If one assumes vice versa that both the underlying augmented-DDH problem and decisional square-DH problem are hard, then this means that there cannot be a reduction as in the theorem to show security of the $\mathsf{nsPRF\text{-}ODH}$ or $\mathsf{snPRF\text{-}ODH}$ assumption.

5.2 DDH-augmented Cryptographic Problems

DDH-augmented problems are cryptographic problems in which the adversary either solves a DDH problem or some abstract (and independent) problem in which it receives some instance inst, can make oracle queries about this instance, and then generates a potential solution sol. The adversary can decide on the fly which of the two problems to solve. In terms of our setting here we build a reduction against such DDH-augmented problems, capturing for example the case that one aims to show security of the $\mathsf{PRF\text{-}ODH}$ assumption by assembling a scheme out of several primitives, including the DDH assumption, and giving reductions to each of them.

We next define cryptographic problems in a general way, where it is convenient to use the threshold of $\frac{1}{2}$ for decisional games to measure the adversary's advantage. We note that we can "lift" common computational games where the threshold is the constant 0 by outputting 1 if the adversary succeeds or if an independent coin flip lands on 1. Formally, a cryptographic problem consists of three probabilistic polynomial-time algorithms $\mathsf{P} = (\mathsf{P.Gen}, \mathsf{P.Ch}, \mathsf{P.Vf})$ such that for any probabilistic polynomial-time algorithm $\mathcal{A}$ we have

$$\mathsf{Adv}^{\mathrm{Prob}}_{\mathsf{P},\mathcal{A}}(\lambda) := 2 \cdot \left(\mathrm{Prob}\left[\mathsf{P.Vf}(\mathsf{secret}, \mathsf{sol}) = 1 : \begin{array}{l} (\mathsf{inst}, \mathsf{secret}) \xleftarrow{\$} \mathsf{P.Gen}(1^\lambda), \\ \mathsf{sol} \xleftarrow{\$} \mathcal{A}^{\mathsf{P.Ch}(\mathsf{secret},\cdot)}(1^\lambda, \mathsf{inst}) \end{array} \right] - \frac{1}{2} \right)$$

is negligible.

A DDH-augmented cryptographic problem $\mathsf{P}^{\mathrm{DDH}}$ for some group $\mathbb{G}$ (or, more precisely, for some sequence of groups) based on a problem P, consists of the following algorithms:

- $\mathsf{P.Gen}^{\mathrm{DDH}}(1^\lambda)$ runs $(\mathsf{inst}, \mathsf{secret}) \xleftarrow{\$} \mathsf{P.Gen}(1^\lambda)$, picks $x, y, z \xleftarrow{\$} \mathbb{Z}_q$ and $b \xleftarrow{\$} \{0, 1\}$, and outputs $\mathsf{inst}^{\mathrm{DDH}} = (g^x, g^y, g^{xy+bz}, \mathsf{inst})$ and $\mathsf{secret}^{\mathrm{DDH}} = (b, \mathsf{secret})$.

- $\mathsf{P.Ch}^{\mathrm{DDH}}(\mathsf{secret}^{\mathrm{DDH}}, \cdot)$ runs $\mathsf{P.Ch}(\mathsf{secret}, \cdot)$.

- $\mathsf{P.Vf}^{\mathrm{DDH}}(\mathsf{secret}^{\mathrm{DDH}}, \mathsf{sol}^{\mathrm{DDH}})$ checks if $\mathsf{sol}^{\mathrm{DDH}} = (''\mathsf{DDH}'', b')$ and, if so, outputs 1 if and only if $b = b'$ for bit b in $\mathsf{secret}^{\mathrm{DDH}}$. If $\mathsf{sol}^{\mathrm{DDH}} = (''\mathsf{P}'', \mathsf{sol})$ then the algorithm here outputs $\mathsf{P.Vf}(\mathsf{secret}, \mathsf{sol})$. In any other case it returns 0.

5.3 Algebraic Reductions for the snPRF-ODH Assumption

Algebraic reductions have been considered in [BV98] and abstractly defined in [PV05]. The idea is that the reduction can only perform group operations in the pre-defined way, e.g., by multiplying given elements. As a consequence, whenever the reduction on input group elements $g_1, g_2, \dots$ generates a group element $A \in \mathbb{G}$ one can output a representation $(\alpha_1, \alpha_2, \dots)$ such that $A = \prod g_i^{\alpha_i}$. In [PV05] this is formalized by assuming the existence of an algorithm which, when receiving the reduction's input and random tape, can output the representation in addition to A.

In order to simplify the presentation here, we simply assume that the reduction, when forwarding some group element to the adversary, outputs the representation itself. The base elements $g_1, g_2, \dots$ for the representation are those which the reduction has received so far, as part of the DDH-part of the input or from the interaction with the adversary. The representation is hidden from the adversary in the simulation, of course, but our meta-reduction may exploit this information.

We consider (algebraic) reductions $\mathcal{R}$ which use the adversary $\mathcal{A}$ in a black-box way. The reduction may invoke multiple copies of the adversary, possibly rewinding copies. We use the common technique of derandomizing our (unbounded) adversary in question by assuming that it internally calls a truly random function on the communication so far, when it needs to generate some randomness. Note that the truly random function is an integral part of the adversary, and that we view the adversary being picked randomly from all adversaries with such an embedded function. Since the reduction is supposed to work for all successful adversaries, it must also work for such randomly chosen adversaries.

It is now convenient to enumerate the adversary's instances which the reduction invokes as $\mathcal{A}_i$ for $i = 1, 2, \dots$. Since our adversary in question is deterministic we can assume that the reduction "abandons" a copy $\mathcal{A}_i$ forever, if it starts the next copy $\mathcal{A}_{i+1}$. This is without loss of generality because the reduction can re-run a fresh copy to the state where it has left the previous instance. This also means that the reduction can effectively re-set executions with the adversary.

The reduction receives as input a triple (g^x, g^y, g^z) and some instance inst and should decide if $g^z = g^{xy}$ or g^z is random, or provide a solution sol to inst with the help of oracle $\mathsf{P.Ch}$. We stress that the reduction is algebraic with respect the DDH-part of the DDH-augmented problem. In particular, encasing a PRF-ODH-like assumption into the general P problem and providing a trivial reduction to the problem itself is not admissible. The group elements (and their representations) handed to the adversary in the reduction are determined by the DDH-part of the input. Finally, we note that we only need that, if $\mathcal{R}$ interacts with an adversary against snPRF-ODH with advantage $1 - 2^{-\lambda}$, then $\mathcal{R}$ solves the DDH-augmented problem with a non-negligible advantage.

5.4 Outline of Steps

Our negative result proceeds in three main steps:

1. We first define an all-powerful adversary $\mathcal{A}_\infty$ which breaks the snPRF-ODH assumption by using its infinite power. This adversary will, besides receiving the challenge at point $x^\star = 0$, ask the ODH_u oracle to get the value at $(S, 1)$ for random $S = g^s$, where the random value s is generated via the integral random function. It then uses its power to compute the Diffie–Hellman key g^{uv}, verifies the answer of oracle ODH_u with the help of s, and only if this one is valid, gives the correct answer concerning the challenge query. In any other case, the adversary aborts.

2. We then show that the algebraic reduction $\mathcal{R}$, potentially spawning many black-box copies of our adversary $\mathcal{A}_\infty$, must answer correctly to the ODH_u query in one copy and use the input values g^x, g^y, g^z non-trivially, or else we can already break the underlying DDH-augmented problem efficiently.

3. Next we show that, if the reduction answers correctly and non-trivially in one of the copies, then we can —using the algebraic nature of the reduction— replace the adversary $\mathcal{A}_\infty$ by an efficient algorithm, the meta-reduction $\mathcal{M}$, and either break the decisional square-DH assumption or refute the pseudorandomness of PRF for a fresh random group element.

The decisional square-DH assumption says that it is infeasible to distinguish (g, g^a, g^{a^2}) from (g, g^a, g^b) for random a, b. It implies the DDH assumption, but is only known to be equivalent to classical DH problem in the computational case [BDZ03]. More formally, we will use the following variation: $(g, g^a, g^{a^2}, g^{a^2b}, g^b, g^{ab})$ is indistinguishable from $(g, g^a, g^{a^2}, g^{a^2b}, g^b, g^c)$ for random a, b, c.

We briefly argue that the above decision problem follows from the decisional square-DH assumption. The latter assumption implies that we can replace g^{a^2} and g^{a^2b} in these tuples by group elements g^d, g^{db} for random d, using knowledge of b to compute the other elements. Then, by the DDH assumption, we can replace g^{ab} in such tuples by a random group element g^c, using knowledge of d to compute the other group elements g^d, g^{bd}. In the last step we can re-substitute g^d, g^{db} again by g^{a^2} and g^{a^2b}, using knowledge of b and c to create the other group elements.

5.5 Defining the All-powerful Adversary

Let us define our adversary $\mathcal{A}_\infty$ (with an internal random function $f : \{0,1\}^* \to \mathbb{Z}_q$) against $\mathsf{snPRF\text{-}ODH}$ formally:

1. Adversary $\mathcal{A}_\infty$ receives g, g^u as input.

2. It then asks the challenge oracle about $x^\star = 0$ to receive $y^\star$ and g^v. We call this the challenge step.

3. It computes $s = f(g^u, g^v, y^\star)$ and $S = g^s$ and asks the ODH_u oracle about $(S, 1)$ to get some y_t. We call this the test step.

4. It computes $S^u = (g^u)^s$ and, using its unbounded computational power, also g^{uv}.

5. If $y_t \neq \mathsf{PRF}(S^u, 1)$ then $\mathcal{A}_\infty$ aborts.

6. Else, $\mathcal{A}_\infty$ outputs 0 if and only if $y^\star = \mathsf{PRF}(g^{uv}, 0)$, and 1 otherwise.

Note that the probability that $\mathcal{A}_\infty$ outputs the correct answer in an actual attack is $1 - 2^{-\lambda}$ and thus optimal; the small error of $2^{-\lambda}$ is due to the case that the random $y^\star$ may accidentally hit the value of the PRF function.

5.6 Reductions Without Help

Ideally we would now first like to conclude that any reduction which does not provide a correct answer for the test step in any of the copies, never exploits the adversary's unlimited power and would thus essentially need to immediately succeed, without the help of $\mathcal{A}_\infty$. We can indeed make this argument formal, simulating $\mathcal{A}_\infty$ efficiently by using lazy sampling techniques for the generation of s and always aborting in Step 5 if reaching this point. However, we need something slightly stronger here.

Assume that the reduction provides some g^u in one of the copies for which it knows the discrete logarithm u, i.e., it is not a non-trivial combination of the input values g^x, g^y, g^z for unknown logarithms. Then the reduction can of course answer the adversary's test query $(S, 1)$ successfully by computing S^u and $\mathsf{PRF}(S^u, 1)$. Yet, in such executions it can also compute the reply to the challenge query itself, even if it does not know the discrete logarithm of g^v. In this sense the reduction cannot gain any knowledge about its DDH input, and we also dismiss such cases as useless.

Formally, we call the j-th run of one of the adversary's copies *useless* if either the instance aborts in (or before) Step 5, or if the representation of the adversary's input g^u in this copy in the given group elements g, g^x, g^y, g^z and the challenge query values $S_1, S_2, \ldots, S_{j-1}$ so far, i.e.,

$$g^u = (g^x)^\alpha (g^y)^\beta (g^z)^\gamma g^\delta \cdot \prod_{i<j} S_i^{\sigma_i},$$

satisfies $x\alpha + y\beta + z\gamma = 0 \bmod q$. Note that the reduction may form g^u with respect to all externally provided group elements, including the S_i's, such that we also need to account for those elements. We will, however, always set all of them to $S_i = g^{s_i}$ for some known s_i, such that we are only interested in the question if combination of the DDH input values g^x, g^y, g^z vanishes.

Let $\mathsf{useless}_j$ be the event that the j-th instance is useless in the above sense. For such a useless copy we can efficiently simulate adversary $\mathcal{A}_\infty$, because it either aborts early enough, or the algebraic reduction outputs some g^u with its representation from which we can compute the discrete logarithm $u = \delta + \sum_{i<j} s_i\sigma_i \bmod q$ and thus execute the decision and test steps of $\mathcal{A}_\infty$. Let $\mathsf{useless}$ be the event that all executions of $\mathcal{A}_\infty$ of the reduction are useless. We next argue that, if the event $\mathsf{useless}$ happens with overwhelming probability, then we can solve the DDH-augmented problem immediately.

The claim holds as we can emulate the all-powerful adversary $\mathcal{A}_\infty$ easily, if the reduction essentially always forgoes to run the adversary till the very end or uses only trivial values g^u. Let $(g^x, g^y, g^z, \mathsf{inst})$ be our input and pass this to the reduction. We simply emulate *all* the copies of the adversary efficiently by:

- using lazy sampling to emulate the random function f,
- for each invocation check at the beginning that $(g^x)^\alpha (g^y)^\beta (g^z)^\gamma = 1$ for the representation received with the input g^u for that instance, in which case we can use the discrete logarithm $u = \delta + \sum_{i<j} s_i\sigma_i \bmod q$ to run this copy of $\mathcal{A}_\infty$, and
- else always abort after having received y in Step 3.

Denote this way of simulating each copy by adversary $\mathcal{A}_{\text{ppt}}$ (even though, technically, the copies share state for the lazy sampling technique and should be thus considered as one big simulated adversary). Then

$$\begin{aligned}
&\text{Prob}\left[\mathsf{P.Vf}(\mathsf{secret}, \mathsf{sol}) = 1 : \begin{array}{l}(\mathsf{inst}, \mathsf{secret}) \xleftarrow{\$} \mathsf{P.Gen}(1^\lambda),\\ \mathsf{sol} \xleftarrow{\$} \mathcal{R}^{\mathsf{P.Ch}(\mathsf{secret},\cdot),\mathcal{A}_\infty}(1^\lambda, \mathsf{inst})\end{array}\right]\\
\leq &\text{Prob}\left[\mathsf{P.Vf}(\mathsf{secret}, \mathsf{sol}) = 1 : \begin{array}{l}(\mathsf{inst}, \mathsf{secret}) \xleftarrow{\$} \mathsf{P.Gen}(1^\lambda),\\ \mathsf{sol} \xleftarrow{\$} \mathcal{R}^{\mathsf{P.Ch}(\mathsf{secret},\cdot),\mathcal{A}_{\text{ppt}}}(1^\lambda, \mathsf{inst})\end{array}\right] + \text{Prob}\left[\overline{\mathsf{useless}}\right]
\end{aligned}$$

The latter probability for event $\overline{\mathsf{useless}}$ is negligible by assumption. We therefore get an efficient algorithm $\mathcal{R}^{\mathcal{A}_{\text{ppt}}}$ which breaks the DDH-augmented problem with non-negligible advantage.

5.7 Our Meta-reduction

We may from now on thus assume that $\text{Prob}\left[\overline{\mathsf{useless}}\right] \not\approx 0$ is non-negligible. This implies that the reduction answers at least in one copy of $\mathcal{A}_\infty$ of the at most polynomial number $q(\lambda)$ in the test queries in Step 3 with the correct value y_t for some non-trivial input g^u, with non-negligible probability. Our meta-reduction will try to guess the first execution k where this happens and to "inject" its input $g^a, g^{a^2}, g^{a^2b}, g^b, g^c$ into that execution in a useful way. More precisely, it will insert these values into g^x, g^y, g^z and S_k such that the expected key K for evaluating PRF for the test query equals a function of g^{ab} if $g^c = g^{ab}$, but is random if g^c is random. In the latter case predicting y is infeasible for the reduction, though, because the PRF is evaluated on a fresh and random key. This allows to distinguish the two cases.

The meta-reduction's injection strategy captures two possible choices of the reduction concerning the equation $x\alpha + y\beta + z\gamma \neq 0 \bmod q$ in the (hopefully correctly guessed) k-th execution. One is for the case that $x\alpha + y\beta \neq 0 \bmod q$, the other one is for the case that $x\alpha + y\beta = 0 \bmod q$ and thus $z\gamma \neq 0 \bmod q$ according to the assumption $x\alpha+y\beta+z\gamma \neq 0 \bmod q$. The meta-reduction will try to predict (via a random bit e) which case will happen and inject the values differently for the cases. This is necessary since the g^z-value, if it is not random, should contain the DH value of the other two elements.

Our meta-reduction $\mathcal{M}$ works as follows:

1. The meta-reduction receives $g^a, g^{a^2}, g^{a^2b}, g^b, g^c$ as input and should decide if $g^c = g^{ab}$. If $a = 0$ then we can decide easily, such that we assume that $a \neq 0$ from now on.

2. The meta-reduction picks an index $k \xleftarrow{\$} \{1, 2, \ldots, q(\lambda)\}$ for the polynomial bound $q(\lambda)$ of adversarial copies the reduction $\mathcal{R}$ runs with $\mathcal{A}_\infty$. It also picks $x', y', z' \xleftarrow{\$} \mathbb{Z}_q^*$, $s_1, \ldots, s_{k-1} \xleftarrow{\$} \mathbb{Z}_q$, $e, d \xleftarrow{\$} \{0, 1\}$, and samples $(\mathsf{inst}, \mathsf{secret}) \xleftarrow{\$} \mathsf{P.Gen}(1^\lambda)$.

3. For the first injection strategy, $e = 0$, it sets

$$g^x = (g^a)^{x'},\ g^y = (g^a)^{y'},\ g^z = (g^{a^2})^{x'y'} \text{ for } d = 0 \text{ resp. } g^z = (g^{a^2})^{z'} \text{ for } d = 1.$$

   For the other injection strategy, $e = 1$, it sets

$$g^x = (g^a)^{x'}, \quad g^y = g^{y'}, \quad g^z = (g^a)^{x'y'} \text{ for } d = 0 \text{ resp. } g^z = g^{az'} \text{ for } d = 1.$$

4. It invokes the reduction $\mathcal{R}$ on input g^x, g^y, g^z as well as inst.

5. The meta-reduction simulates the interactions of $\mathcal{R}$ with $\mathsf{P.Ch}$ and $\mathcal{A}_\infty$ as follows:

   - Each oracle query to $\mathsf{P.Ch}$ is answered by running the original algorithm $\mathsf{P.Ch}$ for secret.
   - Use lazy sampling to emulate the random function f.
   - For each of the first $j < k$ invocations of $\mathcal{A}_\infty$ check at the beginning that $(g^x)^{\alpha_j}(g^y)^{\beta_j}(g^z)^{\gamma_j} = 1$ for the representation received with the input g^{u_j} for that instance, in which case $\mathcal{M}$ can use the discrete logarithm $u_j = \delta_j + \sum_{i<j} s_i\sigma_i \bmod q$ to efficiently run this copy of $\mathcal{A}_\infty$, using $S_j = g^{s_j}$ for the test query.
   - Otherwise, if $(g^x)^{\alpha_j}(g^y)^{\beta_j}(g^z)^{\gamma_j} \neq 1$, for the j-th invocation of an adversarial copy of $\mathcal{A}_\infty$ for $j < k$, up to Step 3, efficiently simulate $\mathcal{A}_\infty$ using $S_j = g^{s_j}$ for the test query, and immediately abort after this step.

6. For the k-th invocation simulate $\mathcal{A}_\infty$ by using $S_k = g^b$. If $\mathcal{M}$ receives a reply y_t from $\mathcal{R}$, do the following. Let g^{u_k} be the input value of this adversary's copy. Since the reduction is algebraic it has also output values $\alpha_k, \beta_k, \gamma_k, \delta_k, \sigma_1, \ldots, \sigma_{k-1} \in \mathbb{Z}_q$ such that

$$g^{u_k} = (g^x)^{\alpha_k}(g^y)^{\beta_k}(g^z)^{\gamma_k} g^{\delta_k} \cdot \prod_{i<k} S_i^{\sigma_i}.$$

   Note that all the base elements, up to this point, only depend on g, g^a (and g^{a^2} in case of strategy $e = 0$) of $\mathcal{M}$'s inputs $g^a, g^{a^2}, g^{a^2b}, g^b, g^c$, because g^{u_k} is output before seeing $S_k = g^b$.[7]

[7] The same is true for g^{v_k} generated in the challenge query before, such that the result applies to the $\mathsf{nsPRF\text{-}ODH}$ assumption accordingly.

7. If strategy $e = 0$ is used and we have $a(x'\alpha_k + y'\beta_k) \neq 0 \bmod q$ (which can be checked for $a \neq 0$ by consulting the known values $x', \alpha_k, y', \beta_k$), then the meta-reduction decides as follows. From the value g^{a^2b} it can compute $g^{bz\gamma_k} = (g^{a^2b})^{x'y'\gamma_k}$ resp. $(g^{a^2b})^{z'\gamma}$ for both cases $d \in \{0,1\}$ and can then set
$$K = (g^c)^{x'\alpha_k + y'\beta_k} g^{bz\gamma_k} (g^b)^{\delta_k + \sum_{i<k} s_i\sigma_i}.$$
It immediately outputs 0 if $y_t = \mathsf{PRF}(K, 1)$, else it continues.

8. If strategy $e = 1$ is used and we have $ax'\alpha_k + y'\beta_k = 0 \bmod q$ (which can be checked by verifying that $(g^a)^{x'\alpha_k} g^{y'\beta_k} = 1$), then the meta-reduction computes the key as
$$K = \begin{cases} (g^c)^{x'y'\gamma_k}(g^b)^{\delta_k + \sum_{i<k} s_i\sigma_i} & \text{for } d = 0 \text{ and} \\ (g^c)^{z'\gamma_k}(g^b)^{\delta_k + \sum_{i<k} s_i\sigma_i} & \text{for } d = 1 \end{cases}$$
and immediately outputs 0 if $y_t = \mathsf{PRF}(K, 1)$; else it continues.

9. In any other case, if the reduction aborts prematurely or if the insertion strategy has been false, i.e., the choice of e does not match the condition on $x'\alpha_k + y'\beta_k \neq 0 \bmod q$, then output a random bit.

5.8 Analysis

For the analysis assume for the moment that our meta-reduction has actually chosen the index k of the first correct and non-trivial answer y_t, i.e., where $x\alpha_k + y\beta_k + z\gamma_k \neq 0 \bmod q$. Additionally, assume that $(x\alpha_k + y\beta_k) \neq 0 \bmod q$. Then $e = 0$ with probability at least $\frac{1}{2}$. This holds since the reduction remains perfectly oblivious about the choice of e, because all values in the interaction have the same distributions in both cases for e. Then the actual key for answering the test query is
$$(g^{u_k})^b = (g^{ab})^{x'\alpha_k + y'\beta_k}(g^{bz})^{\gamma_k}(g^b)^{\delta_k + \sum_{i<k} s_i\sigma_i}.$$
This implies that the meta-reduction's input g^c yields the same key K if $g^c = g^{ab}$ and hence equality for the PRF value. Yet, it yields a random value if g^c is random (since the exponent $x'\alpha_k + y'\beta_k$ does not vanish). In the latter case, since the value g^c is at no point used in the simulation before the reduction outputs y_t, the probability that y_t predicts $\mathsf{PRF}(K, 1)$ for the fresh random key, is negligible. This final step in the argument can be formalized straightforwardly.

Assume next that, besides the correct prediction of index k, we have $e = 1$ and $ax'\alpha_k + y'\beta_k = 0 \bmod q$. Then, since $ax'\alpha_k + y'\beta_k + z\gamma_k \neq 0 \bmod q$, we must have that $z\gamma_k \neq 0 \bmod q$ and therefore also $x'y'\gamma_k \neq 0 \bmod q$ for $d = 0$ resp. $z'\gamma_k \neq 0 \bmod q$ for $d = 1$. The same argument as in the previous case applies now. Namely, for $g^c = g^{ab}$ the meta-reduction computes the expected key, whereas for random g^c the contribution to the computed value K is for a non-zero exponent, such that equality for the PRF value only holds with negligible probability.

Putting the pieces together, for $g^c = g^{ab}$ our algorithm correctly outputs 0 if the reduction uses the adversary's help (with non-negligible probability $\text{Prob}\left[\overline{\mathsf{useless}}\right]$), if the prediction k is correct (with non-negligible probability $\frac{1}{q(\lambda)}$), and if the insertion strategy is correct (with probability at least $\frac{1}{2}$). Let
$$\epsilon(\lambda) \geq \frac{1}{2q(\lambda)} \cdot \text{Prob}\left[\overline{\mathsf{useless}}\right]$$
denote the non-negligible probability that the meta-reduction outputs 0 early. It also outputs 0 with probability $\frac{1}{2}$ in any other case, such that the probability of outputting 0 for $g^c = g^{ab}$ is at least
$$\epsilon(\lambda) + \frac{1}{2} \cdot (1 - \epsilon(\lambda)) \geq \frac{1}{2} + \frac{\epsilon(\lambda)}{2}.$$

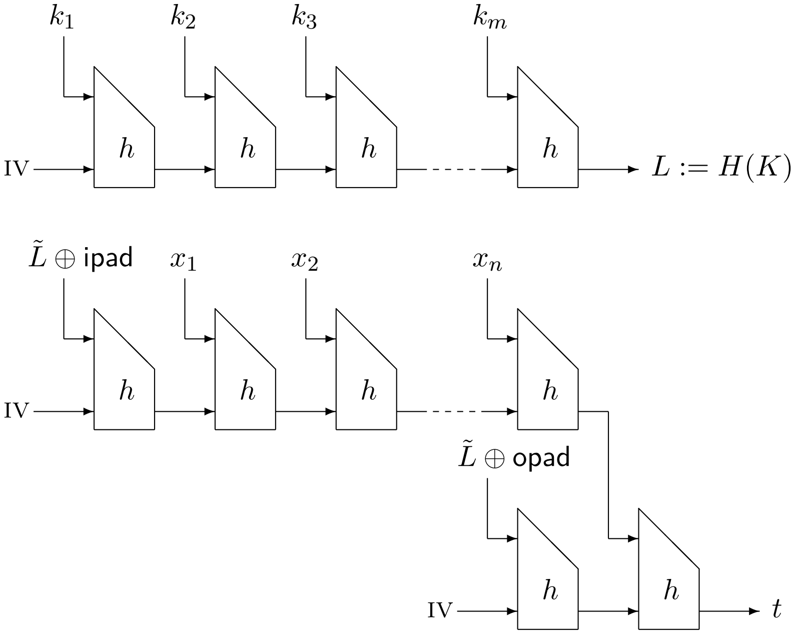
k_1
k_2
k_3
k_m
IV
h
$L := H(K)$
$\tilde{L} \oplus \mathsf{ipad}$
x_1
x_2
x_n
$\tilde{L} \oplus \mathsf{opad}$
t

Figure 3: Illustration of Merkle-Damgård computation of $t := \mathsf{HMAC}(K, X)$ with key $K = k_1 k_2 \ldots k_m$ of size $m \cdot b$ and label $X = x_1 x_2 \ldots x_n$ of size $n \cdot b$.

In case that g^c is random, our meta-reduction only outputs 0 if the PRF value matches y_t for the random key K, or if the final randomly chosen bit equals 0. The probability of this happening is only negligibly larger than $\frac{1}{2}$. This conversely means that the meta-reduction correctly outputs 1 in this case with probability at least $\frac{1}{2} - \text{negl}(\lambda)$ for some negligible function $\text{negl}(\lambda)$.

Overall, the probability of distinguishing the cases is at least

$$\frac{1}{2} \cdot \left(\frac{1}{2} + \frac{\epsilon(\lambda)}{2}\right) + \frac{1}{2} \cdot \left(\frac{1}{2} - \text{negl}(\lambda)\right) \geq \frac{1}{2} + \frac{\epsilon(\lambda)}{4} - \frac{\text{negl}(\lambda)}{2},$$

which is non-negligibly larger than $\frac{1}{2}$ for non-negligible $\epsilon(\lambda)$.

6 PRF-ODH Security of HMAC

In this section we examine the PRF-ODH security of HMAC [KBC97], augmenting previous results on the PRF security of HMAC [CDMP05, Kra10, BL15]. We show that $\mathsf{HMAC}(K, X)$ as well as its dual-PRF [Bel06] usage $\mathsf{HMAC}(X, K)$, as encountered in TLS 1.3 (see below), are mmPRF-ODH secure, which is our strongest notion of PRF-ODH security.

6.1 Description of HMAC

The basic construction of HMAC is illustrated in Figure 3. Let $h : \{0,1\}^c \times \{0,1\}^b \to \{0,1\}^c$ be a compression function and let $H : \{0,1\}^\star \to \{0,1\}^c$ be the cryptographic hash function associated with h. The iterated compression function is denoted by $h^\star : \{0,1\}^c \times B^+ \to \{0,1\}^c$ with B^+ the set of all bit strings of length $n \cdot b$ with $n \in \mathbb{N}^+$. On input key L and message $\tilde{M} = m_1 m_2 \ldots m_n$ of b-bit blocks, the output of $h^\star$ is a_n computed as

$$a_0 \quad = \quad L, \quad a_1 = h(a_0, m_1), \quad \ldots \quad a_n = h(a_{n-1}, m_n).$$

With the above convention we have $H(M) = h^\star(\mathsf{IV}, \tilde{M})$, where $\mathsf{IV} \in \{0,1\}^c$ is the initialization vector fixed by the description of H and $\tilde{M} \in B^+$ the message M padded to a multiple of the block size b. Finally, for key $K \in \{0,1\}^b$ and label X, HMAC is defined as $\mathsf{HMAC}(K, X) := H(K \oplus \mathsf{opad}||H(K \oplus \mathsf{ipad}||X))$, where opad and ipad are fixed (distinct) constants in $\{0,1\}^b$. In terms of the iterated compression function we have

$$\mathsf{HMAC}(K, X) = h^\star(\mathsf{IV}, K \oplus \mathsf{opad}||h^\star(\mathsf{IV}, K \oplus \mathsf{ipad}||X||\mathsf{padding})||\mathsf{padding}).$$

HMAC is in general also defined for keys whose lengths differs from the block size b. The minimal requirements from a security point of view is a length of c bits, i.e., the output length of the underlying hash function H. Keys of this minimal length are simply padded to the correct length. Shorter keys are not recommended from a security perspective. Longer keys K are first hashed down to $H(K) \in \{0,1\}^c$ and then padded. For ease of notation, we introduce the auxiliary function HMAC' that takes as input keys of length greater than b and is defined as $\mathsf{HMAC}'(K, X) := \mathsf{HMAC}(\tilde{L}, X)$ where $L := H(K)$ and $\tilde{L}$ is the padded value of L (cf. Figure 3).

6.2 Security of HMAC

Theorem 6.1. *Assume that the underlying compression function $h : \{0,1\}^c \times \{0,1\}^b \to \{0,1\}^c$ of* HMAC *is a random oracle. Then* HMAC *is* $\mathsf{mmPRF\text{-}ODH}$*-secure under the* StDH *assumption. More precisely, for any efficient adversary $\mathcal{A}$ against the* $\mathsf{mmPRF\text{-}ODH}$ *security of* HMAC*, there exists an efficient algorithm $\mathcal{B}$ such that*

$$\mathsf{Adv}^{\mathsf{mmPRF\text{-}ODH}}_{\mathsf{HMAC},\mathcal{A}} \leq \sqrt{\mathsf{Adv}^{\mathsf{StDH}}_{\mathbb{G},\mathcal{B}}} + (q_{\mathsf{RO}} + (q_{\mathsf{ODH}_u} + q_{\mathsf{ODH}_v}) \cdot \ell_{\mathsf{ODH}} + 1)^2 \cdot 2^{-c}$$

where q with the respective index denotes the maximal number of the according oracle queries, and ℓ_{ODH} the maximal number of oracle calls to h in each evaluation of any ODH *oracle call.*

Proof. Let the compression function h underlying HMAC be modeled as a random oracle RO. We note that we can assume each output of the compression function (genuine or simulated, as below) to be unique and to never hit the initialization vector IV. This assumption is reflected in the loss of the factor $(q_{\mathsf{RO}} + (q_{\mathsf{ODH}_u} + q_{\mathsf{ODH}_v}) \cdot \ell_{\mathsf{ODH}} + 1)^2 \cdot 2^{-c}$ in the above security statement, applying the birthday bound to the maximal number of queries (adding 1 for IV) and noting that in the simulation we simulate the same number of random oracle evaluations as in the actual attack.

A key insight for unique outputs of the compression function is that we can determine *evaluation chains* in a list of random oracle queries and answers. That is, starting from any input/output pair of the random oracle, we can try to go backwards along the HMAC iteration via the unique pre-images in the table, to check if we end up at a value (IV, L') for the hash function's initialization vector IV. Since we assume that no random oracle evaluation yields IV we can easily identify the beginning of such a chain. This holds for both the inner and outer branch of the HMAC evaluation such that we can check if a value has been derived as a full HMAC evaluation, for the key value L. From there on we can also check if we have a full evaluation chain of the key (if the key is larger than the block size). To distinguish the two cases we call the former an HMAC evaluation chain for key value L, and the latter a *complete* HMAC evaluation chain. They coincide for keys which fit into the block size.

For evaluation chains we can also extract the input string (and possibly the key input). In particular, at any point during the simulation we can take any entry in the simulated random oracle table and determine if there is an evaluation chain for this entry. This implies that we can determine all existing evaluation chains at any point in time. For some of such completed chains, we will only know an implicit presentation $[Q, R]$ of the key K with $\mathsf{DDH}(Q, R, K) = 1$, as in the proof of Theorem 3.3.

We now show that if there exists an adversary $\mathcal{A}$ against the $\mathsf{mmPRF\text{-}ODH}$-security of HMAC, then there necessarily also exists an adversary $\mathcal{B}$ that can break StDH with the corresponding advantage. In the following we discuss the case where the group $\mathbb{G}$ is such that its canonical bit string representation exceeds b bits and we use upstream hashing of the key. The cases concerning keys from groups with block sized representations (or even shorter) can be proven analogously.

The StDH-adversary $\mathcal{B}$ simulates the $\mathsf{mmPRF\text{-}ODH}$ environment for $\mathcal{A}$ and programs the random oracle. In addition to the $\mathsf{DDH}(g^u, \cdot, \cdot)$ oracle we assume that $\mathcal{B}$ has a claw-verifying oracle Claw which for inputs S, T checks that $S^u = T^v$, and that it can simulate the $\mathsf{DDH}(g^v, \cdot, \cdot)$ oracle. These extra oracles are accounted for as in the monolithic random oracle case by using the square root of the advantage against StDH.

Once $\mathcal{B}$ has obtained its challenge (g, g^u, g^v), it runs the $\mathsf{mmPRF\text{-}ODH}$-adversary $\mathcal{A}$ as a sub-routine on input (g, g^u). Adversary $\mathcal{A}$ then has access to the oracles RO and ODH_u (and later also ODH_v). Here, the random oracle RO corresponds to the compression function h and thus takes inputs of the form $(A, x) \in \{0,1\}^c \times \{0,1\}^b$. Oracle ODH_u takes as input $(A, X) \in \mathbb{G} \times \{0,1\}^\star$ and is supposed to return $\mathsf{HMAC}(A^u, X)$. $\mathcal{B}$ must provide sound simulations of these oracles. This is done as follows:

Simulation of ODH_u^m. Repeated queries (S, X) to ODH_u^m are answered consistently by returning the same value as before. Thus, in the following we can assume that the received query of the form (S, X) has not been queried to ODH_u^m beforehand.

Else, algorithm $\mathcal{B}$ first checks if there already exists a evaluation chain for a pair (K, X) such that $K = S^u$; this can be verified with the DDH oracle. If this is the case, $\mathcal{B}$ answers consistently with the final output of the evaluation chain.

If not, it verifies if there is an implicit key $K = [Q, R]$ with $Q = g^u$ and $R = S$, or with $Q = g^v$ and $R^v = S^u$, where the former can be checked directly and the latter can be verified with the help of oracle Claw.

- If $\mathcal{B}$ finds a matching key according to one of the cases above, it looks up the corresponding key value L as before.

- If there is no (explicitly or implicitly) matching key then $\mathcal{B}$ sets $H([g^u, S]) \leftarrow L$ to a uniformly random value $L \in \{0,1\}^c$. It stores the implicit key $[g^u, S]$ with the value L for future use.

$\mathcal{B}$ then iterates the HMAC computation to get the return value y with key L by calling its (simulated) oracle RO an all values. More precisely, $\mathcal{B}$ computes $y := \mathsf{RO}^\star(\mathsf{IV}, \tilde{L} \oplus \mathsf{opad} || \mathsf{RO}^\star(\mathsf{IV}, \tilde{L} \oplus \mathsf{ipad} || X || \mathsf{padding}) || \mathsf{padding})$, where $\tilde{L} = L || \mathsf{padding}$, and $\mathcal{B}$ then returns y as response. Note that, if the evaluation chain had already been computed before, the outcome of this evaluation is consistent with the previous result.

Simulation of ODH_v^m. Analogously to ODH_u^m.

Simulation of RO. Outputs of RO for equal input queries are answered consistently. If a previously unseen query, say, $(\tilde{A}, \tilde{x})$ is received, then $\mathcal{B}$ must consider if $\mathsf{RO}(\tilde{A}, \tilde{x})$ completes an $H([g^u, S])$ (or $H([g^v, S])$) computation for only implicitly known key $K = S^u$ (or $K = S^v$) that has been set beforehand to a uniformly random value, say, L (see the simulation of oracle ODH_u or ODH_v). This can be checked again with the DDH oracle. If this is the case, then the reduction answers the query consistently with value L. Otherwise, a response y is drawn uniformly at random from $\{0,1\}^c$ and returned.

At some point, $\mathcal{A}$ issues a challenge query $x^\star$. The reduction emulates the challenger by replying with g^v and some value $y^\star$, drawn uniformly at random from $\{0,1\}^c$. $\mathcal{A}$ can now query ODH_u^m and RO further, with the sole limitation that it may not query the pair $(g^v, x^\star)$ to ODH_u^m. Additionally, $\mathcal{A}$ acquires access to the ODH_v^m oracle which is simulated analogously to the ODH_u^m oracle and may not be queried with $(g^u, x^\star)$. Eventually, $\mathcal{A}$ stops and outputs a guess bit b'. If $\mathsf{DDH}(g^u, g^v, \tilde{K}) = 1$ for some completed chain of RO queries with associated key $\tilde{K}$, $\mathcal{B}$ outputs $\tilde{K}$ in the StDH game.

The rest of the proof is as before, given that the simulation of all oracle queries is sound. That is, $\mathcal{B}$ outputs the correct value g^{uv} with high probability if $\mathcal{A}$ wins $\mathsf{mmPRF\text{-}ODH}$ with non-negligible advantage. We show this by arguing that $\mathcal{A}$ can win the $\mathsf{mmPRF\text{-}ODH}$ game in the random oracle model with non-negligible advantage if and only if g^{uv} is an input key of a completed chain of RO queries with input (padded) label $x^\star$. To this end note that $\mathcal{A}$ expects $y^\star$ to be either $y_0 := \mathsf{HMAC}(g^{uv}, x^\star)$ or $y_1 \xleftarrow{\$} \{0,1\}^c$. By the nature of random oracles, y_0 and y_1 are indistinguishable for $\mathcal{A}$ since both are drawn uniformly at random from $\{0,1\}^c$. Even if $\mathcal{A}$ can correctly determine the value g^{uv}, it cannot compute y_0 by itself to compare with the received challenge. Thus, $\mathcal{A}$ must iteratively query RO on the full $\mathsf{HMAC}(g^{uv}, x^\star)$ computation, including the key g^{uv}, in order to distinguish y_0 and y_1. Furthermore, $\mathcal{B}$ is efficient, since $\mathcal{A}$ is efficient and asks at most polynomially many queries to each oracle. □

6.3 Application to HKDF

As mentioned earlier, one specific use case of the $\mathsf{PRF\text{-}ODH}$ assumption arises in the setting of TLS 1.3. Here, the HKDF scheme [Kra10, KE10] is adapted for key derivation. In particular, the function $\mathsf{HKDF.Extract}$ is used to derive an internal key K' as

$$K' \leftarrow \mathsf{HKDF.Extract}(X, K) := \mathsf{HMAC}(X, K),$$

where an adversarially known value X is used as the HMAC key while the secret randomness source in the form of a DH shared secret $K = g^{uv}$ is used as the label. At a first glance, this swapping of inputs may seem odd. However, the specified purpose of $\mathsf{HKDF.Extract}$ is to extract uniform randomness from its *second* component.

One way to prove that K' is indeed a random key (as long as g^{uv} is not revealed to the adversary) is to model $\mathsf{HKDF.Extract}(X, \cdot)$ as a random oracle. An alternative approach is pursued in [DFGS15b, DFGS16, FG17] where the authors prove the statement under the assumption that $\mathsf{HKDF.Extract}(\mathit{XTS}, \mathit{IKM}) = \mathsf{HMAC}(\mathit{XTS}, \mathit{IKM})$ is $\mathsf{PRF\text{-}ODH}$ secure when understood as a PRF keyed with $\mathit{IKM} \in \mathbb{G}$ (i.e., when the key is the *second* input). In this light, it is beneficial to show that $\mathsf{HMAC}(X, K)$ remains $\mathsf{PRF\text{-}ODH}$ secure for key $K \in \mathbb{G}$ and $X \in \{0,1\}^\star$.[8] Fortunately, our general treatment of $\mathsf{HMAC}(K, X)$ in Theorem 6.1 with arbitrarily long keys allows us to conclude the analogous result for $\mathsf{HMAC}(X, K)$ with swapped key and label.

Corollary 6.2. *Let $h : \{0,1\}^c \times \{0,1\}^b \to \{0,1\}^c$ be the underlying compression function of* HMAC. *If h is modeled as a random oracle, then* $\mathsf{HMAC}'(K, X) := \mathsf{HMAC}(X, K)$ *is* $\mathsf{mmPRF\text{-}ODH}$ *secure under* StDH.

Alternatively, one may wish to prove Theorem 6.1 along the results established by Coron et al. [CDMP05]. In more detail, Coron et al. showed that if the compression function h is modeled as a random oracle then a variant of HMAC can be shown to be *indifferentiable from a random oracle* in the sense of Maurer et al. [MRH04]. Krawczyk [Kra10] mentions that the above result can also immediately be applied to the unmodified plain HMAC design. However, we believe that providing a detailed proof of Theorem 6.1 is beneficial since it enables us to argue in a straightforward manner that the "reversed" result presented in Corollary 6.2, i.e., $\mathsf{HMAC}(X, K)$ with swapped label and key as inputs, also holds. Else one would need to check if HMAC with a hashed key would also be a random oracle.

In recent developments initiated by the NIST hash function competition it has been established that sponge-based constructions can be used to build cryptographic hash functions. We are confident that the

[8] Though formally defined for arbitrary length, recall that the minimal recommended length is c bits

proof of Theorem 6.1 can be adapted to achieve the same result for HMAC if the underlying cryptographic hash function H is replaced by a sponge-based construction such as SHA-3 with the random permutation π modeled as a random oracle.[9] This proof can also be established along the lines of Bertoni et al. [BDPV08] who provide results showing that the sponge construction is indifferentiable from a random oracle when being used with a random transformation or a random permutation.

7 Conclusion

To the best of our knowledge, this is the first systematic study of the relations between different variants of the PRF-ODH assumption which is prominently being used in the realm of analyzing major real-world key exchange protocols. We provide a generic definition of the PRF-ODH assumption subsuming those different variants and show separations between most of the variants. Our results give strong indications that instantiating the PRF-ODH assumption without relying on the random oracle methodology is a challenging task, even though it can be formalized in the standard model. In particular, we show that it is implausible to instantiate the assumption in the standard model via algebraic black-box reductions to DDH-augmented problems.

Despite our negative result, we emphasize that using the PRF-ODH assumption still provides some advantage over the StDH assumption in the random oracle model. Namely, it supports a modular approach to proving key exchange protocols to be secure, shifting the heavy machinery of random-oracle reductions to StDH in the context of complex key exchange protocols to a much simpler assumption. As the PRF-ODH naturally appears in such protocols and enables simpler proofs, it is still worthwhile to use the assumption directly.

Acknowledgments

We thank the anonymous reviewers for valuable comments. This work has been co-funded by the DFG as part of project S4 within the CRC 1119 CROSSING and as part of project D.2 within the RTG 2050 “Privacy and Trust for Mobile Users”.

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[9] SHA-3 is part of the Keccak sponge function family [BDPA11]. It has been standardized in the FIPS Publication 202 [NIS15], wherein it is explicitly approved for usage in HMAC.

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