

On the Impact of Known-Key Attacks on Hash Functions

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Abstract. Hash functions are often constructed based on permutations or blockciphers, and security proofs are typically done in the ideal permutation or cipher model. However, once these random primitives are instantiated, vulnerabilities of these instantiations may nullify the security. At ASIACRYPT 2007, Knudsen and Rijmen introduced known-key security of blockciphers, which gave rise to many distinguishing attacks on existing blockcipher constructions. In this work, we analyze the impact of such attacks on primitive-based hash functions. We present and formalize the weak cipher model, which captures the case a blockcipher has a certain weakness but is perfectly random otherwise. A specific instance of this model, considering the existence of sets of B queries whose XOR equals 0 at bit-positions C , where C is an index set, covers a wide range of known-key attacks in literature. We apply this instance to the PGV compression functions, as well as to the Grøstl (based on two permutations) and Shrimpton-Stam (based on three permutations) compression functions, and show that these designs do not seriously succumb to any differential known-key attack known to date.

Keywords. Hash functions, known-key security, Knudsen-Rijmen, PGV, Grøstl, Shrimpton-Stam, collision resistance, preimage resistance.

1 Introduction

Cryptographic hash functions are conventionally built on top of compression functions, and in turn on one or more blockciphers. Since the first appearance of such compression function $F(h, m) = \text{DES}_m(h)$ by Rabin [49] in the late 70s, many blockcipher-based functions appeared in the literature [23, 25, 29, 30, 40, 43, 48, 59]. These all enjoy security proofs in the ideal model, where the underlying ciphers are assumed to behave ideally. Characteristic to these designs is that the key input to the cipher depends on the input to the compression function, and that the key scheduling needs to be sufficiently strong. For instance, Biryukov et al. [6] derived a related-key attack on AES and claimed that it invalidates the security of the Davies-Meyer compression function when the underlying primitive is instantiated with AES. A more recent approach to compression function design is to base them on a limited number of permutations [8, 41, 42, 51, 57]. These permutations could be designed from scratch, or obtained by fixing a small set of keys and using a blockcipher for these keys only. Related- or chosen-key attacks on blockciphers do not help the adversary here, as the keys are fixed.

Known-Key Security of Blockciphers. While in the classical security models for blockciphers the key is secret and randomly drawn and the adversary's target is to distinguish the instantiation of the cipher from a random permutation (also known as (strong) pseudorandom permutation security), this notion does not apply if the key is known to the adversary. At ASIACRYPT 2007, Knudsen and Rijmen [27] introduced known-key security of blockciphers. Here, the key is presumed known, and the adversary succeeds in distinguishing if it identifies a structural property of the cipher. Andreeva et al. [1] proposed a way to formalize the known-key security of blockciphers based on the underlying primitives. The model is

derived from the indistinguishability framework [37] and hence all composition results carry over. Intuitively: suppose some cryptosystem F is proven to achieve a certain level of security in the ideal permutation model, and consider F' to be F with the permutations replaced by independent blockcipher instantiations. Then, F' achieves the same level of security as F , up to the known-key indistinguishability bound of the underlying blockciphers.

In [1], several blockcipher constructions are proven to be known-key indistinguishable, such as the multiple Even-Mansour cipher and 14 rounds of balanced Feistel with random functions (using a result of Holenstein et al. [24]). For such ciphers, the above approach works well, although for Even-Mansour the composition is trivial (one essentially replaces an ideal permutation by an ideal permutation) and for Feistel with 14 rounds security is only guaranteed up to $2^{n/32}$ queries, where n is the state size of the cipher.

Known-Key Attacks on Blockciphers. Knudsen and Rijmen also demonstrated that the Feistel network on n bits with 7 rounds (called “Feistel₇”) is *not* known-key indistinguishable [1, 27]: an adversary can generically find $2^{n/2}$ plaintext/ciphertext tuples (m, c) and (m', c') satisfying $\text{Ri}_{n/2}(m \oplus c \oplus m' \oplus c') = 0$ (where $\text{Ri}_r(x)$ outputs the r rightmost bits of x). This result has led to a wave of other known-key attacks on practical constructions, including generalized/extended variants of Feistel [1, 27, 47, 53, 56], reduced versions of AES or Rijndael [22, 27, 38, 44, 52], reduced variants of the blockciphers underlying SHA-2 and SHA-3 finalists BLAKE and Skein [2, 7, 31, 34, 61], and many more [3, 11, 12, 14, 17, 18, 28, 33, 46, 47, 54, 55]. This paper will mostly be concerned with differential known-key attacks, including rebound- and boomerang-based attacks (the majority of above-mentioned attacks). We highlight two results that are among the best-known ones and that exemplify the idea of the other attacks. Gilbert and Peyrin [22] used the rebound technique [39] to derive a known-key attack on 8 rounds of AES (called “AES₈”). It starts from the middle, and results in a differential trail with four active words in the beginning, and four at the end. These active words are overlapping at two positions, hence one could consider this result as two tuples (m, c) and (m', c') satisfying $m \oplus c \oplus m' \oplus c' = 0$ at $10n/16$ bit-positions. The adversary has $2^{15} \leq 2^{n/8}$ degrees of freedom in the attack, and for any choice it results in such a tuple with a certain probability. (The bound of $2^{n/8}$ is used for simplicity later on.) The second attack we highlight is by Yu et al. [61], who employ the boomerang technique [60] to attack 36 rounds of the blockcipher Threefish-512 (called “Threefish₃₆”) used in Skein. This attack results in four tuples $(m^1, c^1), \dots, (m^4, c^4)$ satisfying $m^1 \oplus \dots \oplus c^4 = 0$. The adversary has 2^n degrees of freedom, but any trial succeeds with probability approximately 2^{-454} . Therefore, the expected number of solutions is about $2^{n-454} \leq 2^{n/8}$. This attack is in fact a known-related-key attack, where a fixed difference in the key exists. For simplicity, we condone this, observing that an attack with *no* key difference must logically be harder.

In any of these cases, the traditional and commonly employed ideal cipher/permutation model falls short: results achieved in this model do not *necessarily* hold if the primitives are instantiated with Feistel₇, AES₈, Threefish₃₆, or any other known-key distinguishable cipher.

1.1 Our Contributions

In their seminal work, Knudsen and Rijmen state: “In some cases blockciphers are used with a key that is known to the adversary, and at least to a certain extent, the key is under the adversary’s control. Our attacks are quite relevant to this case.” We investigate this fundamental question whether known-key attacks invalidate the security of primitive-based hash functions, but we do so in a much more general way. At a high level, we present a model that goes beyond the traditional ideal cipher model as well as the principle of known-key attacks and that allows to generically analyze the impact of various weaknesses of blockciphers on various blockcipher- and permutation-based cryptosystems.

Model. A naive approach to analyzing the impact of known-key attacks would be to simply plug a certain blockcipher construction into a hash function and to analyze its security,

but this would be a devious and complex combinatorial task: for a function based on r permutations, plugging Feistel₇ into it would lead to $7r$ underlying primitive calls. Note that proving security of the Feistel construction itself is already extraordinarily hard [16, 24, 32]. Instead, we model the blockciphers in such a way that they behave randomly, except that an adversary can exploit the particular relation. More formally, we pose a certain predicate Φ , and we draw blockciphers randomly from the set of all ciphers *that comply with predicate Φ* . Throughout, we refer to this model as the “weak cipher model (WCM).” It corresponds to the ideal cipher model if Φ is trivial.

We present an explicit description of a random weak cipher for the case where Φ implies for each key k the existence of A sets of B queries $\{(k, m^1, c^1), \dots, (k, m^B, c^B)\}$ that comply with a certain condition φ . These ciphers are modeled to have three interfaces: forward queries, inverse queries, and predicate queries. Forward and inverse queries are as usual; on a predicate query, an adversary is given a set of B queries satisfying φ . Multiple technicalities are involved in this formalization. Most importantly, predicate Φ applies to tuples of queries, rather than single queries only, and some query responses may have a reduced entropy.

Above-mentioned known-key attacks are covered by our model if the condition φ states for some $C \subseteq \{1, \dots, n\}$ that

$$\text{Bits}_C(m^1 \oplus c^1 \oplus \dots \oplus m^B \oplus c^B) = 0, \quad (1)$$

where $\text{Bits}_C(x)$ outputs a string consisting of all bits of x whose index is in C . (In fact, our model is much more general: above-mentioned attacks aim to generate only *one* relation, while we allow an adversary to see multiple relations.) The value A usually depends on n and C is regularly a large subset. We consider B being a relatively small number (independent of n). For the above-mentioned attack on Feistel₇, $A = 2^{n/2}$, $B = 2$, and C corresponds to the rightmost $n/2$ bits. Similarly, the attacks on AES₈ (for $A = 2^{n/8}$, $B = 2$, and C a certain set of size $10n/16$) and Threefish₃₆ (for $A = 2^{n/8}$, $B = 4$, and $C = \{1, \dots, n\}$) are covered, and so are almost all known differential (rebound- or boomerang-based) known-key attacks. We remark that, on the other hand, the predicate is not well-suited for integral-based known-key attacks: upon a predicate query an attacker would receive $B \approx 2^n$ queries.

The weak cipher model is similar to an approach followed by Bresson et al. [15] for the indistinguishability analysis of the SHA-3 candidate Shabal if the underlying blockcipher shows some non-random behavior, and by Bouillaguet et al. [13] to analyze the indistinguishability security of SIMD when the underlying compression function is distinguishable from a random function. However, in both approaches, the underlying biased primitives were relatively easy to model. For instance in [15] (using our terminology), predicate Φ is a relation that holds for single queries only, and not for combinations of queries. This considerably simplifies the analysis: one can derive a bias β to measure the distance between primitive responses and fully random responses, and consider oracle responses to be drawn from a set of size at least $2^{n-\beta}$, and the original indistinguishability analysis carries over with minor modifications. The predicate used in the analysis in [13], on the other hand, *does* apply to tuples of queries, but the model can simply be described using two sampling algorithms, and an adversary cannot hit a weak pair by accident (which *is* possible in our analysis). Liskov [35] used a similar approach to prove indistinguishability security of the zipper hash if the underlying compression function is invertible up to a certain degree. However, the analysis is significantly simpler, as this primitive can be perfectly modeled. We finally remark that Katz et al. [26] analyze the impact of related-key attacks on blockciphers to hash functions. However, in their model, the differences $\Delta k, \Delta x, \Delta y$ are fixed, an ideal cipher is generated for half of the key space, and for the other half the cipher is adjusted as $E_k(x, y) = E_{k \oplus \Delta k}(x \oplus \Delta x) \oplus \Delta y$. This primitive can be easily modeled, but is also too generous to the attacker.

To our knowledge, this is the first attempt to formally analyze the effect of a wide class of blockcipher attacks on higher level cryptographic functions. Nonetheless, the weak cipher model is in essence still a model: we use an abstraction of the cryptanalytic known-key attacks in such a way that the ideal cipher model can be relaxed to cope them. A further discussion on the accuracy of the model is given in Sect. 7.

Table 1. Security results for the PGV, Grøstl, and Shrimpton-Stam compression functions in the weak cipher model. Ideal cipher/permutation model bounds match the ones of $B \geq 3$. All results are tight except for the case ($B = 1, |C| > n/2$) for Shrimpton-Stam.

B	$ C $	PGV		Grøstl		Shrimpton-Stam	
		collision	preimage	collision	preimage	collision	preimage
1	$\leq n/2$	$2^{(n- C)/2}$	$2^{n- C }$	$2^{(n- C)/4}$	$2^{(n- C)/2}$	$2^{(n- C)/2}$	$2^{n/2}$
	$> n/2$	$2^{(n- C)/2}$	$2^{n- C }$	$2^{(n- C)/4}$	$2^{(n- C)/2}$	$2^{(n- C)/2}$	$2^{n- C }$
2	$\leq n/2$	$2^{n/2}$	2^n	$2^{n/4}$	$2^{n/2}$	$2^{n/2}$	$2^{n/2}$
	$> n/2$	$2^{n- C }$	2^n	$2^{(n- C)/2}$	$2^{n/2}$	$2^{n- C }$	$2^{n/2}$
≥ 3	arbitrary	$2^{n/2}$	2^n	$2^{n/4}$	$2^{n/2}$	$2^{n/2}$	$2^{n/2}$

Application to Blockcipher-Based Hash Functions. Preneel, Govaerts, and Vandewalle (PGV) [48] classified the 64 most basic ways of constructing a $2n$ -to- n -bit compression function from a blockcipher with n -bit key and n -bit state, and claimed security of 12 of them. A formal security analysis of these functions in the ICM has been performed by Black et al. [9], and later by Duo and Li [19], Stam [59], and Black et al. [10]. In more detail, in the ICM these constructions achieve tight collision security up to about $2^{n/2}$ queries and preimage security up to about 2^n queries. Baecher et al. [4] recently showed that the 12 secure PGV functions can be divided into two classes, in such a way that if a primitive makes one function secure it makes the entire class secure.

As first application of our model, we consider the PGV compression functions in the WCM and derive collision and preimage bounds for general (A, B, C) . A schematic summary of the results for various B and C is given in Table 1 (we remark that A is merely a technical parameter that has no influence on the results). We also show that the bounds are optimal, by providing matching attacks. Some of these attacks are similar to methods used in [27, 53, 56] to detect (near-)collisions in certain PGV modes of operations using known-key attacks.

Application to Permutation-Based Hash Functions. We also apply the WCM to permutation-based compression functions. This is particularly interesting for two reasons: (i) it allows us to understand the impact of distinguishers on permutations that are used in hash functions, and (ii) a blockcipher with a fixed and known key is a permutation and can be used as such. In more detail, we consider the Grøstl compression function [21] and the permutation-based equivalent of the Shrimpton-Stam compression function [57] (see also Fig. 4). In the IPM, the former is proven to achieve collision security up to $2^{n/4}$ queries, where n is the state size, and preimage security up to $2^{n/2}$ [20]. Rogaway and Steinberger [51] showed via an automated analysis that the latter function is collision and preimage resistant up to $2^{n/2}$ queries (asymptotically). This has been confirmed in the generalized work of Mennink and Preneel [41].

A summary of our findings for the Grøstl and Shrimpton-Stam compression functions in the WCM is given in Table 1. All results are tight, except for the case ($B = 1, |C| > n/2$) for Shrimpton-Stam, for which we leave proving tightness as an open problem. We remark that the analysis for these schemes is much more demanding as multiple primitives are involved.

Impact. An application of our formalization to the PGV functions and various permutation-based functions shows that these achieve a comparable level of security in the ideal and weak cipher model for a spectrum of choices for (A, B, C) . This result particularly implies that most relevant rebound-based (including [12, 22, 28, 38, 52, 53, 56]) and boomerang-based (including [2, 7, 31, 54, 61]) known-key attacks known to date do not invalidate the security of such functions, or only have a little effect. For instance, the above-discussed attack on Feistel₇ satisfies $B = 2$ and $|C| = n/2$ and it does not affect the security; similarly for Threefish₃₆ for which $B = 4$. The attack on AES₈ is covered for $B = 2$ and $|C| = 10n/16$,

which demonstrates a slight security degradation to $2^{6n/16}$ for the PGV functions, but this may in part be due to our over-generosity to the adversary. We remark that, even though we focused on collision and preimage resistance, the techniques can be generalized to other security notions, such as near-collisions. This may entail differences in the security results.

We stress that these results do not mean that the analyzed functions are secure when the underlying permutations are instantiated with, say, Feistel₇ or Threefish₃₆: it only means that existing known-key attacks, or more general weaknesses such as relation (1), *alone* are not sufficient to invalidate the collision and preimage security of the construction. Indeed, more sophisticated attacks which are not yet covered by our application of the WCM may still invalidate the security of certain modes [6]. It remains a challenging open research problem to generalize the findings to underlying primitives that have multiple or different weaknesses.

1.2 Outline

In Sect. 2, we formally present the “weak cipher model,” and in Sect. 3 we show how it relates to known-key attacks. We apply the model to the PGV functions in Sect. 4, to the Grøstl compression function in Sect. 5, and to Shrimpton-Stam in Sect. 6. We conclude this work in Sect. 7.

2 Weak Cipher Model

If X is a set, by $x \stackrel{\$}{\leftarrow} X$ we denote the uniformly random sampling of an element from X . By $X \stackrel{\cup}{\leftarrow} x$, we denote $X \leftarrow X \cup \{x\}$. For a bit string x , its bits are numbered $x = x_{|x|} \cdots x_2 x_1$. If $C \subseteq \{1, \dots, |x|\}$, the function $\text{Bits}_C(x)$ outputs a string consisting of all bits of x whose index is in C . Abusing notation, $\text{Bits}_{\overline{C}}(x)$ always denotes the remaining bits (technically, $\overline{C} = \{1, \dots, |x|\} \setminus C$). For $0 \leq r \leq |x|$, we consider $\text{Ri}_r(x)$ that outputs the r rightmost bits of x . In other words, $\text{Ri}_r(x) = \text{Bits}_{\{1, \dots, r\}}(x)$. For a function f , by $\text{dom}(f)$ and $\text{rng}(f)$ we denote its domain and range, respectively.

2.1 Security Model

For $\kappa \geq 0$ and $n \geq 1$, by $\text{BC}(\kappa, n)$ we denote the set of all blockciphers with κ -bit key operating on n bits. If $\kappa = 0$, $\text{BC}(n) := \text{BC}(0, n)$ denotes the set of all n -bit permutations. If Φ is a predicate, by $\text{BC}[\Phi](\kappa, n)$ we denote the subset of ciphers of $\text{BC}(\kappa, n)$ that satisfy predicate Φ . For $\pi \in \text{BC}[\Phi](\kappa, n)$, the input-output tuples are denoted (k, x, z) , where $\pi(k, x) = \pi_k(x) = z$ and $\pi^{-1}(k, z) = \pi_k^{-1}(z) = x$. The key k is omitted in case $\kappa = 0$.

Let $F : \{0, 1\}^s \rightarrow \{0, 1\}^n$ be a compressing function instantiated with $\ell \geq 1$ primitives from $\text{BC}[\Phi](\kappa, n)$, for some predicate Φ . Throughout, we consider security of F in an idealized model: we consider an adversary \mathcal{A} that is a probabilistic algorithm with oracle access to a randomly sampled primitive $\pi = (\pi_1, \dots, \pi_\ell) \stackrel{\$}{\leftarrow} \text{BC}[\Phi](\kappa, n)^\ell$. \mathcal{A} is information-theoretic and its complexity is only measured by the number of queries made to its oracles. The adversary can make forward and inverse queries to its oracles, and these queries are stored in a query history \mathcal{Q} .

A collision-finding adversary \mathcal{A} for F aims at finding two distinct inputs to F that compress to the same range value. In more detail, we say that \mathcal{A} succeeds if it finds two distinct inputs X, X' such that $F(X) = F(X')$ and \mathcal{Q} contains all queries required for these evaluations of F . We define by

$$\text{Adv}_F^{\text{col}}(\mathcal{A}) = \Pr \left(\pi \stackrel{\$}{\leftarrow} \text{BC}[\Phi](\kappa, n)^\ell, X, X' \leftarrow \mathcal{A}^\pi : X \neq X' \wedge F(X) = F(X') \right)$$

the probability that \mathcal{A} succeeds in this. By $\text{Adv}_F^{\text{col}}(q)$ we define the maximum collision advantage taken over all adversaries making q queries.

For preimage resistance, we focus on everywhere preimage resistance [50], which captures preimage security for every point of $\{0, 1\}^n$. Let $Z \in \{0, 1\}^n$ be any range value. Then, we say that \mathcal{A} succeeds in finding a preimage if it obtains an input X such that $F(X) = Z$ and \mathcal{Q} contains all queries required for this evaluation of F . We define by

$$\mathbf{Adv}_F^{\text{epre}}(\mathcal{A}) = \max_{Z \in \{0, 1\}^n} \Pr \left(\pi \stackrel{\$}{\leftarrow} \text{BC}[\Phi](\kappa, n)^\ell, X \leftarrow \mathcal{A}^\pi(Z) : F(X) = Z \right)$$

the probability that \mathcal{A} succeeds, maximized over all possible choices for Z . By $\mathbf{Adv}_F^{\text{epre}}(q)$ we define the maximum (everywhere) preimage advantage taken over all adversaries making q queries.

If Φ is a trivial relation, we have $\text{BC}[\Phi](\kappa, n) = \text{BC}(\kappa, n)$, and the above definitions boil down to security in the ideal cipher model (ICM) if $\kappa > 0$ or the ideal permutation model (IPM) if $\kappa = 0$. On the other hand, if Φ is a non-trivial predicate, it strictly reduces the set $\text{BC}(\kappa, n)$. In this case, we will refer to the model as the “weak cipher model (WCM),” for both $\kappa > 0$ and $\kappa = 0$. Very informally, this model still involves random ciphers/permutations, with the difference that an adversary may exploit a certain additional property. The modeling of a randomly drawn weak ciphers is much more delicate.

2.2 Random Weak Cipher

For a certain class of predicates, we discuss how to model a randomly drawn weak cipher π from $\text{BC}[\Phi](\kappa, n)$. Let $A, B \in \mathbb{N}$. We will consider predicates that imply, for every $k \in \{0, 1\}^\kappa$, the existence of A sets of B distinct queries $\{(x^1, z^1), \dots, (x^B, z^B)\}$ that satisfy $\varphi_k(\{(x^1, z^1), \dots, (x^B, z^B)\})$ for some condition φ depending on key k . The predicate is denoted $\Phi(A, B, \varphi)$. A is merely a technical parameter, and throughout we assume it is larger than q , the number of oracle calls an adversary can make. This definition of $\Phi(A, B, \varphi)$ is fairly general. Particularly, predicate B -sets may overlap and the condition φ can represent any function on the inputs. We note that Φ can be easily generalized to tuples of different length and/or to multiple types of conditions at the same time.

Traditionally, an adversary has only forward $\pi_k(x)$ and inverse $\pi_k^{-1}(z)$ query access. In order for the adversary to be able to exploit the weakness present in π , we give it additional access to π via a “predicate query” $\pi_k^\Phi(y)$: on input of $y \in \{1, \dots, A\}$, the adversary obtains a B -set $\{(x^1, z^1), \dots, (x^B, z^B)\}$ that satisfies $\varphi_k(\{(x^1, z^1), \dots, (x^B, z^B)\})$.

A formal description of how to model $\pi \stackrel{\$}{\leftarrow} \text{BC}[\Phi(A, B, \varphi)](\kappa, n)$ is given in Fig. 1. Here, for every $k \in \{0, 1\}^\kappa$, P_k is an initially empty list of π_k -evaluations, where a regular forward/inverse query adds one element (x, z) to P_k and a π_k^Φ -query may add up to B elements. Additionally, P_k^Φ is an initially empty list of queries to π_k^Φ . We denote by $\Sigma_k(P_k, P_k^\Phi) \subseteq (\{0, 1\}^n \times \{0, 1\}^n)^B$ the set of all tuples $\{(x^1, z^1), \dots, (x^B, z^B)\}$ such that

- (i) x^1, \dots, x^B are pairwise distinct and z^1, \dots, z^B are pairwise distinct;
- (ii) $\forall_{\ell=1}^B : x^\ell \in \text{dom}(P_k) \implies z^\ell = P_k(x^\ell)$ and $z^\ell \in \text{rng}(P_k) \implies x^\ell = P_k^{-1}(z^\ell)$;
- (iii) $\varphi_k(\{(x^1, z^1), \dots, (x^B, z^B)\})$ holds;
- (iv) $\{(x^{p(1)}, z^{p(1)}), \dots, (x^{p(B)}, z^{p(B)})\} \notin \text{rng}(P_k^\Phi)$ for any permutation p on $\{1, \dots, B\}$.

For a new query $\pi_k^\Phi(y)$, the response is then randomly drawn from $\Sigma_k(P_k, P_k^\Phi)$. Conditions (i-iii) are fairly self-evident; note particularly that an existing $(x, z) \in P_k$ may appear in multiple predicate queries. Condition (iv) assures that the drawing from $\Sigma_k(P_k, P_k^\Phi)$ is not just an old predicate query or a reordering thereof. The usage of this set $\Sigma_k(P_k, P_k^\Phi)$ allows for a uniform behavior of π_k^Φ for every k , and in general of $\pi \stackrel{\$}{\leftarrow} \text{BC}[\Phi(A, B, \varphi)](\kappa, n)$, modulo the known existence of condition φ . This step is fundamental to our model and new compared with previous approaches of [13, 15, 35]. We remark that the model allows adversaries to make their queries at their own discretion, e.g., duplicate queries and regular queries after predicate queries are allowed.

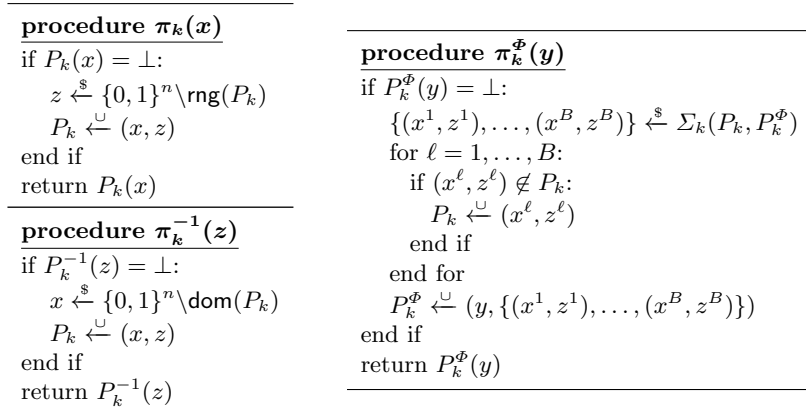


Fig. 1. Random weak cipher π . An adversary has access to π, π^{-1} , and π^Φ .

Note that, for simplicity of analysis, we detached the drawing in π_k and $\pi_k^{-1}(z)$ from $\pi_k^\Phi(y)$. More formally, the drawing in the former sets do not take into account the existence of tuples in $\Sigma_k(P_k, P_k^\Phi)$. This is merely for simplicity of analysis, but it quite accurately captures a random weak cipher as long as $|\Sigma_k(P_k, P_k^\Phi)| \gg 0$. This is particularly the case if we require that $q + AB \ll 2^n$. We thank Damian Vizár for pointing out the necessity of this side condition.

2.3 Random Abortable Weak Cipher

Security analyses in the WCM are significantly more complex than in the ICM or IPM, which is in part because predicate queries may consist of older queries. This will particularly be an issue once collisions among queries are investigated. To suit the analysis for this case, we transform the WCM to an abortable weak cipher model (AWCM), which we denote as $\overline{\text{BC}}[\Phi(A, B, \varphi)](\kappa, n)$. At a high-level, an abortable weak cipher responds to predicate queries with *new* query tuples only, and aborts once it turns out that an older query appears in a newer predicate query.

For any $k \in \{0, 1\}^\kappa$ and partial P_k and P_k^Φ , define by $\bar{\Sigma}_k(P_k^\Phi) \subseteq (\{0, 1\}^n \times \{0, 1\}^n)^B$ the set of all tuples $\{(x^1, z^1), \dots, (x^B, z^B)\}$ such that

- (iii) $\varphi_k(\{(x^1, z^1), \dots, (x^B, z^B)\})$ holds;
- (iv) $\{(x^{p(1)}, z^{p(1)}), \dots, (x^{p(B)}, z^{p(B)})\} \notin \text{rng}(P_k^\Phi)$ for any permutation p on $\{1, \dots, B\}$.

$\bar{\Sigma}_k(P_k^\Phi)$ differs from $\Sigma(P_k, P_k^\Phi)$ in that conditions (i) and (ii) are omitted, and particularly: it is independent of P_k . A formal description of a random cipher $\bar{\pi} \stackrel{\$}{\leftarrow} \overline{\text{BC}}[\Phi(A, B, \varphi)](\kappa, n)$ is given in Fig. 2. It deviates from Fig. 1 as follows: for every key k , $\bar{\pi}_k^\Phi$ responds randomly from $\bar{\Sigma}_k(P_k^\Phi)$, and it aborts if the response violates one of the two skipped conditions of $\Sigma_k(P_k, P_k^\Phi)$.

The next lemma shows that the WCM and AWCM are indistinguishable as long as the abortable weak cipher does not abort, approximately up to the birthday bound. Here, we assume that $\bar{\Sigma}_k(P_k^\Phi)$ is always large enough.

Lemma 1. *Let $\bar{\pi} \stackrel{\$}{\leftarrow} \overline{\text{BC}}[\Phi(A, B, \varphi^C)](\kappa, n)$. Consider an adversary that makes q queries to $\bar{\pi}$. Then,*

$$\Pr(\bar{\pi} \text{ sets abort}) \leq \frac{B^2 q(q+1)}{2^n - \frac{B! q 2^n}{|\bar{\Sigma}_k(\emptyset)|}}.$$

<hr/> <p>procedure $\bar{\pi}_k(x)$ if $P_k(x) = \perp$: $z \xleftarrow{\\$} \{0, 1\}^n \setminus \text{rng}(P_k)$ $P_k \xleftarrow{\cup} (x, z)$ end if return $P_k(x)$</p> <hr/> <p>procedure $\bar{\pi}_k^{-1}(z)$ if $P_k^{-1}(z) = \perp$: $x \xleftarrow{\\$} \{0, 1\}^n \setminus \text{dom}(P_k)$ $P_k \xleftarrow{\cup} (x, z)$ end if return $P_k^{-1}(z)$</p> <hr/>	<hr/> <p>procedure $\bar{\pi}_k^\Phi(y)$ if $P_k^\Phi(y) = \perp$: $\{(x^1, z^1), \dots, (x^B, z^B)\} \xleftarrow{\\$} \bar{\Sigma}_k(P_k^\Phi)$ for $\ell = 1, \dots, B$: if $x^\ell \in \text{dom}(P_k) \wedge z^\ell \neq P_k(x^\ell)$: abort if $z^\ell \in \text{rng}(P_k) \wedge x^\ell \neq P_k^{-1}(z^\ell)$: abort if $(x^\ell, z^\ell) \in \{(x^1, z^1), \dots, (x^{\ell-1}, z^{\ell-1})\}$: abort if $(x^\ell, z^\ell) \notin P_k$: $P_k \xleftarrow{\cup} (x^\ell, z^\ell)$ end if end for $P_k^\Phi \xleftarrow{\cup} (y, \{(x^1, z^1), \dots, (x^B, z^B)\})$ end if return $P_k^\Phi(y)$</p> <hr/>
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Fig. 2. Random abortable weak cipher $\bar{\pi}$. An adversary has access to $\bar{\pi}, \bar{\pi}^{-1}$, and $\bar{\pi}^\Phi$.

Proof. Consider the i^{th} query, for $i \in \{1, \dots, q\}$, and assume it is a predicate query $\bar{\pi}_k^\Phi(y)$. We will consider the probability that this query makes $\bar{\pi}$ abort, provided it has not aborted so far. Prior to this i^{th} query, $|P_k| \leq B(i-1)$ and $|P_k^\Phi| \leq i$. Basic combinatorics shows that

$$|\bar{\Sigma}_k(P_k^\Phi)| = |\bar{\Sigma}_k(\emptyset)| - B! \cdot |P_k^\Phi|,$$

where we use that $\bar{\pi}$ has not aborted so far. This i^{th} query aborts only if for some $\ell \in \{1, \dots, B\}$, the value x^ℓ equals an element in $\text{dom}(P_k) \cup \{x^1, \dots, x^{\ell-1}\}$ or the value z^ℓ equals an element in $\text{rng}(P_k) \cup \{z^1, \dots, z^{\ell-1}\}$.

Define by $\bar{\Sigma}_k^{\text{abort}}(P_k^\Phi)$ the set of all elements of $\bar{\Sigma}_k(P_k^\Phi)$ that would lead to **abort**. We have $2B$ possible values to cause the abort (namely, x^1, \dots, z^B), and it causes the abort if it equals an element in a set of size at most $|P_k| + B$. For any of these $2B(|P_k| + B)$ choices, the number of tuples in $\bar{\Sigma}_k(P_k^\Phi)$ complying with this choice is at most $\frac{|\bar{\Sigma}_k(\emptyset)|}{2^n}$. Thus,

$$\Pr(\bar{\pi}^\Phi(y) \text{ sets abort}) = \frac{|\bar{\Sigma}_k^{\text{abort}}(P_k^\Phi)|}{|\bar{\Sigma}_k(P_k^\Phi)|} \leq \frac{2B(|P_k| + B) \cdot \frac{|\bar{\Sigma}_k(\emptyset)|}{2^n}}{|\bar{\Sigma}_k(\emptyset)| - B! \cdot |P_k^\Phi|} \leq \frac{2B^2 i}{2^n - \frac{B! q 2^n}{|\bar{\Sigma}_k(\emptyset)|}}.$$

The proof is completed by summation over $i = 1, \dots, q$. □

3 Modeling Known-Key Attacks

We next apply the WCM to known-key attacks. For the sake of explanation, we first reconsider the Knudsen-Rijmen attack on Feistel₇ [27]. (A detailed description of the attack is given in App. A.) Let $n \in \mathbb{N}$, and let $\pi := \pi_k$ be an instance of Feistel₇ with fixed key k . Knudsen and Rijmen revealed four functions $f, f', g, g' : \{0, 1\}^{n/2} \rightarrow \{0, 1\}^n$ such that for all $y \in \{0, 1\}^{n/2}$:

$$\begin{aligned} g(y) &= \pi(f(y)) \text{ and } g'(y) = \pi(f'(y)), \\ \text{Ri}_{n/2}(f(y) \oplus g(y)) &= \text{Ri}_{n/2}(f'(y) \oplus g'(y)). \end{aligned} \tag{2}$$

These four functions correspond to the equations of (9) in App. A and depend on the cryptographic primitive underlying Feistel₇ in a complicated way. Therefore, we can safely assume that these functions behave sufficiently random, besides this particular relation (2), and that they are unknown to the adversary. f, f', g, g' are all injective and satisfy $f(y) \neq$

$f'(y)$ and $g(y) \neq g'(y)$ for all y . On the other hand, collisions of the form $f(y) = f'(y')$ and $g(y) = g'(y')$ may occur.

Generically, the attack demonstrates that for key k there exist $2^{n/2}$ possibly overlapping sets of distinct queries $\{(x^1, z^1), (x^2, z^2)\}$ that satisfy $\text{Ri}_{n/2}(x^1 \oplus z^1 \oplus x^2 \oplus z^2) = 0$. In other words, Feistel₇ meets predicate $\Phi(2^{n/2}, 2, \varphi^{\text{Feistel}_7})$, where

$$\varphi_k^{\text{Feistel}_7}(\{(x^1, z^1), (x^2, z^2)\}) : \text{Ri}_{n/2}(x^1 \oplus z^1 \oplus x^2 \oplus z^2) = 0.$$

Here, we remark that the Knudsen-Rijmen attack works for *any* fixed but known key k , and that condition $\varphi_k^{\text{Feistel}_7}$ is in fact independent of the key. In this work, we will consider a more general predicate $\Phi(A, B, \varphi^C)$ for $A, B \in \mathbb{N}$ and $C \subseteq \{1, \dots, n\}$, where

$$\varphi^C(\{(x^1, z^1), \dots, (x^B, z^B)\}) : \text{Bits}_C(x^1 \oplus z^1 \oplus \dots \oplus x^B \oplus z^B) = 0. \quad (3)$$

This generalized predicate considers the case of arbitrary but fixed and known keys, where the adversary can even choose the key every time it makes a predicate query. Note that also the attacks on AES₈ and Threefish₃₆ (see Sect. 1) are covered, as they satisfy $\Phi(2^{n/8}, 2, \varphi^C)$ for certain C of size $10n/16$ and $\Phi(2^{n/8}, 4, \varphi^{\{1, \dots, n\}})$, respectively. In general, all rebound- or boomerang-based known-key attack in literature are covered by predicate $\Phi(A, B, \varphi^C)$ for some A, B, C . Here, B is always a value independent of n (usually 2 or 4) and C is regularly a large subset (of size at least $n/4$). Throughout, we consider A to be sufficiently large.

Basic Computations for AWCM

For the specific condition φ^C of (3), we derive a simpler bound on the probability that a primitive $\bar{\pi} \stackrel{\S}{\leftarrow} \overline{\text{BC}}[\Phi(A, B, \varphi^C)](\kappa, n)$ aborts, along with some other elementary observations for $\bar{\pi}$. To this end, we define the notation “[X],” which equals 1 if X holds and 0 otherwise. For conciseness, we introduce the function $\delta_{B,C}[b]$ defined as

$$\delta_{B,C}[b] = 2^{|C|}[B = b] + [B > b] = \begin{cases} 2^{|C|} & \text{if } B = b, \\ 1 & \text{if } B > b, \\ 0 & \text{otherwise.} \end{cases} \quad (4)$$

Lemma 2. *Let $\bar{\pi} \stackrel{\S}{\leftarrow} \overline{\text{BC}}[\Phi(A, B, \varphi^C)](\kappa, n)$. Consider an adversary that makes $q \leq 2^{n-1}/B$ queries to $\bar{\pi}$. Then,*

$$\Pr(\bar{\pi} \text{ sets abort}) \leq \frac{B^2 q (q + 1)}{2^n - Bq}. \quad (5)$$

Let $k \in \{0, 1\}^\kappa$ and let $Z, Z', Z'' \in \{0, 1\}^n$. Consider any new query $\bar{\pi}_k^\Phi(y)$ and assume it does not abort. Write the response as $\{(x^1, z^1), \dots, (x^B, z^B)\}$. Then,

- (i) $\forall a \in \{1, \dots, B\} : \Pr(x^a = Z), \Pr(z^a = Z) \leq \frac{1}{2^{n-Bq}}$;
- (ii) $\forall a \in \{1, \dots, B\} : \Pr(x^a \oplus z^a = Z) \leq \frac{\delta_{B,C}[1]}{2^{n-Bq}}$;
- (iii) $\forall \{a, b\} \subseteq \{1, \dots, B\} : \Pr(x^a \oplus z^a = Z \wedge x^b \oplus z^b = Z') \leq \frac{\delta_{B,C}[2]}{2^{2n-Bq}}$;
- (iv) $\forall \{a, b\} \subseteq \{1, \dots, B\} : \Pr(x^a = Z \wedge x^b = Z' \wedge x^a \oplus z^a \oplus x^b \oplus z^b = Z'') \leq \frac{\delta_{B,C}[2]}{2^{3n-Bq}}$.

Proof. Recall from the proof of Lem. 1 that

$$|\bar{\Sigma}_k(P_k^\Phi)| = |\bar{\Sigma}_k(\emptyset)| - B!|P_k^\Phi|,$$

where $|P_k^\Phi| \leq q$. For the specific predicate analyzed in this lemma, $|\bar{\Sigma}_k(\emptyset)| = (2^n)^{2B-1} 2^{n-|C|}$. In the remainder, we regularly bound $B! \leq B \cdot (2^n)^{2B-2}$ for $B \geq 1$ or $B! \leq B \cdot (2^n)^{2B-4}$ for $B \geq 2$.

Probability of abortion. The bound of (5) directly follows from Lem. 1, the above-mentioned size of $\bar{\Sigma}_k(\emptyset)$, and the bound on $B!$.

Part (i). Define by $\bar{\Sigma}_k^{(i)}(P_k^\Phi)$ the set of all elements of $\bar{\Sigma}_k(P_k^\Phi)$ that satisfy $x^a = Z$. Then, $|\bar{\Sigma}_k^{(i)}(P_k^\Phi)| \leq (2^n)^{2B-2}2^{n-|C|}$, and

$$\Pr(x^a = Z) = \frac{|\bar{\Sigma}_k^{(i)}(P_k^\Phi)|}{|\bar{\Sigma}_k(P_k^\Phi)|} \leq \frac{1}{2^n - Bq}.$$

A similar analysis applies to the case $z^a = Z$.

Part (ii). Define by $\bar{\Sigma}_k^{(ii)}(P_k^\Phi)$ the set of all elements of $\bar{\Sigma}_k(P_k^\Phi)$ that satisfy $x^a \oplus z^a = Z$. We make a distinction between $B = 1$ and $B > 1$. In case $B > 1$, a similar reasoning as in (i) applies, and we have $|\bar{\Sigma}_k^{(ii)}(P_k^\Phi)| \leq (2^n)^{2B-2}2^{n-|C|}$. On the other hand, if $B = 1$, we have $|\bar{\Sigma}_k^{(ii)}(P_k^\Phi)| = 0$ if $\text{Bits}_C(Z) \neq 0$ and $|\bar{\Sigma}_k^{(ii)}(P_k^\Phi)| \leq 2^n$ if $\text{Bits}_C(Z) = 0$. In any case,

$$|\bar{\Sigma}_k^{(ii)}(P_k^\Phi)| \leq (2^n)^{2B-2}2^{n-|C|}\delta_{B,C}[1],$$

and

$$\Pr(x^a \oplus z^a = Z) = \frac{|\bar{\Sigma}_k^{(ii)}(P_k^\Phi)|}{|\bar{\Sigma}_k(P_k^\Phi)|} \leq \frac{\delta_{B,C}[1]}{2^n - Bq}.$$

Part (iii). This part only applies to $B > 1$; if $B = 1$ the probability equals 0 by construction. Define by $\bar{\Sigma}_k^{(iii)}(P_k^\Phi)$ the set of all elements of $\bar{\Sigma}_k(P_k^\Phi)$ that satisfy $x^a \oplus z^a = Z$ and $x^b \oplus z^b = Z'$. We make a distinction between $B = 2$ and $B > 2$. In case $B > 2$, a similar reasoning as in (i) and (ii) applies, and we have $|\bar{\Sigma}_k^{(iii)}(P_k^\Phi)| \leq (2^n)^{2B-3}2^{n-|C|}$. On the other hand, if $B = 2$, we have $|\bar{\Sigma}_k^{(iii)}(P_k^\Phi)| = 0$ if $\text{Bits}_C(Z \oplus Z') \neq 0$ and $|\bar{\Sigma}_k^{(iii)}(P_k^\Phi)| \leq (2^n)^2$ if $\text{Bits}_C(Z \oplus Z') = 0$. In any case,

$$|\bar{\Sigma}_k^{(iii)}(P_k^\Phi)| \leq (2^n)^{2B-3}2^{n-|C|}\delta_{B,C}[2],$$

and

$$\Pr(x^a \oplus z^a = Z \wedge x^b \oplus z^b = Z') = \frac{|\bar{\Sigma}_k^{(iii)}(P_k^\Phi)|}{|\bar{\Sigma}_k(P_k^\Phi)|} \leq \frac{\delta_{B,C}[2]}{2^{2n} - Bq}.$$

Part (iv). The approach is fairly similar to case (iii). If $B = 1$ the probability is 0 by construction. Define by $\bar{\Sigma}_k^{(iv)}(P_k^\Phi)$ the set of all elements of $\bar{\Sigma}_k(P_k^\Phi)$ that satisfy $x^a = Z$, $x^b = Z'$, and $x^a \oplus z^a \oplus x^b \oplus z^b = Z''$. In case $B > 2$, we have $|\bar{\Sigma}_k^{(iv)}(P_k^\Phi)| \leq (2^n)^{2B-4}2^{n-|C|}$. On the other hand, if $B = 2$, we have $|\bar{\Sigma}_k^{(iv)}(P_k^\Phi)| = 0$ if $\text{Bits}_C(Z'') \neq 0$ and $|\bar{\Sigma}_k^{(iv)}(P_k^\Phi)| \leq 2^n$ if $\text{Bits}_C(Z'') = 0$. In any case,

$$|\bar{\Sigma}_k^{(iv)}(P_k^\Phi)| \leq (2^n)^{2B-4}2^{n-|C|}\delta_{B,C}[2],$$

and

$$\Pr(x^a = Z \wedge x^b = Z' \wedge x^a \oplus z^a \oplus x^b \oplus z^b = Z'') = \frac{|\bar{\Sigma}_k^{(iv)}(P_k^\Phi)|}{|\bar{\Sigma}_k(P_k^\Phi)|} \leq \frac{\delta_{B,C}[2]}{2^{3n} - Bq}. \quad \square$$

4 Application to PGV Compression Functions

We consider the 12 blockcipher-based compression functions from Preneel, Govaerts, and Vandewalle (PGV) [48]. In the ICM these constructions achieve tight collision security up to about $2^{n/2}$ queries and preimage security up to about 2^n queries [9, 10, 19, 59]. The 12

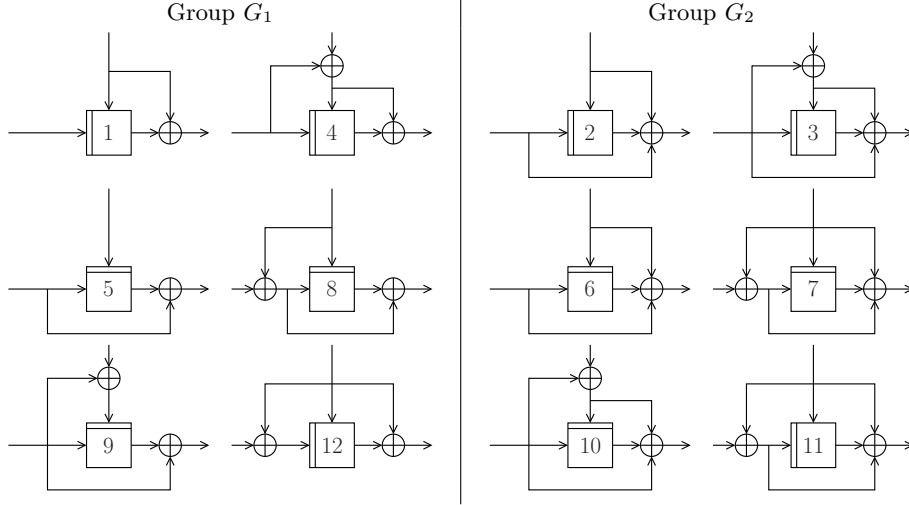


Fig. 3. The 12 PGV compression functions. When in iteration mode, the message comes in at the top. The groups G_1 and G_2 refer to Lem. 3.

constructions are depicted in Fig. 3. Here, we follow the ordering of [10], where PGV1, PGV2, and PGV5 are better known as the Matyas-Meyer-Oseas [36], Miyaguchi-Preneel, and Davies-Meyer [45] compression functions.

Baecher et al. [4] analyzed the 12 PGV constructions under ideal cipher reducibility, which at a high level covers the idea of two constructions being equally secure for the same underlying idealized blockcipher. They divide the PGV functions into two classes, in such a way that if some blockcipher makes one of the constructions secure, it makes all functions in the corresponding class secure. Applied to our WCM, the results of Baecher et al. imply the following:

Lemma 3 (Ideal Cipher Reducibility of PGV [4], informal). *Let $\pi \xleftarrow{\$} \text{BC}[\Phi](n, n)$ for some predicate Φ . Let*

$$G_1 = \{1, 4, 5, 8, 9, 12\}, \text{ and } G_2 = \{2, 3, 6, 7, 10, 11\}.$$

For any $\alpha \in \{1, 2\}$ and $i, j \in G_\alpha$, PGV i and PGV j achieve the same level of collision and preimage security once instantiated with π .

Baecher et al. also derive a reduction between the two classes, but this reduction requires a non-direct transformation on the ideal cipher π ,¹ making it unsuitable for our purposes. Thanks to Lem. 3, it suffices to only analyze PGV1 and PGV2 in the WCM: the bounds carry over to the other 10 PGV constructions. In Sect. 4.1 we analyze the collision security of these functions in the WCM. The preimage security is considered in Sect. 4.2.

4.1 Collision Security

Theorem 1. *Let $n \in \mathbb{N}$. Let $\alpha \in \{1, 2\}$ and consider PGV α . Suppose $\pi \xleftarrow{\$} \text{BC}[\Phi(A, B, \varphi^C)](n, n)$. Then, for $q \leq 2^{n-1}/B$,*

$$\text{Adv}_{\text{PGV}\alpha}^{\text{col}}(q) \leq \frac{B^2 \delta_{B,C}[1]q^2}{2^n} + \binom{B}{2} \frac{2\delta_{B,C}[2]q}{2^n} + \frac{4B^2q^2}{2^n}.$$

¹ If π makes the PGV constructions from group G_1 secure, there is a transformation τ such that τ^π makes the constructions from G_2 secure, and vice versa.

Proof. We focus on PGV2. The analysis for PGV1 is a simplification due to the absence of the feed-forward of the key. We consider any adversary that has query access to $\pi \stackrel{\$}{\leftarrow} \text{BC}[\Phi(A, B, \varphi^C)](n, n)$ and makes q queries. As a first step, we move from π to $\bar{\pi} \stackrel{\$}{\leftarrow} \overline{\text{BC}}[\Phi(A, B, \varphi^C)](n, n)$. By Lem. 2, this costs us an additional term $\frac{B^2 q(q+1)}{2^n - Bq}$.

A collision for PGV2 would imply the existence of two distinct query pairs (k, x, z) , (k', x', z') such that $k \oplus x \oplus z = k' \oplus x' \oplus z'$. We consider the i^{th} query ($i \in \{1, \dots, q\}$) to be the first query to make this condition satisfied, and sum over $i = 1, \dots, q$ at the end. For regular (forward or inverse) queries, the analysis of [9, 10, 59] mostly carries over. The analysis of predicate queries is a bit more technical.

Query $\bar{\pi}_k(x)$ or $\bar{\pi}_k^{-1}(z)$. The cases are the same by symmetry, and we consider $\bar{\pi}_k(x)$ only. Denote the response by z . There are at most $B(i-1)$ possible (k', x', z') . As z is randomly drawn from a set of size at least $2^n - Bq$, it satisfies $z = k \oplus x \oplus k' \oplus x' \oplus z'$ with probability at most $\frac{B(i-1)}{2^n - Bq}$.

Query $\bar{\pi}_k^\Phi(y)$. Denote the query response by $\{(k, x^1, z^1), \dots, (k, x^B, z^B)\}$. In case the B -set contributes only to (k, x, z) , the same reasoning as for regular queries applies with the difference that any query of the B -set may be successful and that the bound of Lem. 2 part (ii) applies: $\frac{B^2 \delta_{B,C}[1](i-1)}{2^n - Bq}$.

Now, consider the case the predicate query contributes to both (k, x, z) and (k, x', z') . There are $\binom{B}{2}$ ways for the predicate query to contribute (or 0 if $B = 1$). By Lem. 2 part (iii), which considers the success probability for any such combination, the predicate query results in a collision with probability at most $\binom{B}{2} \frac{\delta_{B,C}[2]2^n}{2^{2n} - Bq}$.

Conclusion. Taking the maximum of all success probabilities, the i^{th} query is successful with probability at most $\frac{B^2 \delta_{B,C}[1](i-1)}{2^n - Bq} + \binom{B}{2} \frac{\delta_{B,C}[2]2^n}{2^{2n} - Bq}$. Summation over $i = 1, \dots, q$ gives

$$\mathbf{Adv}_{\text{PGV2}}^{\text{col}}(q) \leq \frac{B^2 \delta_{B,C}[1]q^2}{2(2^n - Bq)} + \binom{B}{2} \frac{\delta_{B,C}[2]q}{2^n - Bq} + \frac{B^2 q(q+1)}{2^n - Bq},$$

where the last part of the bound comes from the transition from WCM to AWCM. The proof is completed by using the fact that $2^n - Bq \geq 2^{n-1}$ for $Bq \leq 2^{n-1}$, and that $q+1 \leq 2q$ for $q \geq 1$. \square

We note that the bound gets worse for increasing values of B . This has a technical cause: predicate queries are counted equally expensive as regular queries, but result in up to B new query tuples. This leads to several factors of B in the bound. As this work is mainly concerned with differential known-key attacks for which B is regularly small, these factors are of no major influence.

The implications of the bound of Thm. 1 become more visible when considering particular choices of B and C .

- (i) If $B = 1$, then $\mathbf{Adv}_{\text{PGV}\alpha}^{\text{col}}(q) \leq \frac{2^{|C|}q^2}{2^n} + \frac{4q^2}{2^n}$;
- (ii) If $B = 2$, then $\mathbf{Adv}_{\text{PGV}\alpha}^{\text{col}}(q) \leq \frac{20q^2}{2^n} + \frac{4 \cdot 2^{|C|}q}{2^n}$;
- (iii) If $B \geq 3$ (independent of n), then $\mathbf{Adv}_{\text{PGV}\alpha}^{\text{col}}(q) \leq \frac{5B^2q^2}{2^n} + \frac{B^2q}{2^n}$.

In other words, for $B = 2$ and C with $|C| \leq n/2$, or for $B \geq 3$ constant and C arbitrary, the PGV functions achieve the same $2^{n/2}$ collision security level as in the ICM. On the other hand, if $B = 1$, collisions can be found in about $2^{(n-|C|)/2}$ queries, and if $B = 2$ with $|C| > n/2$, in about $2^{n-|C|} < 2^{n/2}$ queries. See also Table 1.

Tightness

For the cases $B = 1$ and C arbitrary, and $B = 2$ and C arbitrary such that $|C| > n/2$, we derive generic attacks that demonstrate tightness of the bound of Thm. 1. Knudsen and Rijmen [27] and Sasaki et al. [53, 56] already considered how to exploit a known-key pair for

the underlying blockcipher to find a collision for the Matyas-Meyer-Oseas (PGV1) and/or Miyaguchi-Preneel (PGV2) compression functions. Their attacks correspond to our $B = 2$ case.

Proposition 1 ($B = 1$). *Let $n \in \mathbb{N}$. Let $\alpha \in \{1, 2\}$ and consider $\text{PGV}\alpha$. Suppose $\pi \xleftarrow{\$} \text{BC}[\Phi(A, 1, \varphi^C)](n, n)$. Then, $\text{Adv}_{\text{PGV}\alpha}^{\text{col}}(q) \geq \frac{q^2}{2^{n-|C|}}$.*

Proof. We construct a collision-finding adversary \mathcal{A} for PGV2. It fixes key $k = 0$, and makes predicate queries to π_k^Φ on input of distinct values y to obtain q queries (k, x_y, z_y) satisfying $\text{Bits}_C(x_y \oplus z_y) = 0$. Any two such queries collide on the entire state, $k \oplus x_y \oplus z_y = k \oplus x_{y'} \oplus z_{y'}$, with probability at least $\frac{q^2}{2^{n-|C|}}$. The attack for PGV1 is the same as we have taken $k = 0$. \square

Proposition 2 ($B = 2$ and $|C| > n/2$). *Let $n \in \mathbb{N}$. Let $\alpha \in \{1, 2\}$ and consider $\text{PGV}\alpha$. Suppose $\pi \xleftarrow{\$} \text{BC}[\Phi(A, 2, \varphi^C)](n, n)$. Then, $\text{Adv}_{\text{PGV}\alpha}^{\text{col}}(q) \geq \frac{q}{2^{n-|C|}}$.*

Proof. We construct a collision-finding adversary \mathcal{A} for PGV2. It fixes key $k = 0$, and makes predicate queries to π_k^Φ on input of distinct values y to obtain q 2-sets $\{(k, x_y^1, z_y^1), (k, x_y^2, z_y^2)\}$ satisfying $\text{Bits}_C(x_y^1 \oplus z_y^1) = \text{Bits}_C(x_y^2 \oplus z_y^2)$. These two queries collide on the entire state, $k \oplus x_y^1 \oplus z_y^1 = k \oplus x_y^2 \oplus z_y^2$, with probability at least $\frac{1}{2^{n-|C|}}$. If the adversary makes q predicate queries, we directly obtain our bound. The attack for PGV1 is the same as we have taken $k = 0$. \square

4.2 Preimage Security

Theorem 2. *Let $n \in \mathbb{N}$. Let $\alpha \in \{1, 2\}$ and consider $\text{PGV}\alpha$. Suppose $\pi \xleftarrow{\$} \text{BC}[\Phi(A, B, \varphi^C)](n, n)$. Then, for $q \leq 2^{n-2}/B$,*

$$\text{Adv}_{\text{PGV}\alpha}^{\text{epre}}(q) \leq \left(\frac{2Bq}{2^n}\right)^B + \frac{2B^2\delta_{B,C}[1]q}{2^n}.$$

Due to space limitations, the proof is given in App. B. It is much more involved than the one of Thm. 1, particularly as we cannot make use of abortable ciphers. Entering various choices of B and C shows that in the PGV functions remain mostly unaffected in the WCM if $B \geq 2$, and the same security level as in the ICM is achieved [9, 10, 59]. A slight security degradation appears for $B = 1$ as preimages can be found in about $2^{n-|C|}$.

Tightness

For the case $B = 1$, we derive a generic attack that demonstrates the tightness of the bound of Thm. 2.

Proposition 3 ($B = 1$). *Let $n \in \mathbb{N}$. Let $\alpha \in \{1, 2\}$ and consider $\text{PGV}\alpha$. Suppose $\pi \xleftarrow{\$} \text{BC}[\Phi(A, 1, \varphi^C)](n, n)$. Then, $\text{Adv}_{\text{PGV}\alpha}^{\text{epre}}(q) \geq \frac{q}{2^{n-|C|}}$.*

Proof. Let Z be any given range value with $\text{Bits}_C(Z) = 0$ (note that epre guarantees security for every range point). A preimage-finding adversary \mathcal{A} for PGV2 proceeds as follows. It fixes key $k = 0$, and makes predicate queries to π_k^Φ on input of distinct values y to obtain q queries (k, x_y, z_y) satisfying $\text{Bits}_C(x_y \oplus z_y) = 0$. Any such query hits Z on the entire state, $k \oplus x_y \oplus z_y = Z$, with probability at least $\frac{q}{2^{n-|C|}}$. The attack for PGV1 is the same as we have taken $k = 0$. \square

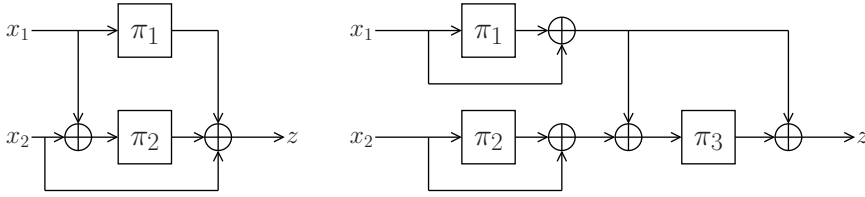


Fig. 4. Grøstl compression function (left) and Shrimpton-Stam (right).

5 Application to Grøstl Compression Function

We consider the provable security of the compression function mode of operation of Grøstl [21] (see also Fig. 4):

$$F_{\text{Grøstl}}(x_1, x_2) = x_2 \oplus \pi_1(x_1) \oplus \pi_2(x_1 \oplus x_2). \quad (6)$$

The Grøstl compression function is in fact designed to operate in a wide-pipe mode, and in the IPM, the function is proven collision secure up to about $2^{n/4}$ queries and preimage secure up to $2^{n/2}$ queries [20]. We consider the security of $F_{\text{Grøstl}}$ in the WCM, where $(\pi_1, \pi_2) \xleftarrow{\$} \text{BC}[\Phi(A, B, \varphi^C)](n)^2$. We remark that in this section we consider keyless primitives, hence $\kappa = 0$ and the k -input is dropped throughout. We furthermore note that finding collisions and preimages for $F_{\text{Grøstl}}$ is equivalent to finding them for

$$F'_{\text{Grøstl}}(x_1, x_2) = x_1 \oplus x_2 \oplus \pi_1(x_1) \oplus \pi_2(x_2), \quad (7)$$

as $F_{\text{Grøstl}}(x_1, x_2) = F'_{\text{Grøstl}}(x_1, x_1 \oplus x_2)$, and we will consider $F'_{\text{Grøstl}}$ throughout.

5.1 Collision Security

Theorem 3. *Let $n \in \mathbb{N}$. Suppose $(\pi_1, \pi_2) \xleftarrow{\$} \text{BC}[\Phi(A, B, \varphi^C)](n)^2$. Then, for $q \leq 2^{n-1}/B$,*

$$\text{Adv}_{F'_{\text{Grøstl}}}^{\text{col}}(q) \leq \frac{B^4 \delta_{B,C}[1]q^4}{2^n} + \binom{B}{2} \frac{2\delta_{B,C}[2](q^2 + 2^{n/2-|C|}q)}{2^n} + \frac{B^2 q^2}{2 \cdot 2^{n/2}} + \frac{4B^2 q^2}{2^n}.$$

The proof is given in App. C. If we enter particular choices of B and C into the bound, we find results comparable to the case of Sect. 4.1. In more detail, for $B = 2$ and C with $|C| \leq n/2$, or for $B \geq 3$ constant and C arbitrary, $F_{\text{Grøstl}}$ achieves the same $2^{n/4}$ collision security level as in the ICM [20]. If $B = 1$, the bound guarantees security up to about $2^{(n-|C|)/4}$, and if $B = 2$ with $|C| > n/2$, collisions can be found in about $2^{(n-|C|)/2}$ queries. See also Table 1. In App. D we show that the bound is optimal, by presenting tight attacks on $F'_{\text{Grøstl}}$ in the WCM.

5.2 Preimage Security

Theorem 4. *Let $n \in \mathbb{N}$. Suppose $(\pi_1, \pi_2) \xleftarrow{\$} \text{BC}[\Phi(A, B, \varphi^C)](n)^2$. Then, for $q \leq 2^{n-1}/B$,*

$$\text{Adv}_{F'_{\text{Grøstl}}}^{\text{epre}}(q) \leq \frac{2B^2 \delta_{B,C}[1](q^2 + 2^{n/2-|C|}q)}{2^n} + \frac{Bq}{2^{n/2}} + \frac{4B^2 q^2}{2^n}.$$

The proof is given in App. E. As before, we find that $F_{\text{Grøstl}}$ remains unaffected in the WCM for most cases, the sole exception being $B = 1$ for which preimages can be found in about $2^{(n-|C|)/2}$. In App. F we show that the bound is optimal, by presenting a tight attack on $F'_{\text{Grøstl}}$ for $B = 1$ in the WCM.

6 Application to Shrimpton-Stam Compression Function

In this section, we consider the provable security of the Shrimpton-Stam compression function [57] (see also Fig. 4):

$$F_{SS}(x_1, x_2) = x_1 \oplus \pi_1(x_1) \oplus \pi_3(x_1 \oplus \pi_1(x_1) \oplus x_2 \oplus \pi_2(x_2)). \quad (8)$$

This function is proven asymptotically optimally collision and preimage secure up to $2^{n/2}$ queries in the IPM [41, 51, 57]. We consider the security of F_{SS} in the WCM, where $(\pi_1, \pi_2, \pi_3) \stackrel{\$}{\leftarrow} \text{BC}[\Phi(A, B, \varphi^C)](n)^3$. (As in Sect. 5 we consider keyless functions, hence $\kappa = 0$ and the key inputs are dropped throughout.) Our findings readily apply to the generalization of F_{SS} of [41]. The analysis of this construction is significantly more complex than the ones of Sect. 4 and Sect. 5.

6.1 Collision Security

Theorem 5. *Let $n \in \mathbb{N}$. Suppose $(\pi_1, \pi_2, \pi_3) \stackrel{\$}{\leftarrow} \text{BC}[\Phi(A, B, \varphi^C)](n)^3$. Then,*

- (i) *If $B = 1$ and C arbitrary, $\text{Adv}_{F_{SS}}^{\text{col}}(2^{(n-|C|)/2-n\epsilon}) \rightarrow 0$ for $n \rightarrow \infty$;*
- (ii) *If $B = 2$ and C with $|C| \leq n/2$, $\text{Adv}_{F_{SS}}^{\text{col}}(2^{n/2-n\epsilon}) \rightarrow 0$ for $n \rightarrow \infty$;*
- (iii) *If $B = 2$ and C with $|C| > n/2$, $\text{Adv}_{F_{SS}}^{\text{col}}(2^{n-|C|-n\epsilon}) \rightarrow 0$ for $n \rightarrow \infty$;*
- (iv) *If $B \geq 3$ (independent of n) and C arbitrary, $\text{Adv}_{F_{SS}}^{\text{col}}(2^{n/2-n\epsilon}) \rightarrow 0$ for $n \rightarrow \infty$.*

Due to the technicality of the proof, the results are expressed in asymptotic terms. The proof is given in App. G. For $B = 2$ and C with $|C| \leq n/2$, or for $B \geq 3$ constant and C arbitrary, F_{SS} achieves the same security level as in the IPM. On the other hand, if $B = 1$, or if $B = 2$ but $|C| > n/2$, Thm. 5 results in a worse bound. See also Table 1. In App. H we show that the bound is optimal, by presenting tight attacks on F_{SS} in the WCM.

6.2 Preimage Security

Theorem 6. *Let $n \in \mathbb{N}$. Suppose $(\pi_1, \pi_2, \pi_3) \stackrel{\$}{\leftarrow} \text{BC}[\Phi(A, B, \varphi^C)](n)^3$. Then,*

- (i) *If $B = 1$ and C with $|C| \leq n/2$, $\text{Adv}_{F_{SS}}^{\text{epre}}(2^{n/2-n\epsilon}) \rightarrow 0$ for $n \rightarrow \infty$;*
- (ii) *If $B = 1$ and C with $|C| > n/2$, $\text{Adv}_{F_{SS}}^{\text{epre}}(2^{n-|C|-n\epsilon}) \rightarrow 0$ for $n \rightarrow \infty$;*
- (iii) *If $B \geq 2$ (independent of n) and C arbitrary, $\text{Adv}_{F_{SS}}^{\text{epre}}(2^{n/2-n\epsilon}) \rightarrow 0$ for $n \rightarrow \infty$.*

As for collision resistance, the results are expressed in asymptotic terms. The proof is given in App. I. The bounds match the ones in the IPM, except for the case of $B = 1$ and $|C| > n/2$. We leave it as an open problem to prove tightness of Thm. 6 part (ii).

7 Conclusions

Since their formal introduction by Knudsen and Rijmen at ASIACRYPT 2007 [27], numerous known-key attacks on blockciphers have appeared in literature. These attacks are often considered delicate, as it is not always clear to what extent they influence the security of cryptographic functions based on these known-key blockciphers. We presented the weak cipher model in order to investigate this impact. For a specific instance of this model, considering the existence of A sets of B queries that satisfy condition φ^C of (3), we proved that the PGV compression functions [48], the Grøstl compression function [21], and the Shrimpton-Stam compression function [57] remain mostly unaffected by the generalized weakness. Additionally, preimage security of the functions turned out to be significantly less susceptible to these types of weaknesses than collision security. The results can be readily generalized to other primitive-based functions, such as the double block length compression

functions Tandem-DM, Abreast-DM, and Hirose’s compression functions [23,30], and to the permutation-based sponge mode [5].

Our model is general enough to cover practically all differential known-key attacks in literature, such as latest results based on the rebound attack [12,22,28,38,52,53,56] and on the boomerang attack [2,7,31,54,61]. To our knowledge, our work provides the first attempt to formally analyze the effect of a wide class of cryptanalytic attacks from a modular and provable security point of view. It is a step in the direction of security beyond the ideal model, connecting practical attacks from cryptanalysis with ideal model provable security. There is still a long way to go: in order to make the connection between the two fields, we abstracted known-key attacks to a certain degree. It remains a highly challenging open research problem to generalize our findings to multiple or different weaknesses, and to different permutation-based cryptographic functions. These generalizations include the analysis of known-key based constructions for more advanced conditions φ (such as arbitrary polynomials).

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A Knudsen-Rijmen Attack on Feistel₇

We briefly discuss the attack of Knudsen and Rijmen [27] on the classical Feistel network on n bits with 7 rounds. Before doing so, we first introduce Feistel.

The Feistel network is a very common blockcipher design strategy, dating back to the design of Lucifer [58], and many generalizations of this design have appeared in literature. We use the notation of [27]. The n -bit Feistel network with 7 rounds, called “Feistel₇,” uses 2 lines of $n/2$ bits, and consists of 7 evaluations of a fixed $n/2$ -bit permutation p . Each evaluation of p is preceded by an XOR of a round key k_1, \dots, k_7 , derived from the master key using some key schedule. It is depicted in Fig. 5.

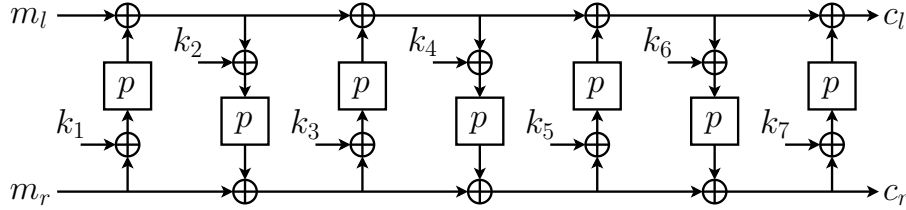


Fig. 5. The n -bit Feistel network with 7 rounds.

Now, we describe the attack by Knudsen and Rijmen on this construction [1, 27]. Assume $k_2 \neq k_6$. First, the adversary fixes an arbitrary value $y \in \{0, 1\}^{n/2}$. Then, from y the adversary derives two plaintext/ciphertext-pairs $((m_l, m_r), (c_l, c_r))$ and $((m'_l, m'_r), (c'_l, c'_r))$ as follows: y is fixed to be the input to the third permutation call p for the first tuple, while for the second tuple that input is $y \oplus \alpha$ for some non-zero value α (determined later). Then, the adversary uses p to compute the tuples in a straightforward way:

$$\begin{aligned}
 (m_l, m_r) &= (z \oplus k_4 \oplus p(y) \oplus p(m_r \oplus k_1), y \oplus k_3 \oplus p(z \oplus k_2 \oplus k_4 \oplus p(y))) \\
 (c_l, c_r) &= (z \oplus k_2 \oplus k_4 \oplus k_6 \oplus p(y) \oplus p(c_r \oplus k_7), \\
 &\quad y \oplus \alpha \oplus k_5 \oplus p(z \oplus k_2 \oplus k_4 \oplus p(y))) \\
 (m'_l, m'_r) &= (z \oplus k_2 \oplus k_4 \oplus k_6 \oplus p(y) \oplus p(m'_r \oplus k_1), \\
 &\quad y \oplus \alpha \oplus k_3 \oplus p(z \oplus k_4 \oplus k_6 \oplus p(y))) \\
 (c'_l, c'_r) &= (z \oplus k_4 \oplus p(y) \oplus p(c'_r \oplus k_7), y \oplus k_5 \oplus p(z \oplus k_4 \oplus k_6 \oplus p(y))),
 \end{aligned} \tag{9}$$

where $\alpha = y \oplus p^{-1}(k_2 \oplus k_6 \oplus p(y))$ and $z = p^{-1}(k_3 \oplus k_5 \oplus \alpha)$. These pairs satisfy $m_r \oplus c_r = m'_r \oplus c'_r$ with probability 1, but this equation is satisfied by an ideal cipher with probability at most $1/2^{n/2}$. This completes the distinguishing attack. An extension of this attack to generalized balanced Feistel networks on r wires and $4r - 1$ rounds was derived in [1].

B Proof of Theorem 2

We focus on PGV2. The analysis for PGV1 is a simplification due to the absence of the feed-forward of the key. We consider any adversary that has query access to $\pi \stackrel{s}{\leftarrow} \text{BC}[\Phi(A, B, \varphi^C)](n, n)$ and makes q queries. Let $Z \in \{0, 1\}^n$. A preimage for Z would imply the existence of a query (k, x, z) such that $x \oplus z = k \oplus Z$. We consider the i^{th} query ($i \in \{1, \dots, q\}$) to be the first query to make this condition satisfied, and sum over $i = 1, \dots, q$ at the end. For regular (forward or inverse) queries, the analysis of [9, 10, 59] mostly carries over. The analysis of predicate queries is a more technical, particularly as we cannot make use of abortable ciphers.

Query $\pi_k(x)$ or $\pi_k^{-1}(z)$. The cases are the same by symmetry, and we consider $\pi_k(x)$ only. Denote the response by z . As z is randomly drawn from a set of size at least $2^n - Bq$, it satisfies $z = x \oplus k \oplus Z$ with probability at most $\frac{1}{2^n - Bq}$.

Query $\pi_k^\Phi(y)$. Denote the query response by $\{(k, x^1, z^1), \dots, (k, x^B, z^B)\}$. If all tuples are old, the query cannot be successful as no earlier query was successful, and so we assume it contains at least one new tuple. The response is drawn uniformly at random from the set $\Sigma_k(P_k, P_k^\Phi)$. For $\ell = 0, \dots, B$, denote by $\Sigma_k^\ell(P_k, P_k^\Phi)$ the subset of all responses that have ℓ new query tuples and $B - \ell$ old query tuples (which already appear in P_k). By construction,

$$\Sigma_k(P_k, P_k^\Phi) = \bigcup_{\ell=0}^B \Sigma_k^\ell(P_k, P_k^\Phi). \quad (10)$$

Define furthermore for $\ell = 1, \dots, B$ by $\Sigma_k^{\ell, \text{pre}}(P_k, P_k^\Phi)$ the subset of elements of $\Sigma_k^\ell(P_k, P_k^\Phi)$ for which one of the new query tuples satisfies $x \oplus z = k \oplus Z$ (recall that we have excluded the case of $\ell = 0$). The predicate query is successful with probability

$$\Pr(\pi_k^\Phi(y) \text{ sets } \text{pre}(\mathcal{Q}_i)) = \sum_{\ell=1}^B \frac{|\Sigma_k^{\ell, \text{pre}}(P_k, P_k^\Phi)|}{|\Sigma_k^\ell(P_k, P_k^\Phi)|}. \quad (11)$$

Using (10), we bound (11) as

$$\Pr(\pi_k^\Phi(y) \text{ sets } \text{pre}(\mathcal{Q}_i)) \leq \frac{|\Sigma_k^{1, \text{pre}}(P_k, P_k^\Phi)|}{|\Sigma_k^B(P_k, P_k^\Phi)|} + \sum_{\ell=2}^B \frac{|\Sigma_k^{\ell, \text{pre}}(P_k, P_k^\Phi)|}{|\Sigma_k^\ell(P_k, P_k^\Phi)|}. \quad (12)$$

The reason why $\ell = 1$ is treated differently, will become clear shortly.

We next bound all relevant sets. Here, for integers $a \geq b \geq 1$, we denote by $a^{\underline{b}} = \frac{a!}{(a-b)!}$ the falling factorial power. Starting with the numerators, for $\ell = 1$ we have

$$|\Sigma_k^{1, \text{pre}}(P_k, P_k^\Phi)| \leq B \cdot |P_k|^{\underline{B-1}} \cdot (2^n - |P_k|).$$

Indeed, we have B positions for the sole new query to appear and $|P_k|^{\underline{B-1}}$ choices for the old queries. For the new query, without loss of generality (k, x^B, z^B) , it needs to satisfy $\text{Bits}_C(x^B \oplus z^B) = \text{Bits}_C(x^1 \oplus \dots \oplus z^{B-1})$ and $x^B \oplus z^B = k \oplus Z$. We have $2^n - |P_k|$ possible choices for x^B , and any choice gives at most one possible z^B . We remark that $|\Sigma_k^{1, \text{pre}}(P_k, P_k^\Phi)|$ will probably be about a factor $2^{-|C|}$ less, as we should only count all possible solutions for the $B - 1$ old queries that satisfy $\text{Bits}_C(x^1 \oplus \dots \oplus z^{B-1}) = \text{Bits}_C(k \oplus Z)$. Deriving a tighter bound would be a cumbersome exercise, but fortunately there is no need to do so: the fraction of elements in $\Sigma_k(P_k, P_k^\Phi)$ consisting of $B - 1$ old tuples is already small enough for the case $B > 1$. This is the reason why we use a special treatment for the case of $\ell = 1$ in (12).

For $\ell \in \{2, \dots, B\}$ we have

$$|\Sigma_k^{\ell, \text{pre}}(P_k, P_k^\Phi)| \leq \binom{B}{\ell} \cdot |P_k|^{\underline{B-\ell}} \cdot (2^n - |P_k|)^\ell \cdot \ell \cdot (2^n - |P_k|)^{\underline{\ell-2}} \cdot 2^{n-|C|}.$$

Again, the first term comes from identifying at which positions the new queries appear and the second term comes from the selection of old queries. Next, we have $(2^n - |P_k|)^\ell$ choices for the x -values and ℓ positions for the “winning query” to occur. For this particular winning query, the corresponding z -value is fixed by the equation $x \oplus z = k \oplus Z$. For the remaining $\ell - 1$ z -values, there are $(2^n - |P_k|)^{\ell-2}$ possibilities to freely fix the first $\ell - 2$ of them, and the last one will be adapted to the predicate condition, and can take at most $2^{n-|C|}$ values.

Regarding the denominators, for $\ell \in \{1, \dots, B\}$ we have

$$|\Sigma_k^\ell(P_k, P_k^\Phi)| \geq \binom{B}{\ell} \cdot |P_k|^{\frac{B-\ell}{\ell}} \cdot \left(\frac{(2^n - |P_k|)^\ell \cdot (2^n - |P_k|)^{\ell-1} \cdot 2^{n-|C|}}{Bq \cdot (2^n - |P_k|)^{\ell-1} \cdot (2^n - |P_k|)^{\ell-1} \cdot 2^{n-|C|}} - \right),$$

which can be seen as follows. As before, we have $\binom{B}{\ell}$ positions for the new queries to appear and $|P_k|^{\frac{B-\ell}{\ell}}$ possible lists of old queries. Regarding the ℓ new queries, without loss of generality $(k, x^1, z^1), \dots, (k, x^\ell, z^\ell)$, these need to satisfy $\text{Bits}_C(x^1 \oplus \dots \oplus z^\ell) = \text{Bits}_C(x^{\ell+1} \oplus \dots \oplus z^B)$. We first compute the number of choices for these new queries where z^ℓ is only used to adapt to this condition *and does not need to satisfy that it is fresh*. For this case, we have precisely $(2^n - |P_k|)^\ell \cdot (2^n - |P_k|)^{\ell-1}$ choices for $x^1, \dots, z^{\ell-1}, x^\ell$, and $2^{n-|C|}$ possibilities for the adaption value z^ℓ .

Now, we subtract the cases where this adapted value happens to collide, either with an older value in $\text{rng}(P_k)$ or with any of the new $z^1, \dots, z^{\ell-1}$. Any of these choices would fix z^ℓ (in total at most $(|P_k| + \ell - 1)$ possibilities). Similarly to the analysis for $|\Sigma_k^{\ell, \text{pre}}(P_k, P_k^\Phi)|$, where now x^ℓ will be used to be adapted to the predicate condition, there are at most

$$(|P_k| + \ell - 1) \cdot (2^n - |P_k|)^{\ell-1} \cdot (2^n - |P_k|)^{\ell-1} \cdot 2^{n-|C|}$$

choices for the fresh values. As $\ell \leq B$, and additionally $|P_k| \leq B(i-1) \leq B(q-1)$ for the current query, we obtain our bound for $|\Sigma_k^\ell(P_k, P_k^\Phi)|$. The bound can be simplified to

$$|\Sigma_k^\ell(P_k, P_k^\Phi)| \geq \binom{B}{\ell} \cdot |P_k|^{\frac{B-\ell}{\ell}} \cdot (2^n - |P_k|)^{\ell-1} \cdot (2^n - |P_k|)^{\ell-1} \cdot 2^{n-|C|} \cdot (2^n - 2Bq),$$

using that $\frac{(2^n - |P_k|)^\ell}{(2^n - |P_k|)^{\ell-1}} = 2^n - |P_k| - (\ell - 1) \geq 2^n - Bq$.

Plugging these bounds into (12), we find for the case $B = 1$:

$$\Pr(\pi_k^\Phi(y) \text{ sets pre}(\mathcal{Q}_i)) \leq \frac{2^n - |P_k|}{2^{n-|C|} \cdot (2^n - 2q)} \leq \frac{2^{|C|}}{2^n - 2q}.$$

For the case $B > 1$ the computation is a bit more elaborate:

$$\Pr(\pi_k^\Phi(y) \text{ sets pre}(\mathcal{Q}_i)) \leq \frac{B \cdot (2^n - |P_k|)}{(2^n - |P_k|)^{\frac{B-1}{\ell}} \cdot 2^{n-|C|} \cdot (2^n - 2Bq)} \cdot \frac{|P_k|^{\frac{B-1}{\ell}}}{(2^n - |P_k|)^{\frac{B-1}{\ell}}} + \sum_{\ell=2}^B \frac{(2^n - |P_k|)^\ell \cdot (2^n - |P_k|)^{\ell-2}}{(2^n - |P_k|)^{\ell-1} \cdot (2^n - |P_k|)^{\ell-1}} \cdot \frac{\ell}{2^n - 2Bq}.$$

For the first fraction we use that $2^n - |P_k| \leq (2^n - |P_k|)^{\frac{B-1}{\ell}}$ as $B > 1$, and additionally that $|C| \leq n$. For the falling factorial powers of the second fraction, we use that $|P_k|^{\frac{B-1}{\ell}} \leq (Bq)^{B-1}$ and $(2^n - |P_k|)^{\frac{B-1}{\ell}} \geq (2^n - |P_k| - (B-1))^{\frac{B-1}{\ell}} \geq (2^n - 2Bq)^{B-1}$. For the fraction in the sum, we use that $\frac{(2^n - |P_k|)^\ell \cdot (2^n - |P_k|)^{\ell-2}}{(2^n - |P_k|)^{\ell-1} \cdot (2^n - |P_k|)^{\ell-1}} = \frac{2^n - |P_k| - (\ell-1)}{2^n - |P_k| - (\ell-2)} \leq 1$. We obtain:

$$\begin{aligned} \Pr(\pi_k^\Phi(y) \text{ sets pre}(\mathcal{Q}_i)) &\leq \frac{B}{2^n - 2Bq} \cdot \frac{(Bq)^{B-1}}{(2^n - 2Bq)^{B-1}} + \sum_{\ell=2}^B \frac{\ell}{2^n - 2Bq} \\ &\leq \frac{B^B q^{B-1}}{(2^n - 2Bq)^B} + \frac{B^2}{2^n - 2Bq}. \end{aligned}$$

Conclusion. Taking the maximum of all success probabilities, the i^{th} query is successful with probability at most $\frac{B^B q^{B-1}}{(2^n - 2Bq)^B} + \frac{B^2 \delta_{B,C}[1]}{2^n - 2Bq}$. Summation over $i = 1, \dots, q$ gives

$$\mathbf{Adv}_{\text{PGV2}}^{\text{epre}}(q) \leq \frac{B^B q^B}{(2^n - 2Bq)^B} + \frac{B^2 \delta_{B,C}[1]q}{2^n - 2Bq}.$$

The proof is completed by using the fact that $2^n - 2Bq \geq 2^{n-1}$ for $Bq \leq 2^{n-2}$.

C Proof of Theorem 3

We consider any adversary that has access to $(\pi_1, \pi_2) \stackrel{\$}{\leftarrow} \text{BC}[\Phi(A, B, \varphi^C)](n)^2$ and makes q queries. These queries are stored in a query history \mathcal{Q} as indexed tuples of the form (κ_i, x_i, z_i) for regular queries, where we denote $z_i = \pi_{\kappa_i}(x_i)$, and $\{(\kappa_i, x_i^1, z_i^1), \dots, (\kappa_i, x_i^B, z_i^B)\}$ for predicate queries. For $q \geq 0$, by \mathcal{Q}_q we define the query history after q queries. As a first step, we move from (π_1, π_2) to $(\bar{\pi}_1, \bar{\pi}_2) \stackrel{\$}{\leftarrow} \overline{\text{BC}}[\Phi(A, B, \varphi^C)](n)^2$. By Lem. 2, this costs us an additional term $\frac{B^2 q(q+1)}{2^n - Bq}$.

A collision for $F'_{\text{Grøstl}}$ would imply the existence of query pairs $(x_1, z_1), (x'_1, z'_1)$ for $\bar{\pi}_1$ and $(x_2, z_2), (x'_2, z'_2)$ for $\bar{\pi}_2$ such that $(x_1, x_2) \neq (x'_1, x'_2)$ and

$$x_1 \oplus z_1 \oplus x_2 \oplus z_2 = x'_1 \oplus z'_1 \oplus x'_2 \oplus z'_2. \quad (13)$$

We define this configuration by $\text{col}(\mathcal{Q}_q)$. For the analysis of $\mathbf{Pr}(\text{col}(\mathcal{Q}_q))$ for the case of $B > 1$ we introduce an auxiliary event $\text{aux}(\mathcal{Q}_q)$. Let $\tau > 0$ be any integral value. We define

$$\text{aux}(\mathcal{Q}_q) : B > 1 \wedge (|\text{set}(\mathcal{Q}_q)| > q + \tau),$$

where

$$\text{set}(\mathcal{Q}_q) = \left\{ (\kappa, x_i, z_i), (\kappa, x_j, z_j) \in \mathcal{Q}_q \mid x_i \neq x_j \wedge \text{Bits}_C(x_i \oplus z_i \oplus x_j \oplus z_j) = 0 \right\}.$$

By basic probability theory:

$$\mathbf{Adv}_{F'_{\text{Grøstl}}}^{\text{col}}(q) \leq \mathbf{Pr}(\text{col}(\mathcal{Q}_q) \wedge \neg \text{aux}(\mathcal{Q}_q)) + \mathbf{Pr}(\text{aux}(\mathcal{Q}_q)) + \frac{B^2 q(q+1)}{2^n - Bq}, \quad (14)$$

where the third part of the bound comes from the transition from WCM to AWCM.

Starting with $\mathbf{Pr}(\text{aux}(\mathcal{Q}_q))$, we make a distinction between $B = 2$ and $B \geq 3$. For $B = 2$, we make a further distinction depending on whether or not $(\kappa, x_i, z_i), (\kappa, x_j, z_j)$ come from the same predicate query. Clearly, if this is the case, they satisfy the condition $\text{Bits}_C(x_i \oplus z_i \oplus x_j \oplus z_j) = 0$ by design. As the adversary makes at most q predicate queries, $\text{set}(\mathcal{Q}_q)$ will contain at most q solutions of this kind. Regarding the case where the two queries are *not* from the same predicate query (i.e., they are from two different predicate queries or at least one of them comes from a regular query), any combination of two tuples satisfies $\text{Bits}_C(x_i \oplus z_i \oplus x_j \oplus z_j) = 0$ with probability $2^{-|C|}$. Consequently, as there are at most $\binom{Bq}{2}$ such combinations,

$$\mathbf{Ex}(|\text{set}(\mathcal{Q}_q)| - q) \leq \frac{\binom{Bq}{2}}{2^{|C|}}.$$

Markov's inequality states that

$$\mathbf{Pr}(\text{aux}(\mathcal{Q}_q)) \leq \frac{\binom{Bq}{2}}{\tau 2^{|C|}}.$$

For the case of $B \geq 3$, the same bound applies with the difference that the case “ (κ, x_i, z_i) , (κ, x_j, z_j) coming from the same predicate query” does not render a solution by design but is instead counted within part two of the analysis (it is already included in the $\binom{Bq}{2}$).

We proceed with the analysis of $\Pr(\text{col}(\mathcal{Q}_q) \wedge \neg \text{aux}(\mathcal{Q}_q))$. We consider the i^{th} query ($i \in \{1, \dots, q\}$) to be the first query to make this condition satisfied, and sum over $i = 1, \dots, q$ at the end. For regular (forward or inverse) queries, the analysis of [20] mostly carries over. The analysis of predicate queries is a bit more technical.

Regular (forward or inverse) query. The cases are the same by symmetry, and we consider $\bar{\pi}_1(x_1)$ only. Denote the query response by z_1 . There are at most $B(i-1)$ possible choices for each of (x'_1, z'_1) , (x_2, z_2) , and (x'_2, z'_2) , where (x_2, z_2) and (x'_2, z'_2) may be the same but the choice of (x'_1, z'_1) is necessarily older than the current query. As z_1 is randomly drawn from a set of size at least $2^n - Bq$, it satisfies (13) with probability at most $\frac{B^3(i-1)^3}{2^n - Bq}$.

Query $\bar{\pi}_1^\Phi(y)$ or $\bar{\pi}_2^\Phi(y)$. The cases are the same by symmetry, and we consider $\bar{\pi}_1^\Phi(y)$ only. Denote the query response by $\{(x_1^B, z_1^B), \dots, (x_1^1, z_1^1)\}$. In case the B -set contributes only to (x_1, z_1) , the same reasoning as for regular queries applies with the difference that any query of the B -set may be successful and that the bound of Lem. 2 part (ii) applies: $\frac{B^4 \delta_{B,C}[1](i-1)^3}{2^n - Bq}$.

Now, consider the case the predicate query contributes to both (x_1, z_1) and (x'_1, z'_1) . There are $\binom{B}{2}$ ways for the predicate query to contribute (or 0 if $B = 1$). By $\neg \text{aux}(\mathcal{Q}_q)$, there are at most $q + \tau$ choices for (x_2, z_2) , (x'_2, z'_2) to satisfy $\text{Bits}_C(x_2 \oplus z_2 \oplus x'_2 \oplus z'_2) = 0$. By Lem. 2 part (iii), the predicate query gives a solution to (13) with probability at most $\binom{B}{2} \frac{\delta_{B,C}[2](q+\tau)2^n}{2^{2n} - Bq}$.

Conclusion. Taking the maximum of all success probabilities, the i^{th} query is successful with probability at most $\frac{B^4 \delta_{B,C}[1](i-1)^3}{2^n - Bq} + \binom{B}{2} \frac{\delta_{B,C}[2](q+\tau)2^n}{2^{2n} - Bq}$. Summation over $i = 1, \dots, q$ gives, using (14),

$$\text{Adv}_{F'_{\text{Grøstl}}}^{\text{col}}(q) \leq \frac{B^4 \delta_{B,C}[1]q^4}{2(2^n - Bq)} + \binom{B}{2} \frac{\delta_{B,C}[2](q^2 + \tau q)2^n}{2^{2n} - Bq} + \frac{\binom{Bq}{2}}{\tau 2^{|C|}} + \frac{B^2 q(q+1)}{2^n - Bq}.$$

The value τ forms a threshold between the second and third fraction (the second fraction increases while the third fraction decreases for larger τ , and vice versa for smaller τ). We put $\tau = 2^{n/2 - |C|}$. The proof is completed by using the fact that $2^n - Bq \geq 2^{n-1}$ for $Bq \leq 2^{n-1}$, and that $q+1 \leq 2q$ for $q \geq 1$.

D Tightness of the Bound of Theorem 3

For the cases $B = 1$ and C arbitrary, and $B = 2$ and C arbitrary such that $|C| > n/2$, we derive generic attacks that demonstrate tightness of the bound of Thm. 3. Mendel et al. [38] already exploited a known-key pair for the underlying permutations of Grøstl to find a semi-free-start collision with a lower complexity, but we approach the problem more generically, particularly considering arbitrary but independent π_1 and π_2 .

Proposition 4 ($B = 1$). *Let $n \in \mathbb{N}$. Suppose $(\pi_1, \pi_2) \stackrel{s}{\leftarrow} \text{BC}[\Phi(A, 1, \varphi^C)](n)^2$. Then, $\text{Adv}_{F'_{\text{Grøstl}}}^{\text{col}}(2q) \geq \frac{q^4}{2^{n-|C|}}$.*

Proof. We construct a collision-finding adversary \mathcal{A} for $F'_{\text{Grøstl}}$. It makes predicate queries to π_1^Φ on input of distinct values y to obtain q queries $(1, x_{1,y}, z_{1,y})$ satisfying $\text{Bits}_C(x_{1,y} \oplus z_{1,y}) = 0$, and likewise for π_2^Φ it obtains q queries $(2, x_{2,y}, z_{2,y})$. Any two queries to π_1^Φ and two queries to π_2^Φ collide on the entire state, $x_{1,y} \oplus z_{1,y} \oplus x_{2,y} \oplus z_{2,y} = x_{1,y'} \oplus z_{1,y'} \oplus x_{2,y'} \oplus z_{2,y'}$, with probability at least $\frac{q^4}{2^{n-|C|}}$. \square

Proposition 5 ($B = 2$ and $|C| > n/2$). *Let $n \in \mathbb{N}$. Suppose $(\pi_1, \pi_2) \stackrel{s}{\leftarrow} \text{BC}[\Phi(A, 2, \varphi^C)](n)^2$. Then, $\text{Adv}_{F'_{\text{Grøstl}}}^{\text{col}}(2q) \geq \frac{q^2}{2^{n-|C|}}$.*

Proof. We construct a collision-finding adversary \mathcal{A} for $F'_{\text{Gröstl}}$. It makes predicate queries to π_1^Φ on input of distinct values y to obtain q 2-sets $\{(1, x_{1,y}^1, z_{1,y}^1), (1, x_{1,y}^2, z_{1,y}^2)\}$ satisfying $\text{Bits}_C(x_{1,y}^1 \oplus z_{1,y}^1) = \text{Bits}_C(x_{1,y}^2 \oplus z_{1,y}^2)$, and likewise for π_2^Φ it obtains q 2-sets $\{(2, x_{2,y}^1, z_{2,y}^1), (2, x_{2,y}^2, z_{2,y}^2)\}$. Any 2-set from π_1^Φ and 2-set from π_2^Φ collide on the entire state, $x_{1,y}^1 \oplus z_{1,y}^1 \oplus x_{2,y}^1 \oplus z_{2,y}^1 = x_{1,y}^2 \oplus z_{1,y}^2 \oplus x_{2,y}^2 \oplus z_{2,y}^2$, with probability at least $\frac{q^2}{2^{n-|C|}}$. \square

E Proof of Theorem 4

We consider any adversary that has access to $(\pi_1, \pi_2) \stackrel{\$}{\leftarrow} \text{BC}[\Phi(A, B, \varphi^C)](n)^2$ and makes q queries. As in App. C, these queries are stored in a query history \mathcal{Q} as indexed tuples of the form (κ_i, x_i, z_i) for regular queries and $\{(\kappa_i, x_i^1, z_i^1), \dots, (\kappa_i, x_i^B, z_i^B)\}$ for predicate queries. We use the same approach as in App. C. In particular, we make the same transition from (π_1, π_2) to $(\bar{\pi}_1, \bar{\pi}_2) \stackrel{\$}{\leftarrow} \overline{\text{BC}}[\Phi(A, B, \varphi^C)](n)^2$.

Let $Z \in \{0, 1\}^n$. A preimage for Z would imply the existence of query pairs (x_1, z_1) for $\bar{\pi}_1$ and (x_2, z_2) for $\bar{\pi}_2$ such that

$$x_1 \oplus z_1 \oplus x_2 \oplus z_2 = Z. \quad (15)$$

We define this configuration by $\text{pre}(\mathcal{Q}_q)$. For the analysis of $\Pr(\text{pre}(\mathcal{Q}_q))$ we introduce an auxiliary event $\text{aux}(\mathcal{Q}_q)$. Let $\tau > 0$ be any integral value. We define

$$\text{aux}(\mathcal{Q}_q) : |\text{set}(\mathcal{Q}_q)| > q + \tau,$$

where

$$\text{set}(\mathcal{Q}_q) = \{(\kappa, x_i, z_i) \in \mathcal{Q}_q \mid \text{Bits}_C(x_i \oplus z_i) = 0\}.$$

By basic probability theory:

$$\text{Adv}_{F'_{\text{Gröstl}}}^{\text{epre}}(q) \leq \Pr(\text{pre}(\mathcal{Q}_q) \wedge \neg \text{aux}(\mathcal{Q}_q)) + \Pr(\text{aux}(\mathcal{Q}_q)) + \frac{B^2 q(q+1)}{2^n - Bq}, \quad (16)$$

where the third part of the bound comes from the transition from WCM to AWCM.

Starting with $\Pr(\text{aux}(\mathcal{Q}_q))$, an identical analysis to the case of the collision resistance of $F'_{\text{Gröstl}}$ in App. C gives

$$\Pr(\text{aux}(\mathcal{Q}_q)) \leq \frac{Bq}{\tau 2^{|C|}}.$$

We proceed with the analysis of $\Pr(\text{pre}(\mathcal{Q}_q) \wedge \neg \text{aux}(\mathcal{Q}_q))$. We consider the i^{th} query ($i \in \{1, \dots, q\}$) to be the first query to make this condition satisfied, and sum over $i = 1, \dots, q$ at the end. For regular (forward or inverse) queries, the analysis of [20] mostly carries over. The analysis of predicate queries is a bit more technical.

Regular (forward or inverse) query. The cases are the same by symmetry, and we consider $\bar{\pi}_1(x_1)$ only. Denote the query response by z_1 . There are at most $B(i-1)$ possible choices for (x_2, z_2) . As z_1 is randomly drawn from a set of size at least $2^n - Bq$, it satisfies (15) with probability at most $\frac{B(i-1)}{2^n - Bq}$.

Query $\bar{\pi}_1^\Phi(y)$ or $\bar{\pi}_2^\Phi(y)$. The cases are the same by symmetry, and we consider $\bar{\pi}_1^\Phi(y)$ only. Denote the query response by $\{(x_1^1, z_1^1), \dots, (x_1^B, z_1^B)\}$. We make a distinction between $B = 1$ and $B > 1$. First, if $B > 1$, the same reasoning as for regular queries applies with the difference that any query of the B -set may be successful and that the bound of Lem. 2 part (ii) applies: $[B > 1] \frac{B^2 \delta_{B,C}[1](i-1)}{2^n - Bq}$. Now, consider $B = 1$. By $\neg \text{aux}(\mathcal{Q}_q)$, there are at most $q + \tau$ choices for (x_2, z_2) to satisfy $\text{Bits}_C(x_2 \oplus z_2) = 0$. By Lem. 2 part (ii), the predicate gives a solution to (15) with probability at most $[B = 1] \frac{\delta_{B,C}[1](q+\tau)}{2^n - Bq}$. Combining, the predicate query succeeds with probability at most $\frac{B^2 \delta_{B,C}[1](q+\tau)}{2^n - Bq}$.

Conclusion. Taking the maximum of all success probabilities, the i^{th} query is successful with probability at most $\frac{B^2\delta_{B,C}[1](q+\tau)}{2^n - Bq}$. Summation over $i = 1, \dots, q$ gives, using (16),

$$\text{Adv}_{F'_{\text{Gröstl}}}^{\text{epre}}(q) \leq \frac{B^2\delta_{B,C}[1](q^2 + \tau q)}{2^n - Bq} + \frac{Bq}{\tau 2^{|C|}} + \frac{B^2q(q+1)}{2^n - Bq}.$$

We put threshold $\tau = 2^{n/2-|C|}$. The proof is completed by using the fact that $2^n - Bq \geq 2^{n-1}$ for $Bq \leq 2^{n-1}$, and that $q+1 \leq 2q$ for $q \geq 1$.

F Tightness of the Bound of Theorem 4

For the case $B = 1$, we derive a generic attack that demonstrates the tightness of the bound of Thm. 4.

Proposition 6 ($B = 1$). *Let $n \in \mathbb{N}$. Suppose $(\pi_1, \pi_2) \stackrel{\$}{\leftarrow} \text{BC}[\Phi(A, 1, \varphi^C)](n)^2$. Then, $\text{Adv}_{F'_{\text{Gröstl}}}^{\text{epre}}(2q) \geq \frac{q^2}{2^{n-|C|}}$.*

Proof. Let Z be any given range value with $\text{Bits}_C(Z) = 0$ (note that epre guarantees security for every range point). A preimage-finding adversary \mathcal{A} for $F'_{\text{Gröstl}}$ proceeds as follows. It makes predicate queries to π_1^Φ on input of distinct values y to obtain q queries $(1, x_{1,y}, z_{1,y})$ satisfying $\text{Bits}_C(x_{1,y} \oplus z_{1,y}) = 0$, and likewise for π_2^Φ it obtains q queries $(2, x_{2,y}, z_{2,y})$. Any query to π_1^Φ and query to π_2^Φ collide with Z on the entire state, $x_{1,y} \oplus z_{1,y} \oplus x_{2,y} \oplus z_{2,y} = Z$, with probability at least $\frac{q^2}{2^{n-|C|}}$. \square

G Proof of Theorem 5

The bulk of the proof is captured in the following proposition.

Proposition 7. *Let $n \in \mathbb{N}$. Suppose $(\pi_1, \pi_2, \pi_3) \stackrel{\$}{\leftarrow} \text{BC}[\Phi(A, B, \varphi^C)](n)^3$. Then, for $q \leq 2^{n-1}/B$ and any positive integral value τ ,*

$$\begin{aligned} \text{Adv}_{F_{\text{SS}}}^{\text{col}}(q) &\leq \frac{2\tau^3 B^2 (B+2)\delta_{B,C}[1]q^2 + 4B^2q^2 + e^2B^2(\delta_{B,C}[1])^2 2^{-|C|}q^2 + e^2B^2q^2}{2^n} + \\ &\binom{B}{2} \left[\frac{2\tau^2\delta_{B,C}[2]q}{2^n} + \frac{2\tau B^2\delta_{B,C}[2]q^3}{2^{2n}} + \frac{2B^4\delta_{B,C}[2]q^5}{2^{3n}} \right] + \\ &2^n \left(\frac{e(B + \binom{B}{2})q}{\tau} \cdot \frac{2B^2\delta_{B,C}[1]q + 2\delta_{B,C}[2]}{2^n} \right)^{\tau/(B + \binom{B}{2})} + \\ &2^{2n} \left(\frac{2eB^3\delta_{B,C}[1]q^2}{\tau 2^n} \right)^{\tau/B} + 2^n \left(\frac{2eB^3\delta_{B,C}[1]q^2}{2^n} + \frac{eB^3\delta_{B,C}[2]q^3}{\tau 2^{2n}} \right)^{2\tau/B}. \end{aligned} \quad (17)$$

The proof follows later in the appendix, but we first prove Thm. 5 using Prop. 7.

Proof of Theorem 5

(i). We put $\tau = n$. Then, for the particular choice of B and C the bound (17) of Prop. 7 simplifies to

$$\text{Adv}_{F_{\text{SS}}}^{\text{col}}(q) \leq \frac{6n^3 2^{|C|}q^2 + 4q^2 + e^2 2^{|C|}q^2 + e^2q^2}{2^n} + 2 \cdot 2^{2n} \left(\frac{2e 2^{|C|}q^2}{2^n} \right)^n.$$

For $q = 2^{(n-|C|)/2-n\epsilon}$, it is clear that all terms approach 0 for $n \rightarrow \infty$.

(ii) and (iii). We put $\tau = n$. Then, for the particular choice of B (with C arbitrary) the bound (17) of Prop. 7 simplifies to

$$\begin{aligned} \mathbf{Adv}_{\text{FSS}}^{\text{col}}(q) &\leq \frac{32n^3q^2 + 16q^2 + 8e^2q^2}{2^n} + \frac{2n^22^{|C|}q}{2^n} + \frac{8n2^{|C|}q^3}{2^{2n}} + \frac{32 \cdot 2^{|C|}q^5}{2^{3n}} + \\ &2^n \left(\frac{24eq^2 + 6e2^{|C|}q}{n2^n} \right)^{n/3} + 2^{2n} \left(\frac{16eq^2}{n2^n} \right)^{n/2} + 2^n \left(\frac{16eq^2}{2^n} + \frac{8e2^{|C|}q^3}{n2^{2n}} \right)^n. \end{aligned}$$

For $q = 2^{n/2-n\epsilon}$ (in case $|C| \leq n/2$), and for $q = 2^{n-|C|-n\epsilon}$ (in case $|C| > n/2$), it is clear that all terms approach 0 for $n \rightarrow \infty$.

(iv). The reasoning of previous cases carries over, and we will not go into detail. Note that we have $\delta_{B,C}[1] = \delta_{B,C}[2] = 1$ for $B \geq 3$.

Proof of Proposition 7

We consider any adversary that has access to $(\pi_1, \pi_2, \pi_3) \stackrel{\$}{\leftarrow} \text{BC}[\Phi(A, B, \varphi^C)](n)^3$ and makes q queries. These queries are stored in a query history \mathcal{Q} as indexed tuples of the form (κ_i, x_i, z_i) for regular queries, where we denote $z_i = \pi_{\kappa_i}(x_i)$, and $\{(\kappa_i, x_i^1, z_i^1), \dots, (\kappa_i, x_i^B, z_i^B)\}$ for predicate queries. For $q \geq 0$, by \mathcal{Q}_q we define the query history after q queries. As a first step, we move from (π_1, π_2, π_3) to $(\bar{\pi}_1, \bar{\pi}_2, \bar{\pi}_3) \stackrel{\$}{\leftarrow} \overline{\text{BC}}[\Phi(A, B, \varphi^C)](n)^3$. By Lem. 2, this costs us an additional term $\frac{B^2q(q+1)}{2^n - Bq}$.

The adversary's goal is to find a collision for FSS , which corresponds to obtaining query pairs $(x_1, z_1), (x'_1, z'_1)$ for $\bar{\pi}_1$, $(x_2, z_2), (x'_2, z'_2)$ for $\bar{\pi}_2$, and $(x_3, z_3), (x'_3, z'_3)$ for $\bar{\pi}_3$ in the query history, such that:

$$(x_1, x_2) \neq (x'_1, x'_2), \quad (18a)$$

$$x_1 \oplus z_1 \oplus x_2 \oplus z_2 = x_3, \quad (18b)$$

$$x'_1 \oplus z'_1 \oplus x'_2 \oplus z'_2 = x'_3, \quad (18c)$$

$$x_1 \oplus z_1 \oplus z_3 = x'_1 \oplus z'_1 \oplus z'_3. \quad (18d)$$

We define this configuration by $\text{col}(\mathcal{Q}_q)$. This means:

$$\mathbf{Adv}_{\text{FSS}}^{\text{col}}(q) = \Pr(\text{col}(\mathcal{Q}_q)) + \frac{B^2q(q+1)}{2^n - Bq}, \quad (19)$$

where the second part of the bound comes from the transition from WCM to AWCM.

For the analysis of $\Pr(\text{col}(\mathcal{Q}_q))$ we introduce an auxiliary event $\text{aux}(\mathcal{Q}_q)$. Let $\tau_2, \tau_3, \tau_4, \tau_5 > 0$ be any integral values. We define $\text{aux}(\mathcal{Q}_q) = \text{aux}_1 \vee \dots \vee \text{aux}_5(\mathcal{Q}_q)$, where

$$\text{aux}_1(\mathcal{Q}_q) : \left| \left\{ \left\{ (\kappa_i, x_i^1, z_i^1), \dots, (\kappa_i, x_i^B, z_i^B) \right\} \in \mathcal{Q}_q, \right. \right. \\ \left. \left. a, b \in \{1, \dots, B\}, a \neq b \mid x_i^a \oplus z_i^a = x_i^b \oplus z_i^b \right\} \right| > 0;$$

$$\text{aux}_2(\mathcal{Q}_q) : \max_{Z \in \{0,1\}^n} \left| \left\{ (\kappa_i, x_i, z_i) \in \mathcal{Q}_q \mid x_i \oplus z_i = Z \right\} \right| > \tau_2;$$

$$\text{aux}_3(\mathcal{Q}_q) : \max_{Z \in \{0,1\}^n} \left| \left\{ (\kappa_i, x_i, z_i), (\kappa_j, x_j, z_j) \in \mathcal{Q}_q \mid \right. \right. \\ \left. \left. (\kappa_i, x_i) \neq (\kappa_j, x_j) \wedge x_i \oplus z_i \oplus x_j \oplus z_j = Z \right\} \right| > \tau_3;$$

$$\text{aux}_4(\mathcal{Q}_q) : \max_{Z, Z' \in \{0,1\}^n} \left| \left\{ (1, x_i, z_i), (2, x_k, z_k), (3, x_l, z_l) \in \mathcal{Q}_q \mid \right. \right. \\ \left. \left. x_k \oplus z_k \oplus x_l = Z' \wedge x_i \oplus z_i \oplus x_k \oplus z_k \oplus x_l \oplus z_l = Z \right\} \right| > \tau_4;$$

$$\text{aux}_5(\mathcal{Q}_q) : \max_{Z \in \{0,1\}^n} \left| \left\{ (1, x_i, z_i), (1, x_j, z_j), (2, x_k, z_k), (3, x_l, z_l) \in \mathcal{Q}_q \mid \right. \right. \\ \left. \left. x_j \oplus z_j \oplus x_k \oplus z_k \oplus x_l = 0 \wedge x_i \oplus z_i \oplus x_k \oplus z_k \oplus x_l \oplus z_l = Z \right\} \right| > \tau_5.$$

We note that $\text{aux}_1(\mathcal{Q}_q)$ only applies to predicate queries, but $\text{aux}_2(\mathcal{Q}_q), \dots, \text{aux}_5(\mathcal{Q}_q)$ apply to regular queries *and* parts of predicate queries (abusing notation). By basic probability

theory, we obtain for (19):

$$\Pr(\text{col}(\mathcal{Q}_q)) \leq \Pr(\text{col}(\mathcal{Q}_q) \wedge \neg \text{aux}(\mathcal{Q}_q)) + \Pr(\text{aux}(\mathcal{Q}_q)). \quad (20)$$

We start with the analysis of $\Pr(\text{col}(\mathcal{Q}_q) \wedge \neg \text{aux}(\mathcal{Q}_q))$. For obtaining a query history that fulfills configuration $\text{col}(\mathcal{Q}_q)$, it may be the case that a query appears at multiple positions. For instance, $x_1 = x'_1$. We split the analysis of $\text{col}(\mathcal{Q}_q)$ into essentially all different possible cases. We define for binary $\alpha_1, \alpha_2, \alpha_3$ by $\text{col}_{\alpha_1 \alpha_2 \alpha_3}(\mathcal{Q})$ the configuration $\text{col}(\mathcal{Q})$ restricted to

$$x_1 = x'_1 \equiv \alpha_1, \quad x_2 = x'_2 \equiv \alpha_2, \quad x_3 = x'_3 \equiv \alpha_3.$$

By construction, $\text{col}(\mathcal{Q}_q) \Rightarrow \bigvee_{\alpha_1, \alpha_2, \alpha_3 \in \{0,1\}} \text{col}_{\alpha_1 \alpha_2 \alpha_3}(\mathcal{Q}_q)$, and from (19-20) we obtain the following bound on $\text{Adv}_{\text{FSS}}^{\text{col}}(q)$:

$$\text{Adv}_{\text{FSS}}^{\text{col}}(q) \leq \sum_{\substack{\alpha_1, \alpha_2, \\ \alpha_3 \in \{0,1\}}} \Pr(\text{col}_{\alpha_1 \alpha_2 \alpha_3}(\mathcal{Q}_q) \wedge \neg \text{aux}(\mathcal{Q}_q)) + \Pr(\text{aux}(\mathcal{Q}_q)) + \frac{B^2 q(q+1)}{2^n - Bq}. \quad (21)$$

The probabilities $\Pr(\text{col}_{\alpha_1 \alpha_2 \alpha_3}(\mathcal{Q}_q) \wedge \neg \text{aux}(\mathcal{Q}_q))$, for all possible $\alpha_1 \alpha_2 \alpha_3$, are bounded in Lems. 4-6. For $\Pr(\text{aux}(\mathcal{Q}_q))$, note that

$$\Pr(\text{aux}(\mathcal{Q}_q)) \leq \Pr(\text{aux}_1(\mathcal{Q}_q)) + \Pr(\text{aux}_2(\mathcal{Q}_q) \mid \neg \text{aux}_1(\mathcal{Q}_q)) + \sum_{i=3}^5 \Pr(\text{aux}_i(\mathcal{Q}_q) \mid \neg \text{aux}_{2 \vee \dots \vee i-1}(\mathcal{Q}_q)). \quad (22)$$

The probabilities on $\text{aux}_1(\mathcal{Q}_q)$ and $\text{aux}_2(\mathcal{Q}_q)$ are rather straightforward and bounded in Lem. 7. The bounds on $\text{aux}_3(\mathcal{Q}_q), \dots, \text{aux}_5(\mathcal{Q}_q)$ are more involved and discussed in Lems. 8-10. In these lemmas, we regularly make use of Stirling's approximation: $t! \geq (t/e)^t$ for any t . The proof ideas of these lemmas at a high level consist of noticing that any query adds a solution to the set in question with a certain maximal probability p , and in that case it adds at most τ' solutions. This means that if τ solutions are required, at least τ/τ' queries have to contribute. Afterwards, all lemmas are combined to obtain a bound on $\text{Adv}_{\text{FSS}}^{\text{col}}(q)$.

Lemma 4. $\Pr(\text{col}_{000}(\mathcal{Q}_q) \wedge \neg \text{aux}(\mathcal{Q}_q)) \leq \frac{(\tau_3^2 + \tau_5)B^2 \delta_{B,C}[1]q^2}{2^n - Bq} + \binom{B}{2} \frac{(\tau_3^2 2^n + \tau_3 B^2 q^2) \delta_{B,C}[2]q}{2^{2n} - Bq} + \binom{B}{2} \frac{B^4 \delta_{B,C}[2]q^5}{2^{3n} - Bq}.$

Proof. In this configuration, we consider (18) with an additional requirement that $x_1 \neq x'_1$, $x_2 \neq x'_2$, and $x_3 \neq x'_3$. Note that (18) can be equivalently stated as

$$x_1 \neq x'_1, \quad x_2 \neq x'_2, \quad x_3 \neq x'_3, \quad (23a)$$

$$x_1 \oplus z_1 \oplus x_2 \oplus z_2 = x_3, \quad (23b)$$

$$x'_1 \oplus z'_1 \oplus x'_2 \oplus z'_2 = x'_3, \quad (23c)$$

$$x_2 \oplus z_2 \oplus x_3 \oplus z_3 = x'_2 \oplus z'_2 \oplus x'_3 \oplus z'_3, \quad (23d)$$

by XORing the second and third equations to the fourth one. We will use both representations interchangeably. For the analysis of $\text{col}_{000}(\mathcal{Q}_q) \wedge \neg \text{aux}(\mathcal{Q}_q)$, we say that the i^{th} query ($i \in \{1, \dots, q\}$) is successful if it makes configuration $\text{col}_{000}(\mathcal{Q}_i)$ satisfied and $\neg \text{aux}(\mathcal{Q}_i)$ holds. Now, by basic probability theory, we can analyze the probability of the i^{th} query being successful, and sum over $i = 1, \dots, q$.

Let $i \in \{1, \dots, q\}$. Clearly, if $\neg \text{aux}_i(\mathcal{Q}_i)$ holds, the i^{th} query can certainly not be successful, so we assume $\neg \text{aux}(\mathcal{Q}_i)$ and analyze the probability the i^{th} query makes $\text{col}_{000}(\mathcal{Q}_i)$ satisfied. For regular (forward or inverse) queries, the analysis of [41] mostly carries over. The difficulty lies in the analysis of predicate queries.

Query $\bar{\pi}_1(x_1)$ or $\bar{\pi}_1^{-1}(z_1)$. The cases are the same by symmetry, and we consider $\bar{\pi}_1(x_1)$ only. Denote the query response by z_1 . There are at most Bq choices for (x'_3, z'_3) . For any such choice, by $\neg\text{aux}_3(Q_q)$, there are at most τ_3 choices for $(x'_1, z'_1), (x'_2, z'_2)$ to fulfill (23c). Again by $\neg\text{aux}_3(Q_q)$, for any such choice there are at most τ_3 choices for $(x_2, z_2), (x_3, z_3)$ to fulfill (23d). As z_1 is randomly drawn from a set of size at least $2^n - Bq$, (23b) is fulfilled with probability at most $\frac{\tau_3^2 Bq}{2^n - Bq}$.

Query $\bar{\pi}_2(x_2)$ or $\bar{\pi}_2^{-1}(z_2)$. The cases are the same by symmetry, and we consider $\bar{\pi}_2(x_2)$ only. Denote the query response by z_2 . There are at most Bq choices for (x_3, z_3) . By $\neg\text{aux}_5(Q_q)$, there are at most τ_5 choices for $(x_1, z_1), (x'_1, z'_1), (x'_2, z'_2), (x'_3, z'_3)$ to satisfy $x'_1 \oplus z'_1 \oplus x'_2 \oplus z'_2 = x'_3$ and $x_1 \oplus z_1 \oplus x'_2 \oplus z'_2 \oplus x'_3 \oplus z'_3 = z_3$, hence to fulfill (18c-18d) for the particular value z_3 (where we XOR the two equations). As z_2 is randomly drawn from a set of size at least $2^n - Bq$, (18b) is fulfilled with probability at most $\frac{\tau_5 Bq}{2^n - Bq}$.

Query $\bar{\pi}_3(x_3)$. Denote the query response by z_3 . There are at most Bq choices for (x'_3, z'_3) . For any such choice, by $\neg\text{aux}_3(Q_q)$, there are at most τ_3 choices for $(x'_1, z'_1), (x'_2, z'_2)$ to fulfill (23c). By $\neg\text{aux}_3(Q_q)$, there are at most τ_3 choices for $(x_1, z_1), (x_2, z_2)$ to fulfill (23b). As z_3 is randomly drawn from a set of size at least $2^n - Bq$, (23d) is fulfilled with probability at most $\frac{\tau_3^2 Bq}{2^n - Bq}$.

Query $\bar{\pi}_3^{-1}(z_3)$. Denote the query response by x_3 . There are at most Bq choices for (x_2, z_2) . By $\neg\text{aux}_5(Q_q)$, there are at most τ_5 choices for $(x_1, z_1), (x'_1, z'_1), (x'_2, z'_2), (x'_3, z'_3)$ to satisfy $x'_1 \oplus z'_1 \oplus x'_2 \oplus z'_2 = x'_3$ and $x_1 \oplus z_1 \oplus x'_2 \oplus z'_2 \oplus x'_3 \oplus z'_3 = z_3$, hence to fulfill (18c-18d) for the particular value z_3 (where we XOR the two equations). As x_3 is randomly drawn from a set of size at least $2^n - Bq$, (18b) is fulfilled with probability at most $\frac{\tau_5 Bq}{2^n - Bq}$.

Query $\bar{\pi}_1^\Phi(y_1)$. Denote the query response by $\{(x_1^1, z_1^1), \dots, (x_1^B, z_1^B)\}$. In case the B -set contributes at one position only (say (x_1, z_1)), the same reasoning as for regular queries applies with the difference that any query of the B -set may be successful and that the bound of Lem. 2 part (ii) applies: $\frac{\tau_3^2 B^2 \delta_{B,C}[1]q}{2^n - Bq}$.

Now consider the case the predicate query contributes to both (x_1, z_1) and (x'_1, z'_1) . There are $\binom{B}{2}$ ways for the predicate query to contribute (or 0 if $B = 1$). By $\neg\text{aux}_3(Q_q)$, there are at most $\tau_3^2 2^n$ choices for $(x_2, z_2), (x'_2, z'_2), (x_3, z_3), (x'_3, z'_3)$ to fulfill (23d). By Lem. 2 part (iii), (23b-23c) are fulfilled with probability at most $\binom{B}{2} \frac{\tau_3^2 \delta_{B,C}[2]2^n}{2^{2n} - Bq}$.

Query $\bar{\pi}_2^\Phi(y_2)$. Denote the query response by $\{(x_2^a, z_2^a), (x_2^b, z_2^b)\}$. In case the B -set contributes at one position only (say (x_2, z_2)), the same reasoning as for regular queries applies with the difference that any query of the B -set may be successful and that the bound of Lem. 2 part (ii) applies: $\frac{\tau_5 B^2 \delta_{B,C}[1]q}{2^n - Bq}$.

Now consider the case the predicate query contributes to both (x_2, z_2) and (x'_2, z'_2) . There are $\binom{B}{2}$ ways for the predicate query to contribute (or 0 if $B = 1$). There are at most $(Bq)^2$ choices for $(x_3, z_3), (x'_3, z'_3)$. By $\neg\text{aux}_3(Q_q)$, there are at most τ_3 choices for $(x_1, z_1), (x'_1, z'_1)$ to fulfill (18d). By Lem. 2 part (iii), (18b-18c) are fulfilled with probability at most $\binom{B}{2} \frac{\tau_3 B^2 \delta_{B,C}[2]q^2}{2^{2n} - Bq}$.

Query $\bar{\pi}_3^\Phi(y_3)$. Denote the query response by $\{(x_3^a, z_3^a), (x_3^b, z_3^b)\}$. In case the B -set contributes at one position only (say (x_3, z_3)), the same reasoning as for regular forward queries applies with the difference that any query of the B -set may be successful and that the bound of Lem. 2 part (ii) applies: $\frac{\tau_3^2 B^2 \delta_{B,C}[1]q}{2^n - Bq}$.

Now consider the case the predicate query contributes to both (x_3, z_3) and (x'_3, z'_3) . There are $\binom{B}{2}$ ways for the predicate query to contribute (or 0 if $B = 1$). There are at most $(Bq)^4$ choices for $(x_1, z_1), (x'_1, z'_1), (x_2, z_2), (x'_2, z'_2)$. By Lem. 2 part (iv), (23b-23d) are fulfilled with probability at most $\binom{B}{2} \frac{B^4 \delta_{B,C}[2]q^4}{2^{3n} - Bq}$.

Conclusion. Taking the maximum of all success probabilities, the i^{th} query is successful with probability at most $\frac{(\tau_3^2 + \tau_5) B^2 \delta_{B,C}[1]q}{2^n - Bq} + \binom{B}{2} \frac{(\tau_3^2 2^n + \tau_3 B^2 q^2) \delta_{B,C}[2]}{2^{2n} - Bq} + \binom{B}{2} \frac{B^4 \delta_{B,C}[2]q^4}{2^{3n} - Bq}$. The claimed bound is obtained by summing over $i = 1, \dots, q$. \square

Lemma 5. $\Pr(\text{col}_{\alpha_1\alpha_2\alpha_3}(\mathcal{Q}_q) \wedge \neg\text{aux}(\mathcal{Q}_q)) = 0$ for $\alpha_1\alpha_2\alpha_3 \in \{001, 010, 011, 101\}$, provided $\tau_2 = 1$.

Proof. We start with the cases 001, 011. For these cases, (18d) reduces to finding $(x_1, z_1), (x'_1, z'_1)$ such that $x_1 \oplus z_1 = x'_1 \oplus z'_1$. By $\neg\text{aux}_2(\mathcal{Q}_q)$ with $\tau_2 = 1$, there do not exist two such tuples. Similarly, for 101, in order to comply with (18b) and (18c), it is required to find $(x_2, z_2), (x'_2, z'_2)$ such that $x_2 \oplus z_2 = x'_2 \oplus z'_2$, and the remaining reasoning is the same. Finally for 010, (18b) and (18c) can be simplified to $x_1 \oplus z_1 \oplus x_3 = x'_1 \oplus z'_1 \oplus x'_3$, which once XORed with (18d) reduces to finding $(x_3, z_3), (x'_3, z'_3)$ such that $x_3 \oplus z_3 = x'_3 \oplus z'_3$, and the remaining reasoning is the same. \square

Lemma 6. $\Pr(\text{col}_{\alpha_1\alpha_2\alpha_3}(\mathcal{Q}_q) \wedge \neg\text{aux}(\mathcal{Q}_q)) = 0$ for $\alpha_1\alpha_2\alpha_3 \in \{100, 110, 111\}$.

Proof. If $\alpha_1\alpha_2 = 11$, the configuration $\text{col}_{\alpha_1\alpha_2\alpha_3}(\mathcal{Q}_q)$ contradicts (18a), and if $\alpha_1\alpha_2\alpha_3 = 100$, it contradicts (18d). In any case, the configuration is invalid. \square

Lemma 7. $\Pr(\text{aux}_1(\mathcal{Q}_q)) + \Pr(\text{aux}_2(\mathcal{Q}_q) \mid \neg\text{aux}_1(\mathcal{Q}_q)) \leq \binom{B}{2} \frac{\delta_{B,C}[2]q2^n}{2^{2n} - Bq} + 2^{n-|C|} \left(\frac{eB\delta_{B,C}[1]q}{(\tau_2+1)(2^n - Bq)} \right)^{\tau_2+1} + 2^n \left(\frac{eBq}{(\tau_2+1)(2^n - Bq)} \right)^{\tau_2+1}$.

Proof. We bound these probabilities separately.

$\Pr(\text{aux}_1(\mathcal{Q}_q))$. This auxiliary event only applies to predicate queries where $B > 1$. Consider the i^{th} query and let it be a predicate query $\{(\kappa_i, x_i^1, z_i^1), \dots, (\kappa_i, x_i^B, z_i^B)\}$. Let $a, b \in \{1, \dots, B\}$ distinct. By Lem. 2 part (iii), we find that $x_i^a \oplus z_i^a = x_i^b \oplus z_i^b$ with probability at most $\frac{\delta_{B,C}[2]2^n}{2^{2n} - Bq}$. Summing over all queries and all choices for a, b , we obtain

$$\Pr(\text{aux}_1(\mathcal{Q}_q)) \leq \binom{B}{2} \frac{\delta_{B,C}[2]q2^n}{2^{2n} - Bq}. \quad (24)$$

$\Pr(\text{aux}_2(\mathcal{Q}_q) \mid \neg\text{aux}_1(\mathcal{Q}_q))$. Let $Z \in \{0, 1\}^n$. Consider the i^{th} query. If this is a regular (forward or inverse) query (κ_i, x_i, z_i) , it satisfies $x_i \oplus z_i = Z$ with probability at most $\frac{1}{2^n - Bq}$. On the other hand, if it is a predicate query $\{(\kappa_i, x_i^1, z_i^1), \dots, (\kappa_i, x_i^B, z_i^B)\}$, one of the B satisfies $x_i \oplus z_i = Z$ with probability at most $\frac{B\delta_{B,C}[1]}{2^n - Bq}$ if $\text{Bits}_C(Z) = 0$ (Lem. 2 part (ii)) and with probability at most $\frac{B}{2^n - Bq}$ if $\text{Bits}_C(Z) \neq 0$ (in this case the equation could never be satisfied if $B = 1$, and $\delta_{B,C}[1] = 1$ for $B > 1$). By $\neg\text{aux}_1(\mathcal{Q}_q)$, at most one of the B queries of the tuple equals Z . If $\text{Bits}_C(Z) = 0$, more than τ_2 queries result in a solution with probability at most $\binom{q}{\tau_2+1} \left(\frac{B\delta_{B,C}[1]}{2^n - Bq} \right)^{\tau_2+1} \leq \left(\frac{eB\delta_{B,C}[1]q}{(\tau_2+1)(2^n - Bq)} \right)^{\tau_2+1}$. If $\text{Bits}_C(Z) \neq 0$, we similarly obtain bound $\left(\frac{eBq}{(\tau_2+1)(2^n - Bq)} \right)^{\tau_2+1}$. Considering any possible choice for Z , we obtain:

$$\Pr(\text{aux}_2(\mathcal{Q}_q) \mid \neg\text{aux}_1(\mathcal{Q}_q)) \leq 2^{n-|C|} \left(\frac{eB\delta_{B,C}[1]q}{(\tau_2+1)(2^n - Bq)} \right)^{\tau_2+1} + (2^n - 2^{n-|C|}) \left(\frac{eBq}{(\tau_2+1)(2^n - Bq)} \right)^{\tau_2+1}. \quad (25)$$

The claim is obtained by adding (24) and a simplified version of (25). \square

The artificial split in the analysis of $\Pr(\text{aux}_2(\mathcal{Q}_q) \mid \neg\text{aux}_1(\mathcal{Q}_q))$ is required as a special treatment is needed for $B = 1$.

Lemma 8. $\Pr(\text{aux}_3(\mathcal{Q}_q) \mid \neg\text{aux}_2(\mathcal{Q}_q)) \leq 2^n \left(\frac{e(\tau_2 B + \binom{B}{2})q}{\tau_3 + 1} \left(\frac{B^2\delta_{B,C}[1]q}{2^n - Bq} + \frac{\delta_{B,C}[2]2^n}{2^{2n} - Bq} \right) \right)^{\frac{\tau_3+1}{\tau_2 B + \binom{B}{2}}}$.

Proof. Let $Z \in \{0, 1\}^n$. Consider the i^{th} query. A special treatment is needed for predicate queries as two queries from one single B -set may make the entire equation satisfied.

Regular (forward or inverse) query. Denote the query by (κ_i, x_i, z_i) . For any other tuple (κ_j, x_j, z_j) (at most Bq choices), the equation is satisfied with probability at most $\frac{1}{2^n - Bq}$. Hence, the query is successful with probability at most $\frac{Bq}{2^n - Bq}$. By $\neg\text{aux}_2(\mathcal{Q}_q)$, each hit adds at most τ_2 solutions.

Predicate query. Denote the query response by $\{(\kappa_i, x_i^1, z_i^1), \dots, (\kappa_i, x_i^B, z_i^B)\}$. In case the B -set contributes at one position only (say (κ_i, x_i, z_i)), the same reasoning as for regular queries applies with the difference that any query of the B -set may be successful and that the bound of Lem. 2 part (ii) applies: $\frac{B^2 \delta_{B,C}[1]q}{2^n - Bq}$. Again, by $\neg\text{aux}_2(\mathcal{Q}_q)$, each hit adds at most $\tau_2 B$ solutions.

Now consider the case the predicate query contributes to both (κ_i, x_i, z_i) and (κ_j, x_j, z_j) . By Lem. 2 part (iii), we find that the equation is satisfied with probability at most $\frac{\delta_{B,C}[2]2^n}{2^{2n} - Bq}$, and it adds at most $\binom{B}{2}$ solutions.

Conclusion. Thus, any query results in a solution with probability at most $\frac{B^2 \delta_{B,C}[1]q}{2^n - Bq} + \frac{\delta_{B,C}[2]2^n}{2^{2n} - Bq}$, and each hit adds at most $\tau_2 B + \binom{B}{2}$ solutions. Consequently, more than τ_3 solutions are obtained with probability at most

$$\begin{aligned} & \binom{q}{\frac{\tau_3+1}{\tau_2 B + \binom{B}{2}}} \left(\frac{B^2 \delta_{B,C}[1]q}{2^n - Bq} + \frac{\delta_{B,C}[2]2^n}{2^{2n} - Bq} \right)^{\frac{\tau_3+1}{\tau_2 B + \binom{B}{2}}} \\ & \leq \left(\frac{e(\tau_2 B + \binom{B}{2})q}{\tau_3 + 1} \left(\frac{B^2 \delta_{B,C}[1]q}{2^n - Bq} + \frac{\delta_{B,C}[2]2^n}{2^{2n} - Bq} \right) \right)^{\frac{\tau_3+1}{\tau_2 B + \binom{B}{2}}}. \end{aligned}$$

Considering any possible choice for Z , we obtain the claimed bound. \square

Lemma 9. $\Pr(\text{aux}_4(\mathcal{Q}_q) \mid \neg\text{aux}_{2 \vee 3}(\mathcal{Q}_q)) \leq 2^{2n} \left(\frac{e\tau_2\tau_3 B^3 \delta_{B,C}[1]q^2}{(\tau_4 + 1)(2^n - Bq)} \right)^{\frac{\tau_4+1}{\tau_3 B}}$.

Proof. Let $Z, Z' \in \{0, 1\}^n$. The goal is to find tuples $(1, x_i, z_i), (2, x_k, z_k), (3, x_l, z_l) \in \mathcal{Q}_q$ such that

$$x_k \oplus z_k \oplus x_l = Z', \tag{26a}$$

$$x_i \oplus z_i \oplus x_k \oplus z_k \oplus x_l \oplus z_l = Z. \tag{26b}$$

Consider the i^{th} query.

Regular (forward or inverse) query to $\bar{\pi}_1$. Without loss of generality, consider a forward query. Denote the query by $(1, x_i, z_i)$. There are at most Bq choices for $(3, x_l, z_l)$. For any such choice, by $\neg\text{aux}_2(\mathcal{Q}_q)$, there are at most τ_2 choices for $(2, x_k, z_k)$ to fulfill (26a). As z_i is randomly drawn from a set of size at least $2^n - Bq$, (26b) is fulfilled with probability at most $\frac{\tau_2 Bq}{2^n - Bq}$. By $\neg\text{aux}_3(\mathcal{Q}_q)$, each hit adds at most τ_3 solutions (for fixed $(1, x_i, z_i)$, there are at most τ_3 choices for $(2, x_k, z_k), (3, x_l, z_l)$ to satisfy (26b)).

Regular (forward or inverse) query to $\bar{\pi}_2$. Without loss of generality, consider a forward query. Denote the query by $(2, x_k, z_k)$. There are at most Bq choices for $(3, x_l, z_l)$. As z_k is randomly drawn from a set of size at least $2^n - Bq$, (26a) is fulfilled with probability at most $\frac{Bq}{2^n - Bq}$. Again, by $\neg\text{aux}_3(\mathcal{Q}_q)$, each hit adds at most τ_3 solutions.

Regular (forward or inverse) query to $\bar{\pi}_3$. Consider (26) with (26a) added to (26b). Without loss of generality, consider a forward query. Denote the query by $(3, x_l, z_l)$. There are at most Bq choices for $(1, x_i, z_i)$. As z_l is randomly drawn from a set of size at least $2^n - Bq$, the replaced (26b) is fulfilled with probability at most $\frac{Bq}{2^n - Bq}$. Again, by $\neg\text{aux}_3(\mathcal{Q}_q)$, each hit adds at most τ_3 solutions.

Predicate query. The same reasoning as for regular queries applies with the difference that any query of the B -set may be successful and that the bound of Lem. 2 part (ii) applies: $\frac{\tau_2 B^2 \delta_{B,C}[1]q}{2^n - Bq}$. Again, by $\neg \text{aux}_3(\mathcal{Q}_q)$, each hit adds at most $\tau_3 B$ solutions.

Conclusion. Thus, any query results in a solution with probability at most $\frac{\tau_2 B^2 \delta_{B,C}[1]q}{2^n - Bq}$, and each hit adds at most $\tau_3 B$ solutions. Consequently, more than τ_4 solutions are obtained with probability at most $\left(\frac{q}{\tau_3 B}\right) \left(\frac{\tau_2 B^2 \delta_{B,C}[1]q}{2^n - Bq}\right)^{\frac{\tau_4+1}{\tau_3 B}} \leq \left(\frac{e \tau_2 \tau_3 B^3 \delta_{B,C}[1]q^2}{(\tau_4+1)(2^n - Bq)}\right)^{\frac{\tau_4+1}{\tau_3 B}}$. Considering any possible choice for Z, Z' , we obtain the claimed bound. \square

Lemma 10. $\Pr(\text{aux}_5(\mathcal{Q}_q) \mid \neg \text{aux}_{2 \vee 3 \vee 4}(\mathcal{Q}_q)) \leq$

$$2^n \left(\frac{e(\tau_3 \binom{B}{2} + \tau_2 \tau_3 B + \tau_4 B)q}{\tau_5 + 1} \left(\frac{2\tau_3 B^2 \delta_{B,C}[1]q}{2^n - Bq} + \frac{B^2 \delta_{B,C}[2]q^2}{2^{2n} - Bq} \right) \right)^{\frac{\tau_5+1}{\tau_3 \binom{B}{2} + \tau_2 \tau_3 B + \tau_4 B}}.$$

Proof. Let $Z \in \{0, 1\}^n$. The goal is to find distinct tuples $(1, x_i, z_i), (1, x_j, z_j), (2, x_k, z_k), (3, x_l, z_l) \in \mathcal{Q}_q$ such that

$$x_j \oplus z_j \oplus x_k \oplus z_k \oplus x_l = 0, \quad (27a)$$

$$x_i \oplus z_i \oplus x_k \oplus z_k \oplus x_l \oplus z_l = Z. \quad (27b)$$

Consider the i^{th} query. A special treatment is needed for predicate queries to $\bar{\pi}_1$ as it may contribute two queries to one solution.

Regular (forward or inverse) query to $\bar{\pi}_1$. Without loss of generality, consider a forward query. Note that it may contribute as tuple $(1, x_i, z_i)$ or as tuple $(1, x_j, z_j)$ (but not both).

First consider it a contribution to $(1, x_i, z_i)$. There are at most Bq choices for $(3, x_l, z_l)$. For any such choice, by $\neg \text{aux}_3(\mathcal{Q}_q)$, there are at most τ_3 choices for $(1, x_j, z_j), (2, x_k, z_k)$ to fulfill (27a). As z_i is randomly drawn from a set of size at least $2^n - Bq$, (27b) is fulfilled with probability at most $\frac{\tau_3 Bq}{2^n - Bq}$. By $\neg \text{aux}_{2 \vee 3}(\mathcal{Q}_q)$, each hit adds at most $\tau_2 \tau_3$ solutions (for fixed $(1, x_i, z_i)$, there are at most τ_3 choices for $(2, x_k, z_k), (3, x_l, z_l)$ to satisfy (27b) and at most τ_2 choices $(1, x_j, z_j)$ to then satisfy (27a)).

Next, consider it a contribution to $(1, x_j, z_j)$. There are at most Bq choices for $(3, x_l, z_l)$. For any such choice, by $\neg \text{aux}_3(\mathcal{Q}_q)$, there are at most τ_3 choices for $(1, x_i, z_i), (2, x_k, z_k)$ to fulfill (27b). As z_j is randomly drawn from a set of size at least $2^n - Bq$, (27a) is fulfilled with probability at most $\frac{\tau_3 Bq}{2^n - Bq}$. By $\neg \text{aux}_4(\mathcal{Q}_q)$, each hit adds at most τ_4 solutions (for fixed $(1, x_j, z_j)$, there are at most τ_4 choices for $(1, x_i, z_i), (2, x_k, z_k), (3, x_l, z_l)$ to satisfy (27)).

In total, the query is successful with probability at most $\frac{2\tau_3 Bq}{2^n - Bq}$ and each hit adds at most $\tau_2 \tau_3 + \tau_4$ solutions.

Regular (forward or inverse) query to $\bar{\pi}_2$. Consider (27) with (27a) added to (27b). Without loss of generality, consider a forward query. Denote the query by $(2, x_k, z_k)$. There are at most Bq choices for $(3, x_l, z_l)$. For any such choice, by $\neg \text{aux}_3(\mathcal{Q}_q)$, there are at most τ_3 choices for $(1, x_i, z_i), (1, x_j, z_j)$ to fulfill the replaced (27b). As z_k is randomly drawn from a set of size at least $2^n - Bq$, (27a) is fulfilled with probability at most $\frac{\tau_3 Bq}{2^n - Bq}$. Again, by $\neg \text{aux}_{2 \vee 3}(\mathcal{Q}_q)$, each hit adds at most $\tau_2 \tau_3$ solutions.

Regular (forward or inverse) query to $\bar{\pi}_3$. Consider (27) with (27a) added to (27b). Without loss of generality, consider a forward query. Denote the query by $(3, x_l, z_l)$. There are at most Bq choices for $(1, x_i, z_i)$. By $\neg \text{aux}_3(\mathcal{Q}_q)$, there are at most τ_3 choices for $(1, x_j, z_j), (2, x_k, z_k)$ to fulfill (27a). As z_l is randomly drawn from a set of size at least $2^n - Bq$, (27a) is fulfilled with probability at most $\frac{\tau_3 Bq}{2^n - Bq}$. Again, by $\neg \text{aux}_{2 \vee 3}(\mathcal{Q}_q)$, each hit adds at most $\tau_2 \tau_3$ solutions.

Predicate query to $\bar{\pi}_1$. Denote the query response by $\{(x_i^1, z_i^1), \dots, (x_i^B, z_i^B)\}$. In case the B -set contributes at one position only (either $(1, x_i, z_i)$ or $(1, x_j, z_j)$), the same reasoning as for regular queries applies with the difference that any query of the B -set may be successful and that the bound of Lem. 2 part (ii) applies: $\frac{2\tau_3 B^2 \delta_{B,C}[1]q}{2^n - Bq}$. Again, by $\neg \text{aux}_{2 \vee 3 \vee 4}(\mathcal{Q}_q)$, each hit adds at most $\tau_2 \tau_3 B + \tau_4 B$ solutions.

Now consider the case the predicate query contributes to both $(1, x_i, z_i)$ and $(1, x_j, z_j)$. There are at most $(Bq)^2$ choices for $(2, x_k, z_k), (3, x_l, z_l)$. By Lem. 2 part (iii), we find that (27) is fulfilled with probability at most $\frac{B^2\delta_{B,C}[2]q^2}{2^{2n}-Bq}$. By $\neg\text{aux}_3(\mathcal{Q}_q)$, each hit adds at most $\tau_3\binom{B}{2}$ solutions.

Predicate query to $\bar{\pi}_2$ or $\bar{\pi}_3$. The same reasoning as for regular queries applies with the difference that any query of the B -set may be successful and that the bound of Lem. 2 part (ii) applies: $\frac{\tau_3 B^2 \delta_{B,C}[1]q}{2^n - Bq}$. Again, by $\neg\text{aux}_{2\vee 3}(\mathcal{Q}_q)$, each hit adds at most $\tau_2\tau_3 B$ solutions.

Conclusion. Thus, any query results in a solution with probability at most $\frac{2\tau_3 B^2 \delta_{B,C}[1]q}{2^n - Bq} + \frac{B^2 \delta_{B,C}[2]q^2}{2^{2n} - Bq}$, and each hit adds at most $\tau_3\binom{B}{2} + \tau_2\tau_3 B + \tau_4 B$ solutions. Consequently, more than τ_5 solutions are obtained with probability at most

$$\begin{aligned} & \left(\frac{q}{\tau_3\binom{B}{2} + \tau_2\tau_3 B + \tau_4 B} \right) \left(\frac{2\tau_3 B^2 \delta_{B,C}[1]q}{2^n - Bq} + \frac{B^2 \delta_{B,C}[2]q^2}{2^{2n} - Bq} \right)^{\frac{\tau_5 + 1}{\tau_3\binom{B}{2} + \tau_2\tau_3 B + \tau_4 B}} \\ & \leq \left(\frac{e(\tau_3\binom{B}{2} + \tau_2\tau_3 B + \tau_4 B)q}{\tau_5 + 1} \left(\frac{2\tau_3 B^2 \delta_{B,C}[1]q}{2^n - Bq} + \frac{B^2 \delta_{B,C}[2]q^2}{2^{2n} - Bq} \right) \right)^{\frac{\tau_5 + 1}{\tau_3\binom{B}{2} + \tau_2\tau_3 B + \tau_4 B}}. \end{aligned}$$

Considering any possible choice for Z , we obtain the claimed bound. \square

From (21), (22), and the results of Lems. 4-10 we find (recall that in Lem. 5 we require $\tau_2 = 1$):

$$\begin{aligned} \text{Adv}_{\text{FSS}}^{\text{col}}(q) & \leq \frac{(\tau_3^2 + \tau_5)B^2\delta_{B,C}[1]q^2 + B^2q(q+1)}{2^n - Bq} + \frac{e^2 B^2 (\delta_{B,C}[1])^2 q^2 2^{n-|C|} + e^2 B^2 q^2 2^n}{4(2^n - Bq)^2} + \\ & \quad \binom{B}{2} \frac{(\tau_3^2 q 2^n + \tau_3 B^2 q^3 + q 2^n) \delta_{B,C}[2]}{2^{2n} - Bq} + \binom{B}{2} \frac{B^4 \delta_{B,C}[2]q^5}{2^{3n} - Bq} + \\ & \quad 2^n \left(\frac{e(B + \binom{B}{2})q}{\tau_3 + 1} \left(\frac{B^2 \delta_{B,C}[1]q}{2^n - Bq} + \frac{\delta_{B,C}[2]2^n}{2^{2n} - Bq} \right) \right)^{\frac{\tau_3 + 1}{B + \binom{B}{2}}} + \\ & \quad 2^{2n} \left(\frac{e\tau_3 B^3 \delta_{B,C}[1]q^2}{(\tau_4 + 1)(2^n - Bq)} \right)^{\frac{\tau_4 + 1}{\tau_3 B}} + \\ & \quad 2^n \left(\frac{e(\tau_3\binom{B+1}{2} + \tau_4 B)q}{\tau_5 + 1} \left(\frac{2\tau_3 B^2 \delta_{B,C}[1]q}{2^n - Bq} + \frac{B^2 \delta_{B,C}[2]q^2}{2^{2n} - Bq} \right) \right)^{\frac{\tau_5 + 1}{\tau_3\binom{B+1}{2} + \tau_4 B}}. \end{aligned}$$

The proof of Prop. 7 is completed by making the following simplifications: we use the fact that $2^n - Bq \geq 2^{n-1}$ for $Bq \leq 2^{n-1}$, and put $\tau_3 = \tau - 1$, $\tau_4 = \tau^2 - \tau - 1$, and $\tau_5 = 2\tau^3 + \tau^2(B-1) - \tau(B+3) - 1$ for some integral $\tau > 0$. Additionally, we use $q+1 \leq 2q$ for $q \geq 1$.

H Tightness of the Bound of Theorem 5

For the cases $B = 1$ and C arbitrary, and $B = 2$ and C arbitrary such that $|C| > n/2$, we derive generic attacks that demonstrate tightness of the bounds of Thm. 5.

Proposition 8 ($B = 1$). *Let $n \in \mathbb{N}$. Suppose $(\pi_1, \pi_2, \pi_3) \stackrel{\$}{\leftarrow} \text{BC}[\Phi(A, 1, \varphi^C)](n)^3$. Then, $\text{Adv}_{\text{FSS}}^{\text{col}}(q) \geq \frac{q^2}{2^{n-|C|}}$.*

Proof. We construct a collision-finding adversary \mathcal{A} whose goal is to find a collision where $x_2 = x'_2$. By construction, to find a collision of such form for FSS (see (8)) it suffices to find two distinct values x_1, x'_1 such that $x_1 \oplus \pi_1(x_1) = x'_1 \oplus \pi_1(x'_1)$. To this end, the adversary makes

predicate queries to π_1^Φ on input of distinct values y to obtain q queries (x_y, z_y) satisfying $\text{Bits}_C(x_y \oplus z_y) = 0$. Any two such queries collide on the entire state, $x_y \oplus z_y = x_{y'} \oplus z_{y'}$, with probability at least $\frac{q^2}{2^{n-|C|}}$. \square

Proposition 9 ($B = 2$ and $|C| > n/2$). *Let $n \in \mathbb{N}$. Suppose $(\pi_1, \pi_2, \pi_3) \stackrel{\$}{\leftarrow} \text{BC}[\Phi(A, 2, \varphi^C)](n)^3$. Then, $\text{Adv}_{\text{FSS}}^{\text{col}}(q) \geq \frac{q^2}{2^{n-|C|}}$.*

Proof. We construct a collision-finding adversary \mathcal{A} whose goal is to find a collision where $x_2 = x'_2$. By construction, to find a collision of such form for FSS (see (8)) it suffices to find two distinct values x_1, x'_1 such that $x_1 \oplus \pi_1(x_1) = x'_1 \oplus \pi_1(x'_1)$. To this end, the adversary makes predicate queries to π_1^Φ on input of distinct values y to obtain q 2-sets $\{(x_y^1, z_y^1), (x_y^2, z_y^2)\}$ satisfying $\text{Bits}_C(x_y^1 \oplus z_y^1) = \text{Bits}_C(x_y^2 \oplus z_y^2)$. These two queries collide on the entire state, $x_y^1 \oplus z_y^1 = x_y^2 \oplus z_y^2$, with probability at least $\frac{1}{2^{n-|C|}}$. If the adversary makes q predicate queries, we directly obtain our bound. \square

I Proof of Theorem 6

The bulk of the proof is captured in the following proposition.

Proposition 10. *Let $n \in \mathbb{N}$. Suppose $(\pi_1, \pi_2, \pi_3) \stackrel{\$}{\leftarrow} \text{BC}[\Phi(A, B, \varphi^C)](n)^3$. Then, for $q \leq 2^{n-1}/B$ and any positive integral values τ, τ', τ'' ,*

$$\begin{aligned} \text{Adv}_{\text{FSS}}^{\text{epre}}(q) &\leq \frac{2\tau B^3 q^2 + 4\tau''\tau[B=1]B2^{|C|}q + 2\tau'\tau B^3 q + 4B^2 q^2}{2^n} + \frac{2[B=1]2^{|C|}q^3}{2^{2n}} + \\ &2^{n-|C|} \left(\frac{2eB\delta_{B,C}[1]q}{\tau 2^n} \right)^\tau + 2^n \left(\frac{2eBq}{\tau 2^n} \right)^\tau + 2^n \left(\frac{2eB^2\delta_{B,C}[1]q^2}{\tau' 2^n} \right)^{\tau'} + \\ &[B=1] \left(\frac{2eq}{\tau'' 2^{|C|}} \right)^{\tau''} + 2[B=1] \left(\frac{2e2^{|C|}q}{2^n} + \frac{2e\tau Bq}{\tau'' 2^{|C|}} + \frac{2eq^2}{\tau'' 2^n} \right)^{\tau''}. \end{aligned} \quad (28)$$

The proof follows later in the appendix, but we first prove Thm. 6 using Prop. 10.

Proof of Theorem 6

(i) and (ii). We put $\tau = n$, $\tau' = n2^{|C|}$, and $\tau'' = n2^{\max\{0, n/2 - |C|\}}$. Then, for the particular choice of B (with C arbitrary) the bound (28) of Prop. 10 simplifies to

$$\begin{aligned} \text{Adv}_{\text{FSS}}^{\text{epre}}(q) &\leq \frac{2nq^2 + 4q^2 + 4eq^2 + 6n^2 2^{\max\{|C|, n/2\}} q + 4e2^{|C|}q}{2^n} + \frac{2 \cdot 2^{|C|}q^3}{2^{2n}} + \\ &\frac{6eq}{2^{\max\{|C|, n/2\}}} + 2^{n-|C|} \left(\frac{2e2^{|C|}q}{n2^n} \right)^n + 2 \cdot 2^n \left(\frac{2eq^2}{n2^n} \right)^n. \end{aligned}$$

For $q = 2^{n/2-n\epsilon}$ (in case $|C| \leq n/2$), and for $q = 2^{n-|C|-n\epsilon}$ (in case $|C| > n/2$), it is clear that all terms approach 0 for $n \rightarrow \infty$.

(iii). We put $\tau = 2$ and $\tau' = n$ (τ'' disappears for $B \geq 2$). The reasoning of previous cases carries over, and we will not go into detail. Note that we have $\delta_{B,C}[1] = 1$ for $B \geq 2$.

Proof of Proposition 10

We consider any adversary that has access to $(\pi_1, \pi_2, \pi_3) \stackrel{\$}{\leftarrow} \text{BC}[\Phi(A, B, \varphi^C)](n)^3$ and makes q queries. As in App. G, these queries are stored in a query history \mathcal{Q} as indexed tuples of the form (κ_i, x_i, z_i) for regular queries and $\{(\kappa_i, x_i^1, z_i^1), \dots, (\kappa_i, x_i^B, z_i^B)\}$ for predicate queries. We use the same approach as in App. G. In particular, we make the same transition from (π_1, π_2, π_3) to $(\bar{\pi}_1, \bar{\pi}_2, \bar{\pi}_3) \stackrel{\$}{\leftarrow} \overline{\text{BC}}[\Phi(A, B, \varphi^C)](n)^3$.

Let $Z \in \{0, 1\}^n$. A preimage for Z corresponds to obtaining query pairs (x_1, z_1) for $\bar{\pi}_1$, (x_2, z_2) for $\bar{\pi}_2$, and (x_3, z_3) for $\bar{\pi}_3$ in the query history, such that:

$$x_1 \oplus z_1 \oplus x_2 \oplus z_2 = x_3, \quad (29a)$$

$$x_1 \oplus z_1 \oplus z_3 = Z. \quad (29b)$$

We define this configuration by $\mathbf{pre}(\mathcal{Q}_q)$. This means:

$$\mathbf{Adv}_{\text{FSS}}^{\text{epre}}(q) = \mathbf{Pr}(\mathbf{pre}(\mathcal{Q}_q)) + \frac{B^2 q(q+1)}{2^n - Bq}, \quad (30)$$

where the second part of the bound comes from the transition from WCM to AWCM.

As before, for the analysis of $\mathbf{Pr}(\mathbf{pre}(\mathcal{Q}_q))$ we introduce an auxiliary event $\mathbf{aux}(\mathcal{Q}_q)$. Let $\tau_1, \dots, \tau_5 > 0$ be any integral values. We define $\mathbf{aux}(\mathcal{Q}_q) = \mathbf{aux}_{1 \vee \dots \vee 5}(\mathcal{Q}_q)$, where

$$\mathbf{aux}_1(\mathcal{Q}_q) : \left| \left\{ (\kappa_i, x_i, z_i) \in \mathcal{Q}_q \mid B = 1 \wedge \text{Bits}_C(x_i) = 0 \wedge \text{Bits}_C(z_i) = Z \right\} \right| > \tau_1;$$

$$\mathbf{aux}_2(\mathcal{Q}_q) : \max_{Z \in \{0, 1\}^n} \left| \left\{ (\kappa_i, x_i, z_i) \in \mathcal{Q}_q \mid x_i \oplus z_i = Z \right\} \right| > \tau_2;$$

$$\mathbf{aux}_3(\mathcal{Q}_q) : \max_{Z \in \{0, 1\}^n} \left| \left\{ \begin{array}{l} (\kappa_i, x_i, z_i), (\kappa_j, x_j, z_j) \in \mathcal{Q}_q \\ \kappa_i \neq \kappa_j \wedge x_i \oplus z_i \oplus x_j \oplus z_j = Z \end{array} \right\} \right| > \tau_3;$$

$$\mathbf{aux}_4(\mathcal{Q}_q) : \left| \left\{ \begin{array}{l} (2, x_i, z_i), (3, x_j, z_j) \in \mathcal{Q}_q \\ B = 1 \wedge x_i \oplus z_i \oplus x_j \oplus z_j = Z \wedge \text{Bits}_C(x_i \oplus z_i \oplus x_j) = 0 \end{array} \right\} \right| > \tau_4;$$

$$\mathbf{aux}_5(\mathcal{Q}_q) : \left| \left\{ \begin{array}{l} (1, x_i, z_i), (3, x_j, z_j) \in \mathcal{Q}_q \\ B = 1 \wedge x_i \oplus z_i \oplus z_j = Z \wedge \text{Bits}_C(x_i \oplus z_i \oplus x_j) = 0 \end{array} \right\} \right| > \tau_5.$$

Here, $\mathbf{aux}_{2 \vee 3}(\mathcal{Q}_q)$ is similar to $\mathbf{aux}_{2 \vee 3}(\mathcal{Q}_q)$ of App. G. Events $\mathbf{aux}_{1 \vee 4 \vee 5}(\mathcal{Q}_q)$ are only employed when $B = 1$ as a special treatment is needed for this case, and they only apply to Z being the target range value. Again, by basic probability theory, we obtain for (30):

$$\mathbf{Pr}(\mathbf{pre}(\mathcal{Q}_q)) \leq \mathbf{Pr}(\mathbf{pre}(\mathcal{Q}_q) \wedge \neg \mathbf{aux}(\mathcal{Q}_q)) + \mathbf{Pr}(\mathbf{aux}(\mathcal{Q}_q)). \quad (31)$$

We consider $\mathbf{Pr}(\mathbf{pre}(\mathcal{Q}_q) \wedge \neg \mathbf{aux}(\mathcal{Q}_q))$ in Lem. 11. For $\mathbf{Pr}(\mathbf{aux}(\mathcal{Q}_q))$, note that

$$\mathbf{Pr}(\mathbf{aux}(\mathcal{Q}_q)) \leq \sum_{i=1}^2 \mathbf{Pr}(\mathbf{aux}_i(\mathcal{Q}_q)) + \sum_{i=3}^5 \mathbf{Pr}(\mathbf{aux}_i(\mathcal{Q}_q) \mid \neg \mathbf{aux}_{1 \vee 2}(\mathcal{Q}_q)). \quad (32)$$

The probabilities on $\mathbf{aux}_1(\mathcal{Q}_q)$, $\mathbf{aux}_2(\mathcal{Q}_q)$, and $\mathbf{aux}_3(\mathcal{Q}_q)$ are straightforward and bounded in Lem. 12. The bounds on $\mathbf{aux}_4(\mathcal{Q}_q)$ and $\mathbf{aux}_5(\mathcal{Q}_q)$ are more involved and discussed in Lem. 13 and Lem. 14, respectively. Afterwards, all lemmas are combined to obtain a bound on $\mathbf{Adv}_{\text{FSS}}^{\text{epre}}(q)$.

Lemma 11. $\mathbf{Pr}(\mathbf{pre}(\mathcal{Q}_q) \wedge \neg \mathbf{aux}(\mathcal{Q}_q)) \leq \frac{(\tau_4 + \tau_5)[B = 1]2^{|C|}q + \tau_2 B^2 q^2 + \tau_3 Bq}{2^n - Bq} + \frac{[B = 1]2^{|C|}q^3}{2^{2n} - Bq}.$

Proof. Note that (29) can be equivalently stated as

$$x_1 \oplus z_1 \oplus x_2 \oplus z_2 = x_3, \quad (33a)$$

$$x_2 \oplus z_2 \oplus x_3 \oplus z_3 = Z, \quad (33b)$$

by XORing the two equations. We will use both representations interchangeably. For the analysis of $\mathbf{pre}(\mathcal{Q}_q) \wedge \neg \mathbf{aux}(\mathcal{Q}_q)$, we say that the i^{th} query ($i \in \{1, \dots, q\}$) is successful if it makes configuration $\mathbf{pre}(\mathcal{Q}_i)$ satisfied and $\neg \mathbf{aux}(\mathcal{Q}_i)$ holds. Now, by basic probability theory, we can analyze the probability of the i^{th} query being successful, and sum over $i = 1, \dots, q$.

Let $i \in \{1, \dots, q\}$. Clearly, if $\neg \text{aux}_i(\mathcal{Q}_i)$ holds, the i^{th} query can certainly not be successful, so we assume $\neg \text{aux}(\mathcal{Q}_i)$ and analyze the probability the i^{th} query makes $\text{pre}(\mathcal{Q}_i)$ satisfied. For regular (forward or inverse) queries, the analysis of [41] mostly carries over. The difficulty lies in the analysis of predicate queries.

Query $\bar{\pi}_1(\mathbf{x}_1)$ or $\bar{\pi}_1^{-1}(\mathbf{z}_1)$. The cases are the same by symmetry, and we consider $\bar{\pi}_1(x_1)$ only. Denote the query response by z_1 . By $\neg \text{aux}_3(\mathcal{Q}_q)$, there are at most τ_3 choices for $(x_2, z_2), (x_3, z_3)$ to fulfill (33b). As z_1 is randomly drawn from a set of size at least $2^n - Bq$, (33a) is fulfilled with probability at most $\frac{\tau_3}{2^n - Bq}$.

Query $\bar{\pi}_2(x_2)$ or $\bar{\pi}_2^{-1}(z_2)$. The cases are the same by symmetry, and we consider $\bar{\pi}_2(x_2)$ only. Denote the query response by z_2 . There are at most Bq choices for (x_3, z_3) . By $\neg \text{aux}_2(\mathcal{Q}_q)$, there are at most τ_2 choices for (x_1, z_1) to fulfill (29b). As z_2 is randomly drawn from a set of size at least $2^n - Bq$, (29a) is fulfilled with probability at most $\frac{\tau_2 Bq}{2^n - Bq}$.

Query $\bar{\pi}_3(x_3)$. Denote the query response by z_3 . By $\neg \text{aux}_3(\mathcal{Q}_q)$, there are at most τ_3 choices for $(x_1, z_1), (x_2, z_2)$ to fulfill (33a). As z_3 is randomly drawn from a set of size at least $2^n - Bq$, (33b) is fulfilled with probability at most $\frac{\tau_3}{2^n - Bq}$.

Query $\bar{\pi}_3^{-1}(z_3)$. Denote the query response by x_3 . There are at most Bq choices for (x_2, z_2) . By $\neg \text{aux}_2(\mathcal{Q}_q)$, there are at most τ_2 choices for (x_1, z_1) to fulfill (29b). As x_3 is randomly drawn from a set of size at least $2^n - Bq$, (29a) is fulfilled with probability at most $\frac{\tau_2 Bq}{2^n - Bq}$.

Query $\bar{\pi}_1^\Phi(\mathbf{y}_1)$. Denote the query response by $\{(x_1^1, z_1^1), \dots, (x_1^B, z_1^B)\}$. We make a distinction between $B = 1$ and $B > 1$. First, if $B > 1$, the same reasoning as for regular queries applies with the difference that any query of the B -set may be successful and that the bound of Lem. 2 part (ii) applies: $[B > 1] \frac{\tau_3 B \delta_{B,C}[1]}{2^n - Bq} \leq \frac{\tau_3 B}{2^n - Bq}$. Now, consider $B = 1$. By $\neg \text{aux}_4(\mathcal{Q}_q)$, there are at most τ_4 choices for $(x_2, z_2), (x_3, z_3)$ to satisfy $x_2 \oplus z_2 \oplus x_3 \oplus z_3 = Z$, hence to fulfill (33b), and $\text{Bits}_C(x_2 \oplus z_2 \oplus x_3) = 0$. By Lem. 2 part (ii), the predicate query fulfills (33) with probability at most $[B = 1] \frac{\tau_4 2^{|C|}}{2^n - Bq}$. Thus, the predicate query succeeds with probability at most $\frac{\tau_4 [B=1] 2^{|C|} + \tau_3 B}{2^n - Bq}$.

Query $\bar{\pi}_2^\Phi(\mathbf{y}_2)$. Denote the query response by $\{(x_2^1, z_2^1), \dots, (x_2^B, z_2^B)\}$. We make a distinction between $B = 1$ and $B > 1$. First, if $B > 1$, the same reasoning as for regular queries applies with the difference that any query of the B -set may be successful and that the bound of Lem. 2 part (ii) applies: $[B > 1] \frac{\tau_2 B^2 \delta_{B,C}[1]q}{2^n - Bq} \leq \frac{\tau_2 B^2 q}{2^n - Bq}$. Now, consider $B = 1$. By $\neg \text{aux}_5(\mathcal{Q}_q)$, there are at most τ_5 choices for $(x_1, z_1), (x_3, z_3)$ to satisfy $x_1 \oplus z_1 \oplus x_3 = Z$, hence to fulfill (29b), and $\text{Bits}_C(x_1 \oplus z_1 \oplus x_3) = 0$. By Lem. 2 part (ii), the predicate query fulfills (29) with probability at most $[B = 1] \frac{\tau_5 2^{|C|}}{2^n - Bq}$. Thus, the predicate query succeeds with probability at most $\frac{\tau_5 [B=1] 2^{|C|} + \tau_2 B^2 q}{2^n - Bq}$.

Query $\bar{\pi}_3^\Phi(\mathbf{y}_3)$. Denote the query response by $\{(x_3^1, z_3^1), \dots, (x_3^B, z_3^B)\}$. We make a distinction between $B = 1$ and $B > 1$. First, if $B > 1$, the same reasoning as for regular forward queries applies with the difference that any query of the B -set may be successful and that the bound of Lem. 2 part (ii) applies: $[B > 1] \frac{\tau_3 B \delta_{B,C}[1]}{2^n - Bq} \leq \frac{\tau_3 B}{2^n - Bq}$. Now, consider $B = 1$. There are at most q^2 choices for $(x_1, z_1), (x_2, z_2)$. Similar to Lem. 2 part (iii), one can prove that the predicate query fulfills (29) with probability at most $[B = 1] \frac{2^{|C|} q^2}{2^{2n} - Bq}$. Thus, the predicate query succeeds with probability at most $\frac{[B=1] 2^{|C|} q^2}{2^{2n} - Bq} + \frac{\tau_3 B}{2^n - Bq}$.

Conclusion. Taking the maximum of all success probabilities, the i^{th} query is successful with probability at most $\frac{(\tau_4 + \tau_5)[B=1] 2^{|C|} + \tau_2 B^2 q + \tau_3 B}{2^n - Bq} + \frac{[B=1] 2^{|C|} q^2}{2^{2n} - Bq}$. The claimed bound is obtained by summing over $i = 1, \dots, q$. \square

Lemma 12. $\Pr(\text{aux}_1(\mathcal{Q}_q)) + \Pr(\text{aux}_2(\mathcal{Q}_q)) + \Pr(\text{aux}_3(\mathcal{Q}_q) \mid \neg \text{aux}_{1 \vee 2}(\mathcal{Q}_q)) \leq$

$$[B = 1] \left(\frac{e 2^{n-|C|} q}{(\tau_1 + 1)(2^n - q)} \right)^{\tau_1 + 1} + 2^{n-|C|} \left(\frac{e B^2 \delta_{B,C}[1] q}{(\tau_2 + 1)(2^n - Bq)} \right)^{\frac{\tau_2 + 1}{B}} +$$

$$2^n \left(\frac{eB^2q}{(\tau_2 + 1)(2^n - Bq)} \right)^{\frac{\tau_2+1}{B}} + 2^n \left(\frac{e\tau_2 B^3 \delta_{B,C}[1]q^2}{(\tau_3 + 1)(2^n - Bq)} \right)^{\frac{\tau_3+1}{\tau_2 B}}.$$

Proof. We bound these probabilities separately.

Pr (aux₁(Q_q)). Recall that aux₁(Q_q) only applies to B = 1. Let Z be the challenge range value. Consider the ith query. If this is a regular (forward or inverse) query (κ_i, x_i, z_i), it satisfies the condition with probability at most $\frac{2^{n-|C|}}{2^n - q}$. If it is a predicate query (κ_i, x_i, z_i), it satisfies the condition with probability at most $\frac{2^{n-|C|}}{2^n - q}$ (Lem. 2 part (i)). More than τ₁ solutions are obtained with probability at most $\left(\frac{q}{\tau_1 + 1}\right) \left(\frac{2^{n-|C|}}{2^n - q}\right)^{\tau_1 + 1} \leq \left(\frac{e2^{n-|C|}q}{(\tau_1 + 1)(2^n - q)}\right)^{\tau_1 + 1}$. As the event only applies to the challenge value Z, we obtain:

$$\Pr(\text{aux}_1(Q_q)) \leq [B = 1] \left(\frac{e2^{n-|C|}q}{(\tau_1 + 1)(2^n - q)} \right)^{\tau_1 + 1}. \quad (34)$$

Pr (aux₂(Q_q)). The analysis differs from the one of Lem. 7 in the fact that each hit may add at most B solutions (rather than 1). Taking this into account results in the following bound:

$$\Pr(\text{aux}_2(Q_q)) \leq 2^{n-|C|} \left(\frac{eB^2 \delta_{B,C}[1]q}{(\tau_2 + 1)(2^n - Bq)} \right)^{\frac{\tau_2+1}{B}} + 2^n \left(\frac{eB^2q}{(\tau_2 + 1)(2^n - Bq)} \right)^{\frac{\tau_2+1}{B}}. \quad (35)$$

Pr (aux₃(Q_q) | ¬aux₂(Q_q)). Event aux₃(Q_q) differs from aux₃(Q_q) in the fact that a single predicate query cannot make the entire equation satisfied. Leaving out this special case in the analysis of Lem. 8 results in the following bound:

$$\Pr(\text{aux}_3(Q_q) \mid \neg \text{aux}_2(Q_q)) \leq 2^n \left(\frac{e\tau_2 B^3 \delta_{B,C}[1]q^2}{(\tau_3 + 1)(2^n - Bq)} \right)^{\frac{\tau_3+1}{\tau_2 B}}. \quad (36)$$

The claim is obtained by adding (34-36). \square

Lemma 13. $\Pr(\text{aux}_4(Q_q) \mid \neg \text{aux}_{1 \vee 2}(Q_q)) \leq [B = 1] \left(\frac{e\tau_1 \tau_2 2^{|C|}q + e\tau_2 q^2}{(\tau_4 + 1)(2^n - q)} \right)^{\frac{\tau_4+1}{\tau_2}}$.

Proof. Recall that aux₄(Q_q) only applies to B = 1. Let Z be the challenge range value. The goal is to find distinct tuples (2, x₂, z₂) and (3, x₃, z₃) ∈ Q_q such that x₂ ⊕ z₂ ⊕ x₃ ⊕ z₃ = Z and Bits_C(x₂ ⊕ z₂ ⊕ x₃) = 0. Consider the ith query.

Regular (forward or inverse) query. We follow the analysis of aux₃(Q_q), only focussing on x₂ ⊕ z₂ ⊕ x₃ ⊕ z₃ = Z. The query is successful with probability at most $\frac{q}{2^n - q}$ and each hit adds at most τ₂ solutions.

Predicate query to π₂. Denote the query response by (x₂, z₂), which will satisfy Bits_C(x₂ ⊕ z₂) = 0. By ¬aux₁(Q_q), there are at most τ₁ choices for (x₃, z₃) to satisfy Bits_C(x₃) = 0 and Bits_C(z₃) = Bits_C(Z). By Lem. 2 part (ii), we find that the equation is satisfied with probability at most $\frac{\tau_1 2^{|C|}}{2^n - q}$. By ¬aux₂(Q_q), each hit adds at most τ₂ solutions.

Predicate query to π₃. Denote the query response by (x₃, z₃), which will satisfy Bits_C(x₃ ⊕ z₃) = 0. There are at most q choices for (x₂, z₂). Similar to Lem. 2 part (ii), one can prove that the equation is satisfied with probability at most $\frac{2^{n-|C|}}{2^n 2^{n-|C|-q}} \cdot q \leq \frac{q}{2^n - q}$: we need the predicate query to satisfy Bits_C(x₃) = Bits_C(x₂ ⊕ z₂), Bits_C(x₃ ⊕ z₃) = 0 = Bits_C(x₂ ⊕ z₂ ⊕ Z), and Bits_{C̄}(x₃ ⊕ z₃) = Bits_{C̄}(x₂ ⊕ z₂ ⊕ Z), and this gives at most 2^{n-|C|} possible tuples of $\bar{\Sigma}_k(P, P^q)$. By ¬aux₂(Q_q), each hit adds at most τ₂ solutions.

Conclusion. Thus, if B = 1, any query results in a solution with probability at most $\frac{\tau_1 2^{|C|} + q}{2^n - q}$, and each hit adds at most τ₂ solutions. Consequently, more than τ₄ solutions are obtained with probability at most $\left(\frac{q}{\tau_4 + 1}\right) \left(\frac{\tau_1 2^{|C|} + q}{2^n - q}\right)^{\frac{\tau_4+1}{\tau_2}} \leq \left(\frac{e\tau_1 \tau_2 2^{|C|}q + e\tau_2 q^2}{(\tau_4 + 1)(2^n - q)}\right)^{\frac{\tau_4+1}{\tau_2}}$. \square

Lemma 14. $\Pr(\text{aux}_5(\mathcal{Q}_q) \mid \neg \text{aux}_{1 \vee 2}(\mathcal{Q}_q)) \leq$
 $[B = 1] \left(\frac{e\tau_1\tau_2 2^{|C|}q + e\tau_2^2 2^{n-|C|}q + e\tau_2 q^2}{(\tau_5 + 1)(2^n - q)} \right)^{\frac{\tau_5+1}{\tau_2}}.$

Proof. Recall that $\text{aux}_5(\mathcal{Q}_q)$ only applies to $B = 1$. Let Z be the challenge range value. The goal is to find distinct tuples $(1, x_1, z_1)$ and $(3, x_3, z_3) \in \mathcal{Q}_q$ such that $x_1 \oplus z_1 \oplus z_3 = Z$ and $\text{Bits}_C(x_1 \oplus z_1 \oplus x_3) = 0$, or equivalently $x_1 \oplus z_1 \oplus z_3 = Z$ and $\text{Bits}_C(x_3 \oplus z_3) = 0$. Consider the i^{th} query.

Regular (forward or inverse) query to $\bar{\pi}_1$. Without loss of generality, consider a forward query. Denote the query by (x_1, z_1) . There are at most q choices for (x_3, z_3) . As z_1 is randomly drawn from a set of size at least $2^n - q$, the equation is satisfied with probability at most $\frac{q}{2^n - q}$. Each hit adds at most 1 solution.

Regular forward query to $\bar{\pi}_3$. By $\neg \text{aux}_2(\mathcal{Q}_q)$, there are at most $\tau_2 2^{n-|C|}$ choices for (x_1, z_1) to satisfy $\text{Bits}_C(x_1 \oplus z_1 \oplus x_3) = 0$. As z_3 is randomly drawn from a set of size at least $2^n - q$, the equation is satisfied with probability at most $\frac{\tau_2 2^{n-|C|}}{2^n - q}$. By $\neg \text{aux}_2(\mathcal{Q}_q)$, each hit adds at most τ_2 solutions.

Regular inverse query to $\bar{\pi}_3$. By $\neg \text{aux}_2(\mathcal{Q}_q)$, there are at most τ_2 choices for (x_1, z_1) to satisfy $x_1 \oplus z_1 \oplus z_3 = Z$. As x_3 is randomly drawn from a set of size at least $2^n - q$, the equation is satisfied with probability at most $\frac{\tau_2 2^{n-|C|}}{2^n - q}$. By $\neg \text{aux}_2(\mathcal{Q}_q)$, each hit adds at most τ_2 solutions.

Predicate query to $\bar{\pi}_1$. Denote the query response by (x_1, z_1) , which will satisfy $\text{Bits}_C(x_1 \oplus z_1) = 0$. By $\neg \text{aux}_1(\mathcal{Q}_q)$, there are at most τ_1 choices for (x_3, z_3) to satisfy $\text{Bits}_C(x_3) = 0$ and $\text{Bits}_C(z_3) = \text{Bits}_C(Z)$. By Lem. 2 part (ii), we find that the equation is satisfied with probability at most $\frac{\tau_1 2^{|C|}}{2^n - q}$. Each hit adds at most 1 solution.

Predicate query to $\bar{\pi}_3$. Denote the query response by (x_3, z_3) , which will satisfy $\text{Bits}_C(x_3 \oplus z_3) = 0$. There are at most q choices for (x_1, z_1) . Similar to the proof of Lem. 13, one can prove that the equation is satisfied with probability at most $\frac{2^{n-|C|}}{2^n 2^{n-|C|} - q} \cdot q \leq \frac{q}{2^n - q}$: we need the predicate query to satisfy $\text{Bits}_C(x_3) = \text{Bits}_C(x_1 \oplus z_1)$, $\text{Bits}_C(x_3 \oplus z_3) = 0$, and $z_3 = x_1 \oplus z_1 \oplus Z$, and this gives at most $2^{n-|C|}$ possible tuples of $\bar{\Sigma}_k(P, P^\Phi)$. By $\neg \text{aux}_2(\mathcal{Q}_q)$, each hit adds at most τ_2 solutions.

Conclusion. Thus, if $B = 1$, any query results in a solution with probability at most $\frac{\tau_1 2^{|C|} + \tau_2 2^{n-|C|} + q}{2^n - q}$, and each hit adds at most τ_2 solutions. Consequently, more than τ_5 solu-

tions are obtained with probability at most $\left(\frac{q}{\tau_5 + 1} \right) \left(\frac{\tau_1 2^{|C|} + \tau_2 2^{n-|C|} + q}{2^n - q} \right)^{\frac{\tau_5+1}{\tau_2}} \leq$
 $\left(\frac{e\tau_1\tau_2 2^{|C|}q + e\tau_2^2 2^{n-|C|}q + e\tau_2 q^2}{(\tau_5 + 1)(2^n - q)} \right)^{\frac{\tau_5+1}{\tau_2}}. \quad \square$

From (30-31), (32), and the results of Lems. 11-14 we find:

$$\begin{aligned} \text{Adv}_{\text{FSS}}^{\text{epre}}(q) &\leq \frac{(\tau_4 + \tau_5)[B = 1]2^{|C|}q + \tau_2 B^2 q^2 + \tau_3 Bq + B^2 q(q + 1)}{2^n - Bq} + \frac{[B = 1]2^{|C|}q^3}{2^{2n} - Bq} + \\ &[B = 1] \left(\frac{e2^{n-|C|}q}{(\tau_1 + 1)(2^n - q)} \right)^{\tau_1+1} + 2^{n-|C|} \left(\frac{eB^2 \delta_{B,C}[1]q}{(\tau_2 + 1)(2^n - Bq)} \right)^{\frac{\tau_2+1}{B}} + \\ &2^n \left(\frac{eB^2 q}{(\tau_2 + 1)(2^n - Bq)} \right)^{\frac{\tau_2+1}{B}} + 2^n \left(\frac{e\tau_2 B^3 \delta_{B,C}[1]q^2}{(\tau_3 + 1)(2^n - Bq)} \right)^{\frac{\tau_3+1}{\tau_2 B}} + \\ &[B = 1] \left(\frac{e\tau_1\tau_2 2^{|C|}q + e\tau_2 q^2}{(\tau_4 + 1)(2^n - q)} \right)^{\frac{\tau_4+1}{\tau_2}} + \\ &[B = 1] \left(\frac{e\tau_1\tau_2 2^{|C|}q + e\tau_2^2 2^{n-|C|}q + e\tau_2 q^2}{(\tau_5 + 1)(2^n - q)} \right)^{\frac{\tau_5+1}{\tau_2}}. \end{aligned}$$

The proof of Prop. 10 is completed by making the following simplifications: we use the fact that $2^n - Bq \geq 2^{n-1}$ for $Bq \leq 2^{n-1}$, and put $\tau_2 = \tau B - 1$, $\tau_3 = \tau' \tau_2 B - 1$, $\tau_1 = \tau'' - 1$, and $\tau_4 = \tau_5 = \tau'' \tau_2 - 1$ for some integral $\tau, \tau', \tau'' > 0$. Additionally, we use $q + 1 \leq 2q$ for $q \geq 1$.