

# Circular chosen-ciphertext security with compact ciphertexts

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## Abstract

A key-dependent message (KDM) secure encryption scheme is secure even if an adversary obtains encryptions of messages that depend on the secret key. Such key-dependent encryptions naturally occur in scenarios such as harddisk encryption, formal cryptography, or in specific protocols. However, there are not many provably secure constructions of KDM-secure encryption schemes. Moreover, only one construction, due to Camenisch, Chandran, and Shoup (Eurocrypt 2009) is known to be secure against active (i.e., CCA) attacks.

In this work, we construct the first public-key encryption scheme that is KDM-secure against active adversaries and has compact ciphertexts. As usual, we allow only circular key dependencies, meaning that encryptions of arbitrary *entire* secret keys under arbitrary public keys are considered in a multi-user setting.

Technically, we follow the approach of Boneh, Halevi, Hamburg, and Ostrovsky (Crypto 2008) to KDM security, which however only achieves security against passive adversaries. We explain an inherent problem in adapting their techniques to active security, and resolve this problem using a new technical tool called “lossy algebraic filters” (LAFs). We stress that we significantly deviate from the approach of Camenisch, Chandran, and Shoup to obtain KDM security against active adversaries. This allows us to develop a scheme with compact ciphertexts that consist only of a constant number of group elements.

**Keywords:** key-dependent messages, chosen-ciphertext security, public-key encryption.

## 1 Introduction

**KDM security.** An encryption scheme is key-dependent message (KDM) secure if it is secure even against an adversary who has access to encryptions of messages that depend on the secret key. Such a setting arises, e.g., in harddisk encryption [11], computational soundness results in formal methods [7, 3], or specific protocols [14]. KDM security does not follow from standard security [2, 16], and there are indications [22, 6] that KDM security (at least in its most general form) cannot be proven using standard techniques; it seems that dedicated constructions and proof techniques are necessary.<sup>1</sup>

**The BHHO approach to KDM-CPA security.** Boneh, Halevi, Hamburg, and Ostrovsky [11] (henceforth BHHO) were the first to construct and prove a public-key encryption (PKE) scheme that is KDM secure under chosen-plaintext attacks (KDM-CPA-secure) in the standard model, under the Decisional Diffie-Hellman (DDH) assumption. While they did not prove their scheme secure under messages that *arbitrarily* depend on the secret key, their result encompasses the important case of *circular (CIRC-CPA) security*. Loosely speaking, a PKE scheme is circular secure if it is secure even in a multi-user setting where encryptions of arbitrary secret keys under arbitrary public keys are known. This notion is sufficient for certain applications [14], and can often be extended to stronger forms of KDM security [6, 13]. Inspired by BHHO, KDM-CPA-secure PKE schemes from other computational assumptions followed [5, 12, 27].

Since we will be using a similar approach, we give a high-level intuition of BHHO’s approach. The crucial property of their scheme is that it is *publicly* possible to construct encryptions of the

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<sup>1</sup>We mention, however, that there are semi-generic transformations that *enhance* the KDM security of an already “slightly” KDM-secure scheme [6, 13, 4].

secret key (under the corresponding public key). Thus, encryptions of the secret key itself do not harm the (IND-CPA) security of that scheme. Suitable homomorphic properties of both keys and ciphertexts allow to extend this argument to circular security (for arbitrarily many users/keys), and to affine functions of all keys.

**Why the BHHO approach fails to achieve KDM-CCA security.** When considering an *active* adversary, we require a stronger form of KDM security. Namely, KDM-CCA, resp. CIRC-CCA security requires security against an adversary who has access to key-dependent encryptions *and* a decryption oracle. (Naturally, to avoid a trivial notion, the adversary is not allowed to submit any of those given KDM encryptions to its decryption oracle.) Now if we want to extend BHHO’s KDM-CPA approach to an adversary with a decryption oracle, the following problem arises: since it is publicly possible to construct (fresh) encryptions of the secret key, an adversary can generate such an encryption and then submit it to its decryption oracle, thus obtaining the full secret key. Hence, the very property that BHHO use to prove KDM-CPA security seemingly contradicts chosen-ciphertext security.

**Our technical tool: lossy algebraic filters (LAFs).** Before we describe our approach to KDM-CCA security, let us present the core technical tool we use. Namely, a *lossy algebraic filter* (LAF) is a family of functions, indexed by a public key and a tag. A function from that family takes a vector  $X = (X_i)_{i=1}^n$  as input. Now if the tag is *lossy*, then the output of the function reveals only a linear combination of the  $X_i$ . If the tag is *injective*, however, then so is the function. We require that there are many lossy tags, which however require a special trapdoor to be found. On the other hand, lossy and injective tags are computationally indistinguishable. This concept is very similar to (parameterized) lossy trapdoor functions [29], and in particular to all-but-many lossy trapdoor functions (ABM-LTFs [23]). However, we do not require efficient inversion, but we do require that lossy functions always reveal *the same* linear combination about the input. In particular, evaluating the same input under many lossy tags will still leave the input (partially) undetermined.

We give a construction of LAFs under the Decision Linear (DLIN) assumption in pairing-friendly groups. Similar to ABM-LTFs, lossy tags correspond to suitably blinded signatures. (This in particular allows to release many lossy tags, while still making the generation of a fresh lossy tag hard for an adversary.) However, unlike with ABM-LTFs, functions with lossy tags always release the same information about its input. Our construction has compact tags with  $\mathbf{O}(1)$  group elements, which will be crucial for our KDM-CCA secure encryption scheme.<sup>2</sup>

**Our approach to KDM-CCA security.** We can now describe our solution to the KDM-CCA dilemma explained above. We will start from a hybrid between the BHHO-like PKE schemes of Brakerski and Goldwasser [12], resp. Malkin et al. [27]. This scheme has compact ciphertexts ( $\mathbf{O}(1)$  group elements), and its KDM-CPA security can be proved under the Decisional Composite Residuosity (DCR) assumption. As with the BHHO scheme, the scheme’s KDM-CPA security relies on the fact that encryptions of its secret key can be publicly generated. Essentially, our modification consists of adding a suitable authentication tag to each ciphertext. This authentication tag comprises the (encrypted) image of the plaintext message under an LAF. During decryption, a ciphertext is rejected in case of a wrong authentication tag.

In our security proof, all authentication tags for the key-dependent encryptions the adversary gets are made with respect to lossy filter tags. This means that information-theoretically, little information about the secret key is released (even with many key-dependent encryptions, resp. LAF evaluations). However, any decryption query the adversary makes must refer (by the LAF properties) to an injective tag. Hence, in order to place a valid key-dependent decryption query, the adversary would have to correctly guess the whole secret key (which is hidden).<sup>3</sup>

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<sup>2</sup>The size of the LAF public key depends on the employed signature scheme. In our main construction, we use Waters signatures, which results in very compact tags, but public keys of  $\mathbf{O}(k)$  group elements, where  $k$  is the security parameter. Alternatively, at the end of Section 3.2, we sketch an LAF with constant-size (but larger than in our main construction) tags *and* constant-size public keys.

<sup>3</sup>As a technical complication, we will also have to protect against a re-use of (lossy) authentication tags. This

Thus, in a nutshell, adding a suitable authentication tag allows to leverage the techniques by BHHO, resp. Brakerski and Goldwasser, Malkin et al. to chosen-ciphertext attacks. In particular, we obtain a CIRC-CCA-secure PKE scheme with compact ciphertexts (of  $\mathbf{O}(1)$  group elements). We prove security under the conjunction of the following assumptions: the DCR assumption (in  $\mathbb{Z}_{N^3}^*$ ), the DLIN assumption (in a pairing-friendly group), and the DDH assumption (somewhat curiously, in the subgroup of order  $(P-1)(Q-1)/4$  of  $\mathbb{Z}_{N^3}^*$ , where  $N = PQ$ ).<sup>4</sup>

**Relation to Camenisch et al.’s CIRC-CCA-secure scheme.** Camenisch, Chandran, and Shoup [15] present the only other known CIRC-CCA-secure PKE scheme in the standard model. They also build upon BHHO techniques, but instead use a Naor-Yung-style double encryption technique [28] to achieve chosen-ciphertext security. As an authentication tag, they attach to each ciphertext a non-interactive zero-knowledge proof that *either* the encryption is consistent (in the usual Naor-Yung sense), *or* that they know a signature for the ciphertext. Since they build on the original, DDH-based BHHO scheme, they can use Groth-Sahai proofs [21] to prove consistency. Compared to our scheme, their system is less efficient: they require  $\mathbf{O}(k)$  group elements per ciphertext, and the secret key can only be encrypted bitwise. However, their sole computational assumption to prove circular security is the DLIN assumption in pairing-friendly groups. One interesting thing to point out is their implicit use of a (one-time) signature scheme. Their argument is conceptually not unlike our LAF argument. However, since they can apply a hybrid argument to substitute all key-dependent encryptions with random ciphertexts, they only require one-time signatures. Furthermore, the meaning of “consistent ciphertext” and “proof” in our case is technically very different. (Unlike Camenisch et al., we apply an argument that rests on the *information* that the adversary has at a certain point about the secret key.)

**Note about concurrent work.** In a work concurrent to ours, Galindo, Herranz, and Villar [20] define and instantiate a strong notion of KDM security for identity-based encryption (IBE) schemes. Using the IBE $\rightarrow$ PKE transformation of Boneh, Canetti, Halevi, and Katz [10], they derive a KDM-CCA-secure PKE scheme. Their concrete construction is entropy-based and achieves only a bounded form of KDM security, much like the KDM-secure SKE scheme from [25]. Thus, while their ciphertexts are very compact, they can only tolerate a number of (arbitrary) KDM queries that is linear in the size of the secret key. In particular, it is not clear how to argue that the encryption of a full secret key in their scheme is secure.

## 2 Preliminaries

**Notation.** For  $n \in \mathbb{N}$ , let  $[n] := \{1, \dots, n\}$ . Throughout the paper,  $k \in \mathbb{N}$  denotes the security parameter. For a finite set  $\mathcal{S}$ , we denote by  $s \leftarrow \mathcal{S}$  the process of sampling  $s$  uniformly from  $\mathcal{S}$ . For a probabilistic algorithm  $A$ , we denote  $y \leftarrow A(x; R)$  the process of running  $A$  on input  $x$  and with randomness  $R$ , and assigning  $y$  the result. We write  $y \leftarrow A(x)$  for  $y \leftarrow A(x; R)$  with uniformly chosen  $R$ . If  $A$ ’s running time is polynomial in  $k$ ,  $A$  is called probabilistic polynomial-time (PPT).

**Key-unique SKE schemes.** A secret-key encryption (SKE) scheme  $(E, D)$  consists of two PPT algorithms. Encryption  $E(K, M)$  takes a key  $K$  and a message  $M$ , and outputs a ciphertext  $C$ . Decryption  $D(K, C)$  takes a key  $K$  and a ciphertext  $C$ , and outputs a message  $M$ . For correctness, we want  $D(K, C) = M$  for all  $M$ , all  $K$ , and all  $C \leftarrow E(K, M)$ . We say that  $(E, D)$  is *key-unique* if for every ciphertext  $C$ , there is at most one key  $K$  with  $D(K, C) \neq \perp$ . For instance, ElGamal encryption can be interpreted as a key-unique SKE scheme through  $E(x, M) := (g^x, g^y, g^{xy} \cdot M)$  (and the obvious  $D$ ). This example assumes a publicly known group  $\mathbb{G} = \langle g \rangle$  in which the DDH

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will be achieved by a combination of one-time signatures and 2-universal hash proof systems [18, 26, 24].

<sup>4</sup>Very roughly, we resort to the DDH assumption since we release *partial* information about our secret keys. Whereas the argument of [12, 27] relies on the fact that the secret key  $sk$  is completely hidden modulo a certain  $N$ , where  $\mathbb{Z}_N$  is message space, we cannot avoid to leak some information modulo about  $sk \bmod N$  by releasing LAF images of  $sk$ . However, using a suitable encoding of messages, we *can* argue that  $sk$  is completely hidden modulo the coprime modulus  $(P-1)(Q-1)/4$ , which enables a reduction to the DDH assumption.

assumption holds.<sup>5</sup> If a larger message space (e.g.,  $\{0, 1\}^*$ ) is desired, hybrid encryption techniques (which preserve key-uniqueness) can be employed.

**PKE schemes.** A public-key encryption (PKE) scheme PKE consists of four<sup>6</sup> PPT algorithms (Pars, Gen, Enc, Dec). The parameter generator Pars( $1^k$ ) outputs public parameters  $pp$  such as a group description. Key generation Gen( $pp$ ) outputs a public key  $pk$  and a secret key  $sk$ . Encryption Enc( $pp, pk, M$ ) takes parameters  $pp$ , a public key  $pk$ , and a message  $M$ , and outputs a ciphertext  $C$ . Decryption Dec( $pp, sk, C$ ) takes public parameters  $pp$ , a secret key  $sk$ , and a ciphertext  $C$ , and outputs a message  $M$ . For correctness, we want Dec( $pp, sk, C$ ) =  $M$  for all  $M$ , all  $pp \leftarrow \text{Pars}(1^k)$ , all  $(pk, sk) \leftarrow \text{Gen}(pp)$ , and all  $C \leftarrow \text{Enc}(pp, pk, M)$ .

**IND-CPA security of SKE and PKE schemes.** An SKE scheme (E, D) is IND-CPA secure iff no efficient adversary  $A$  wins the following game with probability non-negligibly away from  $1/2$ . First,  $A$  selects two equal-length messages  $M_0, M_1$ , then gets an encryption  $E(K, M_b)$  (for random  $K$  and  $b \leftarrow \{0, 1\}$ ), and then takes a guess  $b' \in \{0, 1\}$ . During this,  $A$  gets access to an encryption oracle  $E(K, \cdot)$ . We say that  $A$  wins iff  $b = b'$ . For concrete security analyses, let  $\text{Adv}_{(E,D),A}^{\text{ind-cpa}}(k)$  denote the probability that  $A$  wins this game. This definition can be adapted to the PKE setting by initially giving  $A$  the public parameters  $pp$  and the public key  $pk$  instead of access to an encryption oracle.

**(Chameleon) hashing.** A hash function  $H$  is collision-resistant iff the probability  $\text{Adv}_{H,C}^{\text{cr}}(k)$  that  $C$ , upon input  $H$ , finds  $X \neq X'$  with  $H(X) = H(X')$  is negligible for every PPT  $C$ . A chameleon hash function CH is a keyed and randomized hash function in which key generation outputs a keypair  $(Hpk, Htd)$ . Given a preimage  $X$  and randomness  $R_{\text{CH}}$ , the evaluation key  $Hpk$  allows to efficiently evaluate CH, written  $\text{CH}_{Hpk}(X; R_{\text{CH}})$ . We require collision-resistance in the sense that it is infeasible to find  $(X, R_{\text{CH}}) \neq (X', R'_{\text{CH}})$  with  $\text{CH}_{Hpk}(X; R_{\text{CH}}) = \text{CH}_{Hpk}(X'; R'_{\text{CH}})$ . However, the trapdoor  $Htd$  allows to produce collisions, in the following sense: given arbitrary  $X, R_{\text{CH}}, X'$ ,  $Htd$  allows to efficiently find  $R'_{\text{CH}}$  with  $\text{CH}_{Hpk}(X; R_{\text{CH}}) = \text{CH}_{Hpk}(X'; R'_{\text{CH}})$  for the corresponding  $Hpk$ . We require that the distribution of  $R'_{\text{CH}}$  is uniform given only  $Hpk$  and  $X'$ .

**Signature schemes.** A signature scheme Sig consists of three PPT algorithms (SGen, Sig, Ver). Key generation SGen( $1^k$ ) outputs a verification key  $vk$  and a signing key  $sigk$ . The signature algorithm Sig( $sigk, M$ ) takes a signing key  $sigk$  and a message  $M$  and outputs a signature  $\sigma$ . Verification Ver( $vk, M, \sigma$ ) takes a verification key  $vk$ , a message  $M$  and a potential signature  $\sigma$  and outputs a verdict  $b \in \{0, 1\}$ . For correctness, we require that Ver( $vk, M, \sigma$ ) = 1 for all  $M$ , all  $(vk, sigk) \leftarrow \text{SGen}(1^k)$ , and all  $\sigma \leftarrow \text{Sig}(sigk, M)$ .

**(One-time, strong) existential unforgeability.** A signature scheme Sig is existentially unforgeable (EUF-CMA secure) iff no PPT forger  $F$  wins the following game with non-negligible probability. First,  $F$  gets a verification key  $vk$  as well as access to a signature oracle Sig( $sigk, \cdot$ ).  $A$  wins iff it finally outputs a valid signature  $\sigma$  for a fresh message  $M$  that has not yet been queried to Sig( $sigk, \cdot$ ). Let  $\text{Adv}_{\text{Sig},A}^{\text{euf-cma}}(k)$  denote the probability that  $A$  wins this game. Sig is called one-time existentially unforgeable (OT-EUF-CMA secure) iff no PPT forger  $F$  that makes at most one signature query wins the above game with non-negligible probability. Sig is strongly (OT-)EUF-CMA secure iff it is (OT-)EUF-CMA secure as above, but in a game in which an adversary already wins already if it generates a fresh signature for a (perhaps already signed) message. We let  $\text{Adv}_{\text{Sig},A}^{\text{seuf-cma}}(k)$  denote the probability that  $A$  wins this strong EUF-CMA security game.

**DCR assumption.** The Decisional Composite Residuosity (DCR) assumption over a group  $\mathbb{Z}_{N^{s+1}}^*$  (for  $N = PQ$  with primes  $P, Q$ , and  $s \geq 1$ ) states that for every PPT adversary  $A$ ,

$$\text{Adv}_{\mathbb{Z}_{N^{s+1}}^*, A}^{\text{dcr}}(k) := \Pr[A(N, g) = 1] - \Pr[A(N, g \cdot h) = 1],$$

is negligible, where  $g = \tilde{g}^{N^s}$  for uniform  $\tilde{g} \in \mathbb{Z}_{N^{s+1}}^*$  is a uniformly chosen  $N^s$ -th power, and  $h := 1 + N \in \mathbb{Z}_{N^{s+1}}^*$  is a fixed element of order  $N^s$ . Damgård and Jurik [19] have shown that the DCR assumptions over  $\mathbb{Z}_{N^{s+1}}^*$  and  $\mathbb{Z}_{N^{s'+1}}^*$  are equivalent for any  $s, s'$ .

<sup>5</sup>In view of our application,  $\mathbb{G}$  can be part of the public parameters of our KDM-secure PKE scheme.

<sup>6</sup>We will only use public parameters for PKE schemes, but not, e.g., for signature schemes.

**DDH and DLIN assumptions.** The Decisional Diffie-Hellman (DDH), resp. Decision Linear [8] (DLIN) assumptions over a group  $\mathbb{G}$  of (not necessarily prime) order  $q$  state that for every PPT adversary  $A$ , the respective following functions are negligible:

$$\begin{aligned} \text{Adv}_{\mathbb{G},A}^{\text{ddh}}(k) &:= \Pr [A(g, g^x, g^y, g^{xy}) = 1] - \Pr [A(g, g^x, g^y, g^z) = 1], \\ \text{Adv}_{\mathbb{G},A}^{\text{dlin}}(k) &:= \Pr [A(g, U_1, U_2, g^{s_0}, U_1^{s_1}, U_2^{s_0+s_1}) = 1] - \Pr [A(g, U_1, U_2, g^{s_0}, U_1^{s_1}, U_2^{s_2}) = 1], \end{aligned}$$

where  $g$  is a uniform generator of  $\mathbb{G}$ , and  $U_1, U_2 \leftarrow \mathbb{G}$  and  $x, y, z, s_0, s_1, s_2 \leftarrow \mathbb{Z}_q$  are uniform.

**Pairings.** A (symmetric) pairing is a map  $e : \mathbb{G} \times \mathbb{G} \rightarrow \mathbb{G}_T$  between two cyclic groups  $\mathbb{G}$  and  $\mathbb{G}_T$  that satisfies  $e(g, g) \neq 1$  and  $e(g^a, g^b) = e(g, g)^{ab}$  for all generators  $g$  of  $\mathbb{G}$  and all  $a, b \in \mathbb{Z}$ .

**Waters signatures.** In [30], Waters proves the following signature scheme EUF-CMA secure:<sup>7</sup>

- $\text{Gen}(1^k)$  chooses groups  $\mathbb{G}, \mathbb{G}_T$  of prime order  $p$ , along with a pairing  $e : \mathbb{G} \times \mathbb{G} \rightarrow \mathbb{G}_T$ , a generator  $g \in \mathbb{G}$ , and uniform group elements  $g^\omega, H_0, \dots, H_k \in \mathbb{G}$ . Output is

$$vk = (\mathbb{G}, \mathbb{G}_T, e, p, g, (H_i)_{i=0}^k, e(g, g)^\omega), \quad sigk = (vk, g^\omega).$$

- $\text{Sig}(sigk, M)$ , for  $M = (M_i)_{i=1}^k \in \{0, 1\}^k$ , picks  $r \leftarrow \mathbb{Z}_p$ , and lets  $\sigma := (g^r, g^\omega \cdot (H_0 \prod_{i=1}^k H_i^{M_i})^r)$ .
- $\text{Ver}(vk, M, \sigma)$ , for  $\sigma = (\sigma_0, \sigma_1)$ , outputs 1 iff  $e(g, \sigma_1) = e(g, g)^\omega \cdot e(\sigma_0, H_0 \prod_{i=1}^k H_i^{M_i})$ .

**KDM-CCA and CIRC-CCA security.** Let  $n = n(k)$  and let PKE be a PKE scheme with message space  $\mathcal{M}$ . PKE is chosen-ciphertext secure under key-dependent message attacks ( $n$ -KDM-CCA secure) iff

$$\text{Adv}_{\text{PKE},n,A}^{\text{kdm-cca}}(k) := \Pr \left[ \text{Exp}_{\text{PKE},n,A}^{\text{kdm-cca}}(k) = 1 \right] - 1/2$$

is negligible for all PPT  $A$ , where experiment  $\text{Exp}_{\text{PKE},n,A}^{\text{kdm-cca}}$  is defined as follows. First, the experiment tosses a coin  $b \leftarrow \{0, 1\}$ , and samples public parameters  $pp \leftarrow \text{Pars}(1^k)$  and  $n$  keypairs  $(pk_i, sk_i) \leftarrow \text{Gen}(pp)$ . Then  $A$  is invoked with input  $pp$  and  $(pk_i)_{i=1}^n$ , and access to two oracles:

- a KDM oracle  $\mathcal{KDM}_b(\cdot, \cdot)$  that maps  $i \in [n]$  and a function  $f : (\{0, 1\}^*)^n \rightarrow \{0, 1\}^*$  to a ciphertext  $C \leftarrow \text{Enc}(pp, pk_i, M)$ . If  $b = 0$ , then  $M = f((sk_i)_{i=1}^n)$ ; else,  $M = 0^{|f((sk_i)_{i=1}^n)|}$ .
- a decryption oracle  $\mathcal{DEC}(\cdot, \cdot)$  that takes as input an index  $i \in [n]$  and a ciphertext  $C$ , and outputs  $\text{Dec}(pp, sk_i, C)$ .

When  $A$  finally generates an output  $b' \in \{0, 1\}$ , the experiment outputs 1 if  $b = b'$  (and 0 else). We require that (a)  $A$  never inputs a ciphertext  $C$  to  $\mathcal{DEC}$  that has been produced by  $\mathcal{KDM}_b$  (for the same index  $i$ ), and (b)  $A$  only specifies PPT-computable functions  $f$  that always output messages of the same length. As a relevant special case, PKE is  $n$ -CIRC-CCA-secure if it is  $n$ -KDM-CCA secure against all  $A$  that only query  $\mathcal{KDM}_b$  with functions  $f \in \mathcal{F}$  for

$$\mathcal{F} := \{f_j : f_j((sk_i)_{i=1}^n) = sk_j\}_{j \in [n]} \cup \{f_M : f_M((sk_i)_{i=1}^n) = M\}_{M \in \mathcal{M}}.$$

(Technically, what we call “circular security” is called “clique security” in [11]. We stress, however, that our notion of circular security implies that of [11].) Our main result will be a PKE scheme that is  $n$ -CIRC-CCA-secure for all polynomials  $n = n(k)$ .

## 3 Lossy algebraic filters

### 3.1 Definition

**Informal description.** An  $(\ell_{\text{LAF}}, n)$ -lossy algebraic filter (LAF) is a family of functions indexed by a public key  $Fpk$  and a tag  $t$ . A function  $\text{LAF}_{Fpk,t}$  from the family maps an input  $X = (X_i)_{i=1}^n \in \mathbb{Z}_p^n$  to an output  $\text{LAF}_{Fpk,t}(X)$ , where  $p$  is an  $\ell_{\text{LAF}}$ -bit prime contained in the public key.

<sup>7</sup>In fact, our description is a slight folklore optimization of Waters’s scheme. The original scheme features elements  $g^\alpha, g^\beta$  in  $vk$ , so that  $e(g^\alpha, g^\beta)$  takes the role of  $e(g, g)^\omega$ .

The crucial property of an LAF is its lossiness. Namely, for a given public key  $Fpk$ , we distinguish *injective* and *lossy* tags.<sup>8</sup> For an injective tag  $t$ , the function  $\text{LAF}_{Fpk,t}(\cdot)$  is injective, and thus has an image of size  $p^n$ . However, if  $t$  is lossy, then  $\text{LAF}_{Fpk,t}(\cdot)$  only depends on a linear combination  $\sum_{i=1}^n \omega_i X_i \bmod p$  of its input. In particular, different  $X$  with the same value  $\sum_{i=1}^n \omega_i X_i \bmod p$  are mapped to the same image. Here, the coefficients  $\omega_i \in \mathbb{Z}_p$  only depend on  $Fpk$  (but not on  $t$ ). For a lossy tag  $t$ , the image of  $\text{LAF}_{Fpk,t}(\cdot)$  is thus of size at most  $p$ . Note that the modulus  $p$  is public, while the coefficients  $\omega_i$  may be (and in fact will have to be) computationally hidden.

For this concept to be useful, we require that (a) lossy and injective tags are computationally indistinguishable, (b) lossy tags can be generated using a special trapdoor, but (c) new lossy (or, rather, non-injective) tags cannot be found efficiently without that trapdoor, even when having seen polynomially many lossy tags before. In view of our application, we will work with structured tags: each tag  $t = (t_c, t_a)$  consists of a *core tag*  $t_c$  and an *auxiliary tag*  $t_a$ . The auxiliary tag will be a ciphertext part that is authenticated by a filter image.

**Definition 3.1.** An  $(\ell_{\text{LAF}}, n)$ -lossy algebraic filter (LAF)  $\text{LAF}$  consists of three PPT algorithms:

**Key generation.**  $\text{FGen}(1^k)$  samples a keypair  $(Fpk, Ftd)$ . The public key  $Fpk$  contains an  $\ell_{\text{LAF}}$ -bit prime  $p$  and the description of a tag space  $\mathcal{T} = \mathcal{T}_c \times \{0, 1\}^*$  for efficiently samplable  $\mathcal{T}_c$ . A tag  $t = (t_c, t_a)$  consists of a core tag  $t_c \in \mathcal{T}_c$  and an auxiliary tag  $t_a \in \{0, 1\}^*$ . A tag may be injective, or lossy, or neither.  $Ftd$  is a trapdoor that will allow to sample lossy tags.

**Evaluation.**  $\text{FEval}(Fpk, t, X)$ , for a public key  $Fpk$  and a tag  $t = (t_c, t_a) \in \mathcal{T}$ , maps an input  $X = (X_i)_{i=1}^n \in \mathbb{Z}_p^n$  to a unique output  $\text{LAF}_{Fpk,t}(X)$ .

**Lossy tag generation.**  $\text{FTag}(Ftd, t_a)$ , for a trapdoor  $Ftd$  and  $t_a \in \{0, 1\}^*$ , samples a core tag  $t_c$  such that  $t = (t_c, t_a)$  is lossy.

We require the following:

**Lossiness.** The function  $\text{LAF}_{Fpk,t}(\cdot)$  is injective if  $t$  is injective. If  $t$  is lossy, then  $\text{LAF}_{Fpk,t}(X)$  depends only on  $\sum_{i=1}^n \omega_i X_i \bmod p$  for  $\omega_i \in \mathbb{Z}_p$  that only depend on  $Fpk$ .

**Indistinguishability.** Lossy tags are indistinguishable from random tags. Formally,

$$\text{Adv}_{\text{LAF}, A}^{\text{ind}}(k) := \Pr \left[ A(1^k, Fpk)^{\text{FTag}(Ftd, \cdot)} = 1 \right] - \Pr \left[ A(1^k, Fpk)^{\mathcal{O}_{\mathcal{T}_c}(\cdot)} = 1 \right]$$

is negligible for all PPT  $A$ , where  $(Fpk, Ftd) \leftarrow \text{FGen}(1^k)$ , and  $\mathcal{O}_{\mathcal{T}_c}(\cdot)$  is the oracle that ignores its input and samples a random core tag  $t_c$ .

**Evasiveness.** Non-injective (and in particular lossy) tags are hard to find, even given multiple lossy tags:

$$\text{Adv}_{\text{LAF}, A}^{\text{eva}}(k) := \Pr \left[ t \text{ non-injective} \mid t \leftarrow A(1^k, Fpk)^{\text{FTag}(Ftd, \cdot)} \right]$$

is negligible with  $(Fpk, Ftd) \leftarrow \text{FGen}(1^k)$ , and for any PPT algorithm  $A$  that never outputs a tag obtained through oracle queries (i.e.,  $A$  never outputs  $t = (t_c, t_a)$  when  $t_c$  has been obtained by an oracle query  $\text{FTag}(Ftd, t_a)$ ).

## 3.2 Construction

**Intuition.** We present a construction based on the DLIN problem in a group  $\mathbb{G}$  of order  $p$  with symmetric pairing  $e : \mathbb{G} \times \mathbb{G} \rightarrow \mathbb{G}_T$ . Essentially, each tag corresponds to  $n$  DLIN-encrypted Waters signatures. If the signatures are valid, then the tag is lossy. The actual filter maps an input  $X = (X_i)_{i=1}^n \in \mathbb{Z}_p^n$  to the tuple

$$\text{LAF}_{Fpk,t}(X) := \mathbf{M} \circ X := \left( \prod_{j=1}^n \mathbf{M}_{i,j}^{X_j} \right)_{i=1}^n \in \mathbb{G}_T^n, \quad (1)$$

<sup>8</sup>Technically, there may also be tags that are neither injective nor lossy.

where the matrix  $\mathbf{M} = (\mathbf{M}_{i,j})_{i,j \in [n]} \in \mathbb{G}_T^{\mathbf{n} \times \mathbf{n}}$  is computed from public key and tag. Note that this mapping is lossy if and only if the matrix

$$\widetilde{\mathbf{M}} := (\widetilde{\mathbf{M}}_{i,j}) := (\text{dlog}_{e(g,g)}(\mathbf{M}_{i,j}))_{i,j} \in \mathbb{Z}_p^{\mathbf{n} \times \mathbf{n}} \quad (2)$$

of discrete logarithms (to some arbitrary basis  $e(g, g) \in \mathbb{G}_T$ ) is non-invertible.

For a formal description, let  $\ell_{\text{LAF}}(k), \mathbf{n}(k)$  be two functions.

**Key generation.**  $\text{FGen}(1^k)$  generates cyclic groups  $\mathbb{G}, \mathbb{G}_T$  of prime order  $p$  (where  $p$  is of bitlength  $\lceil \log_2(p) \rceil = \ell_{\text{LAF}}(k)$ ), and a symmetric pairing  $e : \mathbb{G} \times \mathbb{G} \rightarrow \mathbb{G}_T$ . Then  $\text{FGen}$  chooses

- a generator  $g \in \mathbb{G}$  and a uniform exponent  $\omega \leftarrow \mathbb{Z}_p$ ,
- uniform group elements  $U_1, \dots, U_n \leftarrow \mathbb{G}, H_0, \dots, H_k \leftarrow \mathbb{G}$ , and
- a keypair  $(\text{Hpk}, \text{Htd})$  for a chameleon hash function  $\text{CH} : \{0, 1\}^* \rightarrow \{0, 1\}^k$ .

$\text{FGen}$  finally outputs

$$\begin{aligned} \text{Fpk} &:= (\mathbb{G}, \mathbb{G}_T, e, p, g, (H_i)_{i=0}^k, (U_i)_{i=1}^n, W := e(g, g)^\omega, \text{Hpk}) \\ \text{Ftd} &:= (\text{Fpk}, g^\omega, \text{Htd}). \end{aligned}$$

For convenience, write  $U_i = g^{u_i}$  for suitable (unknown) exponents  $u_i$ .

**Tags.** (Core) tags are of the form

$$t_c := (R, (\widetilde{S}_i)_{i=1}^n, (S_{i,j})_{i,j=1}^n, R_{\text{CH}}) \in \mathbb{G} \times \mathbb{G} \times \mathbb{G}^{\mathbf{n} \times \mathbf{n}} \times \mathcal{R}_{\text{CH}},$$

where we require  $e(U_{j'}, S_{i,j}) = e(U_j, S_{i,j'})$  whenever  $i \notin \{j, j'\}$ . This means we can write

$$R = g^r, \quad \widetilde{S}_i = g^{\widetilde{s}_i}, \quad S_{i,j} = U_j^{s_i} \quad (i \neq j)$$

for suitable  $r, s_i, \widetilde{s}_i$ . To a tag  $t = (t_c, t_a)$  (with auxiliary part  $t_a \in \{0, 1\}^*$ ), we associate the matrix  $\mathbf{M} = (\mathbf{M}_{i,j})_{i,j=1}^n \in \mathbb{G}_T^{\mathbf{n} \times \mathbf{n}}$  with

$$\begin{aligned} \mathbf{M}_{i,j} &= e(U_j, \widetilde{S}_i) \cdot e(g, S_{i,j}) = e(g, g)^{u_j(\widetilde{s}_i + s_i)} \quad (i \neq j) \\ \mathbf{M}_{i,i} &= \frac{e(g, S_{i,i})}{W \cdot e(H_0 \prod_{i=1}^k H_i^{T_i}, R)} \end{aligned} \quad (3)$$

for  $(T_i)_{i=1}^k := \text{CH}_{\text{Hpk}}(R, (\widetilde{S}_i)_{i=1}^n, (S_{i,j})_{i,j=1}^n, t_a; R_{\text{CH}})$ . If the matrix  $\widetilde{\mathbf{M}}$  of discrete logarithms (see (2)) is invertible, we say that  $t$  is injective; if  $\widetilde{\mathbf{M}}$  has rank 1, then  $t$  is lossy. Note that for lossy tags, thus  $\mathbf{M}_{i,j} = e(g, g)^{u_j(\widetilde{s}_i + s_i)}$  for all  $i, j$ .

**Evaluation.**  $\text{FEval}(\text{Fpk}, t, X)$ , for  $t = (t_c, t_a)$ ,  $t_a \in \{0, 1\}^*$ ,  $X = (X_i)_{i=1}^n \in \mathbb{Z}_p^n$ , and  $\text{Fpk}$  and  $t_c$  as above, computes  $\mathbf{M}$  as in (3) and then  $(Y_i)_{i=1}^n := \text{LAF}_{\text{Fpk}, t}(X) \in \mathbb{G}_T^n$  as in (1).

**Lossiness.** If we write  $Y_i = e(g, g)^{y_i}$ , the definition of  $\text{FEval}$  implies  $(y_i)_{i=1}^n = \widetilde{\mathbf{M}} \cdot X$ . Since injective tags satisfy that  $\widetilde{\mathbf{M}}$  is invertible, they lead to injective functions  $\text{LAF}_{\text{Fpk}, t}(\cdot)$ . On the other hand, for a lossy tag,  $\widetilde{\mathbf{M}}_{i,j} = u_j(\widetilde{s}_i + s_i)$ , so that

$$y_i = \sum_{j=1}^n u_j(\widetilde{s}_i + s_i) X_j = (\widetilde{s}_i + s_i) \cdot \sum_{j=1}^n u_j X_j \quad \text{mod } p.$$

Specifically,  $\text{LAF}_{\text{Fpk}, t}(X)$  depends only on  $\sum_i \omega_i X_i \text{ mod } p$  for  $\omega_i := u_i$ .

**Lossy tag generation.**  $\text{FTag}(\text{Ftd}, t_a)$ , for  $\text{Ftd}$  as above and  $t_a \in \{0, 1\}^*$ , first chooses a random  $\text{CH}$ -image  $T = (T_i)_{i=1}^k \in \{0, 1\}^k$  that can later be explained, using  $\text{Htd}$ , as the  $\text{CH}$ -image of an arbitrary preimage.  $\text{FTag}$  then chooses uniform  $r, s_i, \widetilde{s}_i \leftarrow \mathbb{Z}_p$  and sets

$$R := g^r, \quad \widetilde{S}_i := g^{\widetilde{s}_i}, \quad S_{i,j} := U_j^{s_i} \quad (i \neq j), \quad S_{i,i} := U_i^{\widetilde{s}_i + s_i} \cdot g^\omega \cdot \left( H_0 \prod_{i=1}^k H_i^{T_i} \right)^r. \quad (4)$$

Finally, FTag chooses CH-randomness  $R_{\text{CH}}$  such that  $\text{CH}_{\text{Hpk}}(R, (\tilde{S}_i)_{i=1}^n, (S_{i,j})_{i,j=1}^n, t_a; R_{\text{CH}}) = T$  and outputs  $t_c = (R, (\tilde{S}_i)_{i=1}^n, (S_{i,j})_{i,j=1}^n, R_{\text{CH}})$ . Intuitively,  $t_c$  consists of  $\mathbf{n}$  DLIN encryptions (with correlated randomness  $s_i, \tilde{s}_i$ ) of Waters signatures  $(g^r, g^\omega \cdot (H_0 \prod_{i=1}^k H_i^{T_i})^r)$  for message  $T$ . Indeed, substituting into (3) yields

$$\mathbf{M}_{i,i} := \frac{e(g, g)^{u_i(\tilde{s}_i + s_i)} \cdot W \cdot e(g, (H_0 \prod_{i=1}^k H_i^{T_i})^r)}{W \cdot e(g^r, H_0 \prod_{i=1}^k H_i^{T_i})} = e(g, g)^{u_i(\tilde{s}_i + s_i)}.$$

Hence,  $\tilde{\mathbf{M}}_{i,j} = u_j(\tilde{s}_i + s_i)$  for all  $i, j$ , and thus the resulting tag  $t = (t_c, t_a)$  is lossy.

**A generalization with constant-size evaluation keys.** The LAF above inherits a rather large public key of  $\mathbf{O}(k)$  group elements from Waters signatures. We now sketch how to generalize LAF to any structure-preserving signature scheme; plugging in, e.g., the DLIN-based signature scheme of Abe et al. [1] yields an LAF with constant-size tags and keys. (Compared to LAF, tags will be larger, however.) The idea is to have tags that directly contain a matrix  $\mathbf{M} \in \mathbb{G}_T^{\mathbf{n} \times \mathbf{n}}$  as above, along with a DLIN ciphertext  $C$ , and a Groth-Sahai [21] proof  $\pi$ . The statement proved by  $\pi$  is that either  $\mathbf{M}$  is injective (e.g., in the sense that there exist  $V_1, \dots, V_n$  with  $\mathbf{M}_{i,i} = e(U_i, V_i) \cdot e(g, g)$  and  $\mathbf{M}_{i,j} = e(U_j, V_i)$  for fixed elements  $U_i$  from the public key and all  $i \neq j$ ), or that  $C$  contains a fresh signature (e.g., of a chameleon hash value  $T$  as above). Evaluation of this LAF takes place as in (1). Lossy tags can be generated using a signing key and proving the “or” branch of the statement. The soundness of Groth-Sahai proofs ensures that any adversarially produced lossy tag (with lossy  $\mathbf{M}$ ) would imply a fresh forged signature.

### 3.3 Security proof

**Theorem 3.2.** *If the DLIN assumption holds in  $\mathbb{G}$ , and CH is a chameleon hash function, then the LAF construction LAF from Section 3.2 satisfies Definition 3.1.*

The lossiness of LAF has already been discussed in Section 3.2. We prove indistinguishability and evasiveness separately.

**Lemma 3.3.** *For every adversary  $A$  on LAF’s indistinguishability, there exists a DLIN distinguisher  $B$  such that*

$$\text{Adv}_{\text{LAF}, A}^{\text{ind}}(k) = \mathbf{n} \cdot \text{Adv}_B^{\text{dlin}}(k). \quad (5)$$

Intuitively, to see Lemma 3.3, observe that lossy tags differ from random tags only in their  $S_{i,i}$  components, and in how the CH randomness  $R_{\text{CH}}$  is generated. For lossy tags, the  $S_{i,i}$  are (parts of) DLIN ciphertexts, which are pseudorandom under the DLIN assumption. Furthermore, the uniformity property of CH guarantees that the distribution of  $R_{\text{CH}}$  is the same for lossy and random tags.

*Proof.* Assume a PPT adversary  $A$ . We proceed in games. In Game  $i$ ,  $A$  gets an input  $\text{Fpk}$  and interacts with an oracle  $\mathcal{O}_i$ . Let  $\text{out}_i$  denote  $A$ ’s output in Game  $i$ .

In **Game 1**, we let  $\mathcal{O}_1(\cdot) := \text{FTag}(\text{Ftd}, \cdot)$ , where  $\text{Ftd}$  is the trapdoor initially sampled alongside  $\text{Fpk}$ . Thus,  $\mathcal{O}_1(t_a)$  outputs core tags  $t_c = (R, (\tilde{S}_i)_{i=1}^n, (S_{i,j})_{i,j=1}^n, R_{\text{CH}})$  generated as in (4).

In **Game 2.  $i^*$**  (for  $0 \leq i^* \leq \mathbf{n}$ ), we let  $\mathcal{O}_2$  generate core tags as in Game 1, but with independently and uniformly chosen  $S_{i,i} \in \mathbb{G}$  for  $i \leq i^*$ . Note that Game 2.0 is equivalent to Game 1. Let furthermore Game 2 be defined as Game 2. $\mathbf{n}$ . We claim

$$\Pr[\text{out}_1 = 1] - \Pr[\text{out}_2 = 1] = \Pr[\text{out}_{2.0} = 1] - \Pr[\text{out}_{2.\mathbf{n}} = 1] = \mathbf{n} \cdot \text{Adv}_B^{\text{dlin}}(k) \quad (6)$$

for a suitable DLIN distinguisher  $B$ . Namely,  $B$  uniformly chooses  $i^* \in [\mathbf{n}]$  and parses its DLIN challenge as  $(g, U, U_{i^*}, g^{\tilde{s}_{i^*}}, U^{s_{i^*}}, C)$ , where  $C = U_{i^*}^{\tilde{s}_{i^*} + s_{i^*}}$  or  $C \in \mathbb{G}$  is uniform.  $B$  then first re-randomizes its input to obtain many tuples  $(g^{\tilde{s}_{i^*, \ell}}, U^{s_{i^*, \ell}}, C_\ell)$ , where (a) the  $\tilde{s}_{i^*, \ell}, s_{i^*, \ell}$  are independently and uniformly random, and (b)  $C_\ell = U_{i^*}^{\tilde{s}_{i^*, \ell} + s_{i^*, \ell}}$  iff  $X = U_{i^*}^{\tilde{s}_{i^*} + s_{i^*}}$  (otherwise, all



$C_\ell$  are independently and uniformly random). Next,  $B$  simulates Game 2. $(i^* - 1)$  or Game 2. $i^*$ , depending on its own challenge  $C$ . Concretely, to prepare a key  $Fpk$  for  $A$ ,  $B$  sets  $U_j = U^{\alpha_j}$  for all  $j \neq i^*$  and uniform  $\alpha_j \leftarrow \mathbb{Z}_p$ . (Like Game 2. $i^*$ ,  $B$  chooses  $\omega \leftarrow \mathbb{Z}_p$  and a CH keypair  $(Hpk, Htd)$  on its own.) When answering  $A$ 's  $\ell$ -th oracle query,  $B$  proceeds as in Game 2. $i^*$ , but sets up (a)  $\widetilde{S}_{i^*} = g^{s_{i^*,\ell}}$ , (b)  $S_{i,i}$  as in Game 1 for  $i > i^*$ , (c)  $S_{i,i} \leftarrow \mathbb{G}$  uniformly (as in Game 2) for  $i < i^*$ , (d)  $S_{i^*,j} = (U^{s_{i^*,\ell}})^{\alpha_j} = U_j^{s_{i^*,\ell}}$  for  $j \neq i^*$ , (e)  $S_{i^*,i^*} = C_\ell \cdot g^\omega \cdot \left( H_0 \prod_{i=1}^k H_i^{T_i} \right)^r$ . This implicitly sets  $s_{i^*} = s_{i^*,\ell}$  and  $\widetilde{s}_{i^*} = \widetilde{s}_{i^*,\ell}$ . (All other  $s_i, \widetilde{s}_i$  are chosen by  $B$ .) Furthermore, if  $C = U_{i^*}^{\widetilde{s}_{i^*} + s_{i^*}}$ , this setting of  $S_{i,i^*}$  yields Game 2. $(i^* - 1)$ ; but if  $C$  is uniform, then all  $C_i$  are independently uniform, and we obtain Game 2. $i^*$ . We get (6).

In **Game 3**, we choose the hash values  $R_{\text{CH}}$  in the core tags output by  $\mathcal{O}_3$  uniformly and independently. Recall that up to Game 2,  $R_{\text{CH}}$  was instead chosen as follows: first choose a random CH-output  $T$ , and later select  $R_{\text{CH}}$  such that  $\text{CH}_{Hpk}(R, (\widetilde{S}_i)_{i=1}^n, (S_{i,j})_{i,j=1}^n; R_{\text{CH}}) = T$  holds. By definition of chameleon hashing, this induces a uniform distribution of  $R_{\text{CH}}$ . Moreover,  $T$  is not used in Game 2 or Game 3. Hence, the change in Game 3 is merely conceptual, and we obtain

$$\Pr[\text{out}_3 = 1] = \Pr[\text{out}_2 = 1].$$

Now note that in Game 3, the tags  $t_c$  output by  $\mathcal{O}_3$  are random tags. Taking things together, (5) follows as desired.  $\square$

**Lemma 3.4.** *For every adversary  $A$  on LAF's evasiveness, there exist adversaries  $B$ ,  $C$ , and  $F$  such that*

$$\text{Adv}_{\text{LAF},A}^{\text{eva}}(k) \leq \left| \text{Adv}_{\text{LAF},B}^{\text{ind}}(k) \right| + \text{Adv}_{\text{CH},C}^{\text{cr}}(k) + \text{Adv}_{\text{Sig}_{\text{Wat}},F}^{\text{euf-cma}}(k). \quad (7)$$

Intuitively, Lemma 3.4 holds because lossy (or, rather, non-injective) tags correspond to DLIN-encrypted Waters signatures. Hence, even after seeing many lossy tags (i.e., encrypted signatures), an adversary cannot produce a fresh encrypted signature. We note that the original Waters signatures from [30] are re-randomizable and thus not *strongly* unforgeable. To achieve evasiveness, we have thus used a chameleon hash function, much like Boneh et al. [9] did to make Waters signatures strongly unforgeable.

*Proof.* Assume a PPT adversary  $A$ . Again, we proceed in games. Let  $\text{bad}_i$  denote the event that  $A$ 's output in Game  $i$  is a fresh non-injective tag. In **Game 1**,  $A$  gets input  $Fpk$  and interacts with an  $\text{FTag}(Ftd, \cdot)$  oracle. By definition,

$$\Pr[\text{bad}_1] = \text{Adv}_{\text{LAF},A}^{\text{eva}}(k).$$

To describe **Game 2**, denote  $A$ 's output by  $t^* = (t_c^*, t_a^*)$ , for  $t_c^* = (R^*, (\widetilde{S}_i)_{i=1}^n, (S_{i,j}^*)_{i,j=1}^n; R_{\text{CH}}^*)$ . Denote by  $\text{bad}_{\text{coll}}$  the event that  $t^*$  induces a CH-collision in the sense that

$$T^* = \text{CH}_{Hpk}(R^*, (\widetilde{S}_i^*)_{i=1}^n, (S_{i,j}^*)_{i,j=1}^n; R_{\text{CH}}^*) = \text{CH}_{Hpk}(R, (\widetilde{S}_i)_{i=1}^n, (S_{i,j})_{i,j=1}^n; R_{\text{CH}}) = T$$

for some hash value  $T$  associated with an  $\text{FTag}$ -output  $t_c = (R, (\widetilde{S}_i)_{i=1}^n, (S_{i,j})_{i,j=1}^n; R_{\text{CH}})$  (and the corresponding query  $t_a$ ). In Game 2, we abort (and do not raise event  $\text{bad}_2$ ) if  $\text{bad}_{\text{coll}}$  occurs. Intuitively, we would expect to use CH's collision resistance directly to argue that  $\text{bad}_{\text{coll}}$  occurs only negligibly often. However, both in Game 1 and Game 2, we use CH's trapdoor  $Htd$  to construct lossy tags for  $A$ .

Hence, we first argue that  $\text{bad}_{\text{coll}}$  occurs with essentially the same probability in a modified **Game 1'**, in which  $A$  gets random tags instead of lossy tags as oracle answers. Indeed, since lossy and random tags are indistinguishable by Lemma 3.3, and  $\text{bad}_{\text{coll}}$  is efficiently recognizable from  $A$ 's view, we obtain

$$\Pr[\text{bad}_{\text{coll}} \text{ in Game 1'}] - \Pr[\text{bad}_{\text{coll}} \text{ in Game 1}] = \text{Adv}_{\text{LAF},B}^{\text{ind}}(k)$$

for a suitable adversary  $B$  on LAF's indistinguishability. Furthermore, since in Game 1', the CH-trapdoor  $Htd$  is not required, we have

$$\Pr [\text{bad}_{\text{coll}} \text{ in Game } 1'] = \text{Adv}_{\text{CH},C}^{\text{cr}}(k)$$

for a suitable collision-finder  $C$ . However, Game 1 and Game 2 only differ when  $\text{bad}_{\text{coll}}$  occurs, and so we finally get

$$|\Pr [\text{bad}_2] - \Pr [\text{bad}_1]| \leq \Pr [\text{bad}_{\text{coll}} \text{ in Game } 1] \leq \left| \text{Adv}_{\text{LAF},B}^{\text{ind}}(k) \right| + \text{Adv}_{\text{CH},C}^{\text{cr}}(k).$$

**The final reduction.** Now that CH-collisions are excluded, we can finally conclude that any occurrence of  $\text{bad}_2$  means that  $A$  has forged a Waters signature. Concretely, we show that

$$\Pr [\text{bad}_2] = \text{Adv}_{\text{Sig}_{\text{Wat}},F}^{\text{uf-cma}}(k) \quad (8)$$

for a suitable forger  $F$  that attacks  $\text{Sig}_{\text{Wat}}$  and internally simulates Game 2 with  $A$ . Namely,  $F$  gets as input a  $\text{Sig}_{\text{Wat}}$  public key  $(\mathbb{G}, \mathbb{G}_T, e, p, g, (H_i)_{i=0}^k, W := e(g, g)^\omega)$ .  $F$  extends this public key to an LAF public key  $Fpk$  by picking  $U_i = g^{u_i}$  and  $Hpk$ . (In particular,  $F$  knows all  $u_i$  and  $Htd$ .) Upon an FTag-query from  $A$ ,  $F$  constructs elements  $\tilde{S}_i$  and  $S_{i,j}$  (for  $i \neq j$ ) exactly as in (4); note, however, that  $F$  cannot directly compute the  $S_{i,i}$ , since  $F$  does not know  $g^\omega$ . Instead,  $F$  requests a  $\text{Sig}_{\text{Wat}}$  signature for the message  $T \in \{0, 1\}^k$  (as derived in (4)). Such a signature is of the form

$$(g^r, g^\omega \cdot \left( H_0 \prod_{i=1}^k H_i^{T_i} \right)^r),$$

from which  $F$  can compute the elements  $R$  and  $S_{i,i}$  as in (4). Since  $F$  also knows the CH-trapdoor  $Htd$ , this allows to construct lossy tags exactly as FTag would do in Game 2.

It remains to describe how  $F$  extracts a  $\text{Sig}_{\text{Wat}}$ -signature out of a lossy tag  $t = (t_c, t_a)$  that  $A$  finally outputs. By our definition of tags, we may assume that  $t_c = (R, (\tilde{S}_i)_{i=1}^n, (S_{i,j})_{i,j=1}^n, R_{\text{CH}})$  is of the form  $R = g^r$ ,  $\tilde{S}_i = g^{\tilde{s}_i}$ , and  $S_{i,j} = U_j^{s_i}$  for suitable  $r, s_i, \tilde{s}_i$  and all  $i \neq j$ . Furthermore, since  $t_c$  is lossy,

$$\text{rank}(\tilde{\mathbf{M}}) < \mathbf{n} \implies \exists i : \tilde{\mathbf{M}}_{i,i} = u_i(\tilde{s}_i + s_i) \implies \exists i : S_{i,i} = U_i^{\tilde{s}_i + s_i} \cdot g^\omega \cdot \left( H_0 \prod_{i=1}^k H_i^{T_i} \right)^r. \quad (9)$$

Since  $F$  knows all  $u_i$ , it can compute

$$\sigma_i := \frac{S_{i,i}}{\tilde{S}_i^{u_i} \cdot S_{i,j}^{u_i/u_j}} = \frac{S_{i,i}}{U_i^{\tilde{s}_i + s_i}}$$

for all  $i$  (and some  $j \neq i$ ). By (9), for some  $i$ , the pair  $(R, \sigma_i)$  forms a valid  $\text{Sig}_{\text{Wat}}$  signature for  $T = \text{CH}_{Hpk}(R, (\tilde{S}_i)_{i=1}^n, (S_{i,j})_{i,j=1}^n; R_{\text{CH}})$ . Because Game 2 aborts in case of a CH-collision, we may further assume that  $T$  is a message for which  $F$  has not yet requested a signature. Consequently,  $F$  can output a forged signature for a fresh message whenever  $\text{bad}_2$  occurs. This yields (8). Putting things together finally gives (7).  $\square$

Combining Lemma 3.3, Lemma 3.4, and the fact that Waters signatures are EUF-CMA secure already under the CDH assumption, we obtain Theorem 3.2.

## 4 CIRC-CCA-secure encryption scheme

### 4.1 The scheme

**Setting and ingredients.** First, we assume an algorithm  $\text{GenN}$  that outputs  $\ell_N$ -bit Blum integers  $N = PQ$  along with their prime factors  $P$  and  $Q$ . If  $N$  is clear from the context, we write

$\mathbb{G}_{\text{rnd}}$  and  $\mathbb{G}_{\text{msg}}$  for the unique subgroups of  $\mathbb{Z}_{N^3}^*$  of order  $(P-1)(Q-1)/4$ , resp.  $N^2$ . We also write  $h := 1 + N \bmod N^3$ , so  $\langle h \rangle = \mathbb{G}_{\text{msg}}$ . Note that it is efficiently possible to compute  $\text{dlog}_h(X) := x$  for  $X := h^x \in \mathbb{G}_{\text{msg}}$  and  $x \in \mathbb{Z}_{N^2}$ . Specifically, it is efficiently possible to test for membership in  $\mathbb{G}_{\text{msg}}$ . In our scheme,  $\mathbb{G}_{\text{msg}}$  will be used to embed a suitably encoded message, and  $\mathbb{G}_{\text{rnd}}$  will be used for blinding. We require that

- $P$  and  $Q$  are safe primes of bitlength between  $\ell_N/2 - k$  and  $\ell_N/2 + k$ ,
- $\gcd((P-1)(Q-1)/4, N) = 1$  (which holds, e.g., for uniform  $P, Q$  of a certain length),
- $\ell_N \geq 25k + 8$  (e.g.,  $k = 80$  and  $\ell_N = 2048$ )<sup>9</sup>
- the DCR assumption holds in  $\mathbb{Z}_{N^3}^*$ , and the DDH assumption holds in  $\mathbb{G}_{\text{rnd}}$ .

We also assume an  $(\ell_{\text{LAF}}, \mathfrak{n})$ -lossy algebraic filter LAF for  $\mathfrak{n} = 6$  and  $\ell_{\text{LAF}} = (\ell_N + k + 1)/(\mathfrak{n} - 2)$ . Our scheme will encrypt messages from the domain

$$\mathcal{M} := \mathbb{Z}_{2^{3k}} \times \mathbb{Z}_{p \cdot 2^k} \times \mathbb{Z}_{N \cdot 2^{k-2}},$$

where  $p$  is the modulus of the used LAF. (The reason for this weird-looking message space will become clearer in the proof.) During encryption, we will have to treat a message  $M = (a, b, c) \in \mathcal{M}$  both as an element of  $\mathbb{Z}_{N^2}$  and as an LAF-input from  $\mathbb{Z}_p^n$ . In these cases, we can encode

$$[M]_{\mathbb{Z}} := a + 2^{3k} \cdot b + p \cdot 2^{4k} \cdot c \in \mathbb{Z}, \quad [M]_{\mathbb{Z}_p^n} := (a, b \bmod p, c_0, \dots, c_{n-3}) \in \mathbb{Z}_p^n \quad (10)$$

for the natural interpretation of  $\mathbb{Z}_i$ -elements as integers between 0 and  $i - 1$ , and  $c$ 's  $p$ -adic representation  $(c_i)_{i=0}^{n-3} \in \mathbb{Z}_p^{n-2}$  with  $c = \sum_{i=0}^{n-3} c_i \cdot p^i$ . By our choice of  $\ell_N$  and  $\ell_{\text{LAF}}$ , we have  $0 \leq [M]_{\mathbb{Z}} < N^2 - 2^k$ . However, the encoding  $[M]_{\mathbb{Z}_p^n}$  is *not* injective, since it only depends on  $b \bmod p$  (while  $0 \leq b < p \cdot 2^k$ ).

Finally, we assume a strongly OT-EUF-CMA secure signature scheme  $\text{Sig} = (\text{SGen}, \text{Sig}, \text{Ver})$  with  $k$ -bit verification keys, and a key-unique IND-CPA secure symmetric encryption scheme  $(\text{E}, \text{D})$  (see Section 2) with  $k$ -bit symmetric keys  $K$  and message space  $\{0, 1\}^*$ .

Now consider the following PKE scheme PKE:

**Public parameters.**  $\text{Pars}(1^k)$  first runs  $(N, P, Q) \leftarrow \text{GenN}(1^k)$ . Recall that this fixes the groups  $\mathbb{G}_{\text{rnd}}$  and  $\mathbb{G}_{\text{msg}}$ . Then,  $\text{Pars}$  selects two generators  $g_1, g_2$  of  $\mathbb{G}_{\text{rnd}}$ . Finally,  $\text{Pars}$  runs  $(\text{Fpk}, \text{Ftd}) \leftarrow \text{FGen}(1^k)$  and outputs

$$pp = (N, g_1, g_2, \text{Fpk}).$$

In the following, we denote with  $p$  the LAF modulus contained in  $\text{Fpk}$ .

**Key generation.**  $\text{Gen}(pp)$  uniformly selects four messages  $s_j = (a_j, b_j, c_j) \in \mathcal{M}$  (for  $1 \leq j \leq 4$ ) as secret key, and sets

$$pk := \left( u := g_1^{[s_1]_{\mathbb{Z}}} g_2^{[s_2]_{\mathbb{Z}}}, v := g_1^{[s_3]_{\mathbb{Z}}} g_2^{[s_4]_{\mathbb{Z}}} \right), \quad sk := (s_j)_{j=1}^4.$$

**Encryption.**  $\text{Enc}(pp, pk, M)$ , for  $pp$  and  $pk$  as above, and  $M \in \mathcal{M}$ , uniformly selects exponents  $r, \tilde{r} \leftarrow \mathbb{Z}_{N/4}$ , a random filter core tag  $t_c$ , a  $\text{Sig}$ -keypair  $(vk, \text{sigk}) \leftarrow \text{SGen}(1^k)$ , and a random symmetric key  $K \in \{0, 1\}^k$  for  $(\text{E}, \text{D})$ , and computes

$$\begin{aligned} (G_1, G_2) &:= (g_1^r, g_2^r) & Z &:= (u^{vk} v)^{r \cdot N^2} \\ (\tilde{G}_1, \tilde{G}_2) &:= (g_1^{\tilde{r}}, g_2^{\tilde{r}}) & \tilde{Z} &:= (u^{vk} v)^r \cdot u^{\tilde{r}} \cdot h^{K+2k \cdot [M]_{\mathbb{Z}}}, \\ C_E &\leftarrow \text{E}(K, \text{LAF}_{\text{Fpk}, t}([M]_{\mathbb{Z}_p^n})), & \sigma &\leftarrow \text{Sig}(\text{sigk}, ((G_j, \tilde{G}_j)_{j=1}^2, Z, \tilde{Z}, C_E, t_c)) \\ C &:= ((G_j, \tilde{G}_j)_{j=1}^2, Z, \tilde{Z}, C_E, t_c, vk, \sigma) \end{aligned}$$

for the auxiliary tag  $t_a := vk$ , and the resulting filter tag  $t := (t_c, t_a)$ .

<sup>9</sup>Depending on the parameter  $\mathfrak{n}$  below, shorter  $\ell_N$  are possible. The relevant inequality that must hold is (17).

**Decryption.**  $\text{Dec}(pp, sk, C)$ , for  $pp$ ,  $sk$  and  $C$  as above, first checks the signature  $\sigma$  and rejects with  $\perp$  if  $\text{Ver}(vk, ((G_j, \tilde{G}_j)_{j=1}^2, Z, \tilde{Z}, C_E, t_c), \sigma) = 0$ , or if

$$Z \neq \left( G_1^{[s_1]_{\mathbb{Z}} \cdot vk + [s_3]_{\mathbb{Z}}} G_2^{[s_2]_{\mathbb{Z}} \cdot vk + [s_4]_{\mathbb{Z}}} \right)^{N^2}.$$

Then  $\text{Dec}$  computes

$$Z' := G_1^{[s_1]_{\mathbb{Z}} \cdot vk + [s_3]_{\mathbb{Z}}} G_2^{[s_2]_{\mathbb{Z}} \cdot vk + [s_4]_{\mathbb{Z}}} \tilde{G}_1^{[s_1]_{\mathbb{Z}}} \tilde{G}_2^{[s_2]_{\mathbb{Z}}}$$

and then  $K \in \{0, 1\}^k$ ,  $M \in \mathcal{M}$  with

$$K + 2^k \cdot [M]_{\mathbb{Z}} := \text{dlog}_h(\tilde{Z}/Z').$$

If  $\tilde{Z}/Z' \notin \mathbb{G}_{\text{msg}}$ , or no such  $M$  exists, or  $D(K, C_E) \neq \text{LAF}_{Fpk, t}([M]_{\mathbb{Z}_p^n})$  (for  $t = (t_c, t_a)$  computed from  $C$  as during encryption), then  $\text{Dec}$  rejects with  $\perp$ . Else,  $\text{Dec}$  outputs  $M$ .

**Secret keys as messages.** Our scheme has secret keys  $s = (s_j)_{j=1}^4 \in \mathcal{M}^4$ ; hence, we can only encrypt one quarter  $s_j$  of a secret key at a time. In the security proof below, we will thus only consider KDM queries that ask to encrypt a specific secret key *part*. Alternatively, we can change our scheme, so that 4-tuples of  $\mathcal{M}$ -elements are encrypted. To avoid malleability (which would destroy CCA security), we of course have to use only one LAF tag for this. Our CIRC-CCA proof below applies to such a changed scheme with minor syntactic changes.

**Efficiency.** When instantiated with our DLIN-based LAF construction from Section 3, and taking  $n = 6$  as above, our scheme has ciphertexts with 43  $\mathbb{G}$ -elements, 6  $\mathbb{Z}_{N^3}$ -elements, plus chameleon hash randomness, a one-time signature and verification key, and a symmetric ciphertext (whose size could be in the range of one  $\mathbb{Z}_{N^2}$ -element plus some encryption randomness). The number of group elements in the ciphertext is constant, and does not grow in the security parameter. The public parameters contain  $\mathcal{O}(k)$  group elements<sup>10</sup> (most of them from  $\mathbb{G}$ ), and public keys contain two  $\mathbb{Z}_{N^3}$ -elements; secret keys consist of four  $\mathbb{Z}_{N^2}$ -elements. While these parameters are not competitive with current non-KDM-secure schemes, they are significantly better than those from the circular-secure scheme of Camenisch et al. [15].

## 4.2 Security proof (1-user case)

It is instructive to first consider the one-user case. In this case, we essentially only require that PKE is IND-CCA secure, even if encryptions of its secret key are made public.

**Theorem 4.1.** *Assume the DCR assumption holds in  $\mathbb{Z}_{N^3}$ , the DDH assumption holds in  $\mathbb{G}_{\text{rnd}}$ , LAF is an LAF, Sig is a strongly OT-EUF-CMA secure signature scheme, H is collision-resistant, and  $(E, D)$  is a key-unique IND-CPA secure SKE scheme. Then PKE is 1-CIRC-CCA-secure.*

*Proof.* Assume a PPT adversary  $A$  on PKE's 1-CIRC-CCA security. Say that  $A$  always makes  $q = q(k)$  KDM queries. We proceed in games. Let  $out_i$  denote the output of Game  $i$ .

**Game 1** is the 1-KDM-CCA experiment with PKE and  $A$ . Thus, by definition,

$$\Pr[out_1 = 1] - 1/2 = \text{Adv}_{\text{PKE}, A}^{\text{kdm-cca}}(k).$$

In **Game 2**, we modify the way KDM queries are answered. Namely, in each ciphertext prepared for  $A$ , we set up  $Z$  and  $\tilde{Z}$  up as

$$\begin{aligned} Z &:= \left( G_1^{[s_1]_{\mathbb{Z}} \cdot vk + [s_3]_{\mathbb{Z}}} G_2^{[s_2]_{\mathbb{Z}} \cdot vk + [s_4]_{\mathbb{Z}}} \right)^{N^2} \\ \tilde{Z} &:= G_1^{[s_1]_{\mathbb{Z}} \cdot vk + [s_3]_{\mathbb{Z}}} G_2^{[s_2]_{\mathbb{Z}} \cdot vk + [s_4]_{\mathbb{Z}}} \tilde{G}_1^{[s_1]_{\mathbb{Z}}} \tilde{G}_2^{[s_2]_{\mathbb{Z}}} \cdot h^{K + 2^k \cdot [M]_{\mathbb{Z}}}. \end{aligned} \tag{11}$$

<sup>10</sup>Using the generalized LAF sketched at the end of Section 3.2, public parameters with  $\mathcal{O}(1)$  group elements are possible, at the cost of a (constant) number of extra group elements per tag.

for the already prepared  $(G_j, \tilde{G}_j) = (g_j^r, g_j^{\tilde{r}})$ . This change is only conceptual by our setup of  $u, v$ , so

$$\Pr[out_2 = 1] = \Pr[out_1 = 1].$$

In **Game 3**, we again change how KDM ciphertexts are prepared. Intuitively, our goal is now to prepare the  $G_j$  and  $\tilde{G}_j$  with additional  $\mathbb{G}_{\text{msg}}$ -components, such that  $\tilde{Z}$ , as computed in (11), is of the form  $g \cdot h^K$  for some  $g \in \mathbb{G}_{\text{rnd}}$ . (That is, we want the  $\mathbb{G}_{\text{msg}}$ -components of the  $G_j, \tilde{G}_j$  to cancel out the  $h^{2^k \cdot [M]_{\mathbb{Z}}}$  term in (11).) To do so, we prepare

$$G_j = g_j^r / h^{\alpha_j \cdot 2^k} \quad \tilde{G}_j = g_j^{\tilde{r}} / h^{\tilde{\alpha}_j \cdot 2^k}$$

for  $j \in \{1, 2\}$  and suitable  $\alpha_j, \tilde{\alpha}_j$  to be determined.  $\tilde{Z}$  is still computed as in (11), so we have

$$\tilde{Z} = g \cdot h^{K + 2^k \cdot [M]_{\mathbb{Z}} - 2^k (\alpha_1([s_1]_{\mathbb{Z}} \cdot vk + [s_3]_{\mathbb{Z}}) + \alpha_2([s_2]_{\mathbb{Z}} \cdot vk + [s_4]_{\mathbb{Z}}) + \tilde{\alpha}_1[s_1]_{\mathbb{Z}} + \tilde{\alpha}_2[s_2]_{\mathbb{Z}})}$$

for

$$g = g_1^{r \cdot ([s_1]_{\mathbb{Z}} \cdot vk + [s_3]_{\mathbb{Z}}) + \tilde{r}[s_1]_{\mathbb{Z}}} g_2^{r \cdot ([s_2]_{\mathbb{Z}} \cdot vk + [s_4]_{\mathbb{Z}}) + \tilde{r}[s_2]_{\mathbb{Z}}} = \left( u^{vk} v \right)^r u^{\tilde{r}} \in \mathbb{G}_{\text{rnd}}.$$

Thus, to prepare a KDM encryption of  $s_{j^*}$  with a  $\tilde{Z}$  of the form  $\tilde{Z} = g \cdot h^K$ , we can set

$$(\alpha_1, \alpha_2, \tilde{\alpha}_1, \tilde{\alpha}_2) := \begin{cases} (0, 0, 1, 0) & \text{for } j^* = 1 \\ (0, 0, 0, 1) & \text{for } j^* = 2 \\ (1, 0, -vk, 0) & \text{for } j^* = 3 \\ (0, 1, -vk) & \text{for } j^* = 4. \end{cases}$$

(Note that  $vk$  can be chosen independently in advance.) The remaining parts of  $C$  are prepared as in Game 2. We claim that

$$\Pr[out_3 = 1] - \Pr[out_2 = 1] \leq 4 \cdot \text{Adv}_{\mathbb{Z}_{N^3}^{\text{dcr}}, B}^{\text{dcr}}(k) + \mathbf{O}(2^{-k}) \quad (12)$$

for a suitable DCR distinguisher  $B$  that simulates Game 2, resp. Game 3. Concretely,  $B$  gets as input a value  $\tilde{W} \in \mathbb{Z}_{N^3}^*$  of the form  $\tilde{W} = \tilde{g}^{N^2} \cdot h^b$  for  $b \in \{0, 1\}$ . Note that if we set  $W := \tilde{W}^{-2^k}$ , we have  $W = \tilde{g}^{\tilde{r}} / h^{b \cdot 2^k} \in \mathbb{Z}_{N^3}^*$ , with uniform  $\tilde{g}^{\tilde{r}} \in \mathbb{G}_{\text{rnd}}$ . First,  $B$  guesses a value of  $j^* \in [4]$ . (This gives a very small hybrid argument, in which in the  $j^*$ -th step, only encryptions of  $s_{j^*}$  are changed.) We only detail  $B$ 's behavior for the case  $j^* = 3$ ; the other cases are easier or analogous. First,  $B$  sets up  $g_1 := W^{N^2}$  and  $g_2 := W^{\gamma N^4}$  for uniform  $\gamma \in \mathbb{Z}_{N/4}$ . To prepare an encryption of  $s_3$ ,  $B$  chooses uniform  $\rho, \tilde{\rho} \in \mathbb{Z}_{N^2/4}$  and sets

$$G_1 := W^{\rho \cdot (\rho^{-1})} \quad G_2 := W^{\gamma \cdot \rho \cdot (\rho^{-1}) \cdot N^2} \quad \tilde{G}_1 := W^{vk \cdot \tilde{\rho} \cdot (\tilde{\rho}^{-1})} \quad \tilde{G}_2 := W^{\gamma \cdot vk \cdot \tilde{\rho} \cdot (\tilde{\rho}^{-1}) \cdot N^2},$$

where the values  $\rho^{-1}, \tilde{\rho}^{-1}$  are computed modulo  $N^2$ . This implicitly sets  $r = \rho \cdot (\rho^{-1}) / N^2 \bmod |\mathbb{G}_{\text{rnd}}|$  and  $\tilde{r} = vk \cdot \tilde{\rho} \cdot (\tilde{\rho}^{-1}) / N^2 \bmod |\mathbb{G}_{\text{rnd}}|$ , both of which are statistically close to uniform. Furthermore,  $G_j = g_j^r / h^{b \cdot \alpha_j \cdot 2^k}$  and  $\tilde{G}_j = g_j^{\tilde{r}} / h^{b \cdot \tilde{\alpha}_j \cdot 2^k}$ ; so, depending on  $B$ 's challenge, encryptions of  $s_3$  are prepared as in Game 2 or Game 3. Similar arguments work for  $j^* = 1, 2, 4$ , and (12) follows. (The  $\mathbf{O}(2^{-k})$  term in (12) accounts for the statistical defect caused by choosing  $\mathbb{G}_{\text{rnd}}$ -exponents from  $\mathbb{Z}_{N/4}$ , resp.  $\mathbb{Z}_{N^2/4}$ .)

Using the definition of  $u$  and  $v$ , our change in Game 3 implies  $\tilde{Z} = (u^{vk} v)^r \cdot u^{\tilde{r}} \cdot h^K$  when a key part  $s_j$  is to be encrypted. (However, note that we still have  $Z = (u^{vk} v)^{r \cdot N^2}$  in any case.) This means that  $A$  still obtains information about the  $s_j$  (beyond what is public from  $pk$ ) from its KDM queries, but this information is limited to values  $\text{LAF}_{Fpk, t}([s_j]_{\mathbb{Z}_p^*})$ . We will now further cap this leaked information by making  $\text{LAF}_{Fpk, t}(\cdot)$  lossy. Namely, in **Game 4**, we use the LAF trapdoor  $Ftd$  initially sampled along with  $Fpk$ . Concretely, when preparing a ciphertext  $C$  for

$A$ , we sample  $t_c$  using  $t_c \leftarrow \text{FTag}(Ftd, t_a)$  for the corresponding auxiliary tag  $t_a = vk$ . A simple reduction shows

$$\Pr[out_4 = 1] - \Pr[out_3 = 1] = \text{Adv}_{\text{LAF}, C_2}^{\text{ind}}(k)$$

for a suitable adversary  $C_2$  on LAF's indistinguishability.

In **Game 5**, we reject all decryption queries of  $A$  that re-use a verification key  $vk$  from one of the KDM ciphertexts. To show that this change does not significantly affect  $A$ 's view, assume a decryption query  $C$  that re-uses a key  $vk = vk^*$  from a KDM ciphertext  $C^*$ . Recall that  $C$  contains a signature  $\sigma$  of  $X := ((G_j, \tilde{G}_j)_{j=1}^2, Z, \tilde{Z}, C_E, t_c)$  under an honestly generated Signature-verification-key  $vk = t_a = t_a^* = vk^*$ . Since we assumed  $t = (t_c, t_a) = (t_c^*, t_a^*) = t^*$ , and  $A$  is not allowed to query unchanged challenge ciphertexts for decryption, we must have  $(X, \sigma) \neq (X^*, \sigma^*)$  for the corresponding signed message  $X^*$  and signature  $\sigma^*$  from  $C^*$ . Hence, Game 4 and Game 5 only differ when  $A$  manages to forge a signature. A straightforward reduction to the strong OT-EUF-CMA security of Sig yields

$$\Pr[out_5 = 1] - \Pr[out_4 = 1] = q(k) \cdot \text{Adv}_{\text{LAF}, F}^{\text{seuf-cma}}(k)$$

for a forger  $F$  against Sig that makes at most one signature query.

In **Game 6.i** (for  $0 \leq i \leq q$ ), the first  $i$  challenge ciphertexts are prepared using  $Z = \hat{g}^{N^2}$  and  $\tilde{Z} = \hat{g} \cdot u^{\tilde{r}} \cdot h^K$  (if a key component  $s_j$  is to be encrypted), resp.  $\tilde{Z} = \hat{g} \cdot u^{\tilde{r}} \cdot h^{K+2^k[M]_Z}$  (if a constant  $M \in \mathcal{M}$  is to be encrypted) for an independently uniform  $\hat{g} \leftarrow \mathbb{G}_{\text{rnd}}$  drawn freshly for each ciphertext. Obviously, Game 6.0 is identical to Game 5:

$$\Pr[out_{6.0} = 1] = \Pr[out_5 = 1].$$

We will move from Game 6.i to Game 6.(i + 1) in several steps. During these steps, we denote with  $C = ((G_j, \tilde{G}_j)_{j=1}^2, Z, \tilde{Z}, C_E, t_c, vk, \sigma)$  the (i + 1)-st KDM ciphertext.

In **Game 6.i.1**, we change the  $\mathbb{G}_{\text{rnd}}$  parts of  $G_1, G_2$  from a Diffie-Hellman tuple (with respect to  $g_1, g_2$ ) to a random tuple. Concretely, if an  $s_j$  is to be encrypted, we set  $(G_1, G_2) = (g_1^{r_1}/h^{\alpha_1 \cdot 2^k}, g_2^{r_2}/h^{\alpha_2 \cdot 2^k})$ ; if a constant  $M$  is encrypted, we set  $(C_1, C_2) = (g_1^{r_1}, g_2^{r_2})$ , in both cases for independently uniform  $r_1, r_2 \leftarrow \mathbb{Z}_{N/4}$ . The  $\mathbb{G}_{\text{msg}}$  parts of  $G_1, G_2$  are thus unchanged compared to Game 6.i. Note that the  $\tilde{G}_j$  are still prepared as  $\tilde{G}_j = \tilde{g}_j^{\tilde{r}}/h^{\tilde{\alpha} \cdot 2^k}$ , resp.  $\tilde{G}_j = \tilde{g}_j^{\tilde{r}}$ . A straightforward reduction to the DDH assumption in  $\mathbb{G}_{\text{rnd}}$  yields

$$\sum_{i=1}^{q(k)} (\Pr[out_{6.i} = 1] - \Pr[out_{6.i.1} = 1]) = q(k) \cdot \text{Adv}_{\mathbb{G}_{\text{rnd}}, D_1}^{\text{ddh}}(k) + \mathbf{O}(2^{-k})$$

for a suitable  $D_1$ . The  $\mathbf{O}(2^{-k})$  error term accounts for the statistical difference caused by the choice of exponents  $r_j \leftarrow \mathbb{Z}_{N/4}$ , which induces an only almost-uniform distribution on group elements  $g^{r_j}$ . Note that at this point,  $Z$  and  $\tilde{Z}$  are still computed as in (11), even if an  $s_j$  is to be encrypted.

In **Game 6.i.2**, we compute  $Z$  and  $\tilde{Z}$  as  $Z = \hat{g}^{N^2}$  and  $\tilde{Z} = \hat{g} \cdot u^{\tilde{r}} \cdot h^K$ , resp.  $\tilde{Z} = \hat{g} \cdot u^{\tilde{r}} \cdot h^{K+2^k[M]_Z}$  for a fresh  $\hat{g} \leftarrow \mathbb{G}_{\text{rnd}}$ . Thus, the difference to Game 6.i.1 is that we substitute a  $\mathbb{G}_{\text{rnd}}$ -element computed as  $G_1^{[s_1]_Z \cdot vk + [s_3]_Z} G_2^{[s_2]_Z \cdot vk + [s_4]_Z}$  with a fresh random  $\hat{g}$ . To show that this change affects  $A$ 's view only negligibly, it suffices to show that  $A$ 's statistical information about

$$X := \text{dlog}_g \left( G_1^{[s_1]_Z \cdot vk + [s_3]_Z} G_2^{[s_2]_Z \cdot vk + [s_4]_Z} \right) = \gamma_1 r_1 ([s_1]_Z \cdot vk + [s_3]_Z) + \gamma_2 r_2 ([s_2]_Z \cdot vk + [s_4]_Z) \pmod{|\mathbb{G}_{\text{rnd}}|}$$

(for some arbitrary generator  $g$  of  $\mathbb{G}_{\text{rnd}}$  and  $\gamma_j = \text{dlog}_g(g_j)$ ) is negligible. This part of the proof will be rather delicate, since we will have to argue that both  $A$ 's KDM queries and  $A$ 's decryption queries yield (almost) no information about  $X$ .

First, observe that  $A$  gets the following information about the  $s_j$ :

- $pk$  reveals (through  $u$  and  $v$ ) precisely the two linear equations  $\gamma_1 [s_1]_Z + \gamma_2 [s_2]_Z \pmod{|\mathbb{G}_{\text{rnd}}|}$  and  $\gamma_1 [s_3]_Z + \gamma_2 [s_4]_Z \pmod{|\mathbb{G}_{\text{rnd}}|}$  about the  $s_j$ , where the  $\gamma_j$  are as above. For  $r_1 \neq r_2$ , these equations are linearly independent of the equation that defines  $X$ . Hence, for uniform  $r_1, r_2$ ,  $X$  is (almost) independent of  $pk$ .

- By LAF's lossiness, KDM ciphertexts reveal (through  $C_E = E(K, \text{LAF}_{Fpk,t}([s_j]_{\mathbb{Z}_p^n}))$ ) in total at most one equation  $\omega_1 a_j + \omega_2 b_j + \sum_{i=0}^{n-2} \omega_{3+i} c_{j,i} \pmod p$  for each  $j$ , where  $(a_j, b_j, c_{j,0}, \dots, c_{j,n-3}) := [s_j]_{\mathbb{Z}_p^n}$ , and the  $\omega_i$  are the (fixed) coefficients from LAF's lossiness property. (Recall the encodings  $[s_j]_{\mathbb{Z}}, [s_j]_{\mathbb{Z}_p^n}$  of the  $s_j = (a_j, b_j, c_j)$  from (10).) Hence, the  $b_j \in \mathbb{Z}_{p \cdot 2^k}$  fully blind the information released about the  $c_j \in \mathbb{Z}_{2^{k-2N}}$  through the KDM ciphertexts. Thus, KDM ciphertexts reveal no information about  $c_j \pmod{|\mathbb{G}_{\text{rnd}}|}$  and hence also about  $[s_j]_{\mathbb{Z}} \pmod{|\mathbb{G}_{\text{rnd}}|}$ . Consequently, even given  $pk$  and the KDM ciphertexts,  $X$  is statistically close to independently uniform. This already shows that our change from Game 6.i.2 affects  $A$ 's view only negligibly if  $A$  makes no decryption queries. It remains to show that decryption queries yield no additional information about the  $s_j$ .

To do so, let us say that a ciphertext  $C$  is *inconsistent* iff there exist no  $r, \tilde{r}$  with  $(G_j, \tilde{G}_j) = (g_j^r, \tilde{g}_j^{\tilde{r}})$  for both  $j \in \{1, 2\}$ . Note that the decryption of a consistent ciphertext yields no information about the  $s_j$  beyond  $pk$ . ( $pk$  and  $r, \tilde{r}$  determine the values  $Z, Z'$  computed during decryption; everything else follows from  $Z'$  and  $C$ .) Thus, it suffices to prove the following lemma (which we do after the main proof):

**Lemma 4.2.** *In the situation of Game 6.i.l (for  $l \in \{1, 2\}$ ), let  $\text{bad}_{\text{query.i.l}}$  be the event that  $A$  places an inconsistent decryption query that is not rejected. Then*

$$\sum_{i=1}^{q(k)} (\Pr[\text{bad}_{\text{query.i.1}}] + \Pr[\text{bad}_{\text{query.i.2}}]) \leq 2 \cdot q(k) \cdot \text{Adv}_{\text{LAF}, F}^{\text{eva}}(k) + \mathbf{O}(2^{-3k}).$$

for a suitable evasiveness adversary  $F$  on LAF.

By our discussion above and Lemma 4.2, we obtain that

$$\sum_{i=1}^{q(k)} |\Pr[\text{out}_{6.i.2} = 1] - \Pr[\text{out}_{6.i.1} = 1]| \leq 2 \cdot q(k) \cdot \text{Adv}_{\text{LAF}, F}^{\text{eva}}(k) + \mathbf{O}(2^{-3k}).$$

In **Game 6.i.3**, we reverse the change from Game 6.i.1. Concretely, we prepare the  $G_j$  as  $G_j = g_j^r / h^{\alpha_j \cdot 2^k}$ , resp.  $G_j = g_j^r$  for a single  $r \leftarrow \mathbb{Z}_{N/4}$ . Another straightforward reduction to the DDH assumption in  $\mathbb{G}_{\text{rnd}}$  yields that

$$\sum_{i=1}^{q(k)} (\Pr[\text{out}_{6.i.3} = 1] - \Pr[\text{out}_{6.i.2} = 1]) = q(k) \cdot \text{Adv}_{\mathbb{G}_{\text{rnd}}, D_2}^{\text{ddh}}(k) + \mathbf{O}(2^{-k})$$

for a suitable  $D_2$ . To close the hybrid argument, note that Games 6.i.3 and 6.(i+1) are identical.

In **Game 7**, we clear the  $\mathbb{G}_{\text{msg}}$ -component of  $\tilde{Z}$  in all ciphertexts prepared for  $A$ . That is, instead of computing  $\tilde{Z} = \hat{g} \cdot u^{\tilde{r}} \cdot h^K$ , resp.  $\tilde{Z} = \hat{g} \cdot u^{\tilde{r}} \cdot h^{K+[M]_{\mathbb{Z}}}$  for a freshly uniform  $\hat{g} \leftarrow \mathbb{G}_{\text{rnd}}$ , we set  $\tilde{Z} = \hat{g} \cdot u^{\tilde{r}}$ . (We stress that we still compute  $Z = \hat{g}^{N^2}$ .) Since all  $\tilde{Z}$  already have an independently uniform  $\mathbb{G}_{\text{rnd}}$ -component, a straightforward reduction to the DCR assumption yields

$$\Pr[\text{out}_{6.q} = 1] - \Pr[\text{out}_7 = 1] = \text{Adv}_{\mathbb{Z}_{N^3}^*, E}^{\text{dcr}}(k) + \mathbf{O}(2^{-k})$$

for a DCR distinguisher  $E$ . Note that because of the re-randomizability of DCR, there is no factor of  $q(k)$ , even though we substitute many group elements at once. However, since the precise order of  $\mathbb{G}_{\text{rnd}}$  is not known, this re-randomization costs us an error term of  $\mathbf{O}(2^{-k})$ .

In **Game 8**, we substitute the symmetric ciphertexts  $C_E$  in all KDM ciphertexts by encryptions of random messages. By our change in Game 7, we do not use the symmetric keys  $K$  used to produce  $C_E$  anywhere else. Thus, a reduction to the IND-CPA security of  $(E, D)$  gives

$$\Pr[\text{out}_7 = 1] - \Pr[\text{out}_8 = 1] = q(k) \cdot \text{Adv}_{(E, D), G}^{\text{ind-cpa}}(k)$$

for an IND-CPA adversary  $G$ . Note that in Game 8,  $A$ 's view is independent of the challenge bit  $b$  initially selected by the KDM challenger. Hence, we have

$$\Pr[out_8 = 1] = 1/2.$$

Taking things together yields the theorem.  $\square$

It remains to prove Lemma 4.2, which we do now:

*Proof.* Let  $\text{bad}_{\text{tag}.i.\ell}$  be the event that in Game 6.i.l,  $A$  submits a decryption query that refers to a lossy tag  $t$ . By our change from Game 5, we may assume that  $vk$  and thus  $t$  is fresh, i.e., has not been generated through  $\text{FTag}$  by the experiment. Thus, by  $\text{LAF}$ 's evasiveness,  $\text{bad}_{\text{tag}.i.\ell}$  can occur only with negligible probability. Concretely, it is easy to construct an evasiveness adversary  $F$  with

$$\sum_{i=1}^{q(k)} (\Pr[\text{bad}_{\text{tag}.i.1}] + \Pr[\text{bad}_{\text{tag}.i.2}]) \leq 2 \cdot q(k) \cdot \text{Adv}_{\text{LAF},F}^{\text{eva}}(k). \quad (13)$$

Now suppose that we are in Game 6.i.l, and say that  $\text{bad}_{\text{tag}.i.\ell}$  does not occur. Consider an inconsistent decryption query  $C = ((G_j, \tilde{G}_j)_{j=1}^2, Z, \tilde{Z}, C_E, t_c, vk, \sigma)$  from  $A$ . Write  $(G_j, \tilde{G}_j) = (g_j^{r_j} \cdot h^{\delta_j}, \tilde{g}_j^{r_j} \cdot h^{\tilde{\delta}_j})$  (for  $j \in \{1, 2\}$ ) and  $\tilde{Z} = \tilde{g} \cdot h^{\tilde{\delta}}$  for  $\tilde{g} \in \mathbb{G}_{\text{rnd}}$ . Recall that decryption first checks

$$Z \stackrel{?}{=} \left( G_1^{[s_1]_{\mathbb{Z}} \cdot vk + [s_3]_{\mathbb{Z}}} G_2^{[s_2]_{\mathbb{Z}} \cdot vk + [s_4]_{\mathbb{Z}}} \right)^{N^2}, \quad (14)$$

then computes

$$Z' = G_1^{[s_1]_{\mathbb{Z}} \cdot vk + [s_3]_{\mathbb{Z}}} G_2^{[s_2]_{\mathbb{Z}} \cdot vk + [s_4]_{\mathbb{Z}}} \tilde{G}_1^{[s_1]_{\mathbb{Z}}} \tilde{G}_2^{[s_2]_{\mathbb{Z}}}, \quad (15)$$

and finally attempts to find  $K \in \{0, 1\}^k, M \in \mathcal{M}$  with

$$K + 2^k \cdot [M]_{\mathbb{Z}} = \text{dlog}_h(\tilde{Z}/Z') = \tilde{\delta} - \delta' \pmod{N^2} \quad (16)$$

for  $\delta' = \delta_1([s_1]_{\mathbb{Z}} \cdot vk + [s_3]_{\mathbb{Z}}) + \delta_2([s_2]_{\mathbb{Z}} \cdot vk + [s_4]_{\mathbb{Z}}) + \tilde{\delta}_1[s_1]_{\mathbb{Z}} + \tilde{\delta}_2[s_2]_{\mathbb{Z}}$ . As usual, we write  $s_j = (a_j, b_j, c_j) \in \mathcal{M} = \mathbb{Z}_{2^{3k}} \times \mathbb{Z}_{p \cdot 2^k} \times \mathbb{Z}_{N \cdot 2^{k-2}}$  for  $i \in \{1, 2\}$ .

**$h$ -inconsistent ciphertexts.** First, consider the case that there is a  $j^* \in \{1, 2\}$  with  $\delta_{j^*} \neq 0 \pmod{N^2}$  or  $\tilde{\delta}_{j^*} \neq 0 \pmod{N^2}$ . (In that case, we may say that  $C$  is  $h$ -inconsistent.) Then, we claim that either  $C$  is rejected, or  $A$  has (information-theoretically) successfully narrowed down the value of  $\delta'$  to a set of size at most  $2^k$ . Indeed,  $C_E$  determines  $K$  and thus  $\text{LAF}_{Fpk,t}([M]_{\mathbb{Z}_p^n}) = \text{D}(K, C_E)$  by (E, D)'s key-uniqueness. Moreover, since we assumed  $\neg \text{bad}_{\text{tag}.i.\ell}$ , the used tag  $t$  is injective, and so  $\text{LAF}_{Fpk,t}([M]_{\mathbb{Z}_p^n})$  determines  $M$  up to  $[b/p] \in \mathbb{Z}_{2^k}$ . (Recall that the encoding  $[M]_{\mathbb{Z}_p^n}$  only depends on  $b \pmod{p}$ .) Thus, a non-rejected ciphertext allows to infer (a  $2^k$ -candidate set for)  $\delta'$  by substituting  $K, M$ , and  $\tilde{\delta}$  (as defined by  $\tilde{Z}$ ) into (16).

However, we will now argue that  $\delta'$  has min-entropy at least  $5k$ , even given  $pk$  and all KDM ciphertexts. Hence,  $A$  cannot predict a correct  $2^k$ -candidate set for  $\delta'$  (and thus cannot supply an  $h$ -inconsistent decryption query that is not rejected) with non-negligible probability. To prove our claim, we need some preparations. For concreteness, say that  $\delta_1 \neq 0 \pmod{N^2}$  (the other cases are similar). Then either  $P^2 \nmid \delta_1$  or  $Q^2 \nmid \delta_1$  (or both) for the factors  $P, Q$  of  $N$ . Without losing generality, say that  $P^2 \nmid \delta_1$ , so that the subterm  $\delta_1 \cdot [s_3]_{\mathbb{Z}} \pmod{N^2}$  of  $\delta'$  reveals  $[s_3]_{\mathbb{Z}} \pmod{P}$ . Furthermore,

$$\begin{aligned} [s_3]_{\mathbb{Z}} &\stackrel{(10)}{<} 2^{5k-2} \cdot p \cdot N \leq 2^{5k+1} \cdot 2^{\ell_{\text{LAF}}} \cdot |\mathbb{G}_{\text{rnd}}| && \ell_{\text{LAF}} = \frac{\ell_N + k + 1}{n-2} < \frac{\ell_N + k + 1}{n-2} \\ &\stackrel{P \geq 2^{(\ell_N/2) - k}}{\leq} 2^{(6+1/(n-2))k+2 - (1/2 - 1/(n-2))\ell_N} \cdot |\mathbb{G}_{\text{rnd}}| \cdot P && \stackrel{\ell_N \geq 25k+8}{\leq} \frac{n=6}{|\mathbb{G}_{\text{rnd}}|} \cdot P. \end{aligned} \quad (17)$$



Using  $\gcd(P, |\mathbb{G}_{\text{rnd}}|) = 1$ , the Chinese Remainder Theorem hence gives that  $[s_3]_{\mathbb{Z}} \bmod |\mathbb{G}_{\text{rnd}}|$  and  $[s_3]_{\mathbb{Z}} \bmod P$  uniquely determine  $[s_3]_{\mathbb{Z}}$ . Thus, since  $[s_3]_{\mathbb{Z}}$  initially has min-entropy at least  $5k - 2 + \ell_{\text{LAF}} + \ell_N$ , even revealing  $[s_3]_{\mathbb{Z}} \bmod |\mathbb{G}_{\text{rnd}}|$  (through  $pk$ ) leaves at least  $5k + \ell_{\text{LAF}}$  bits of min-entropy in  $[s_3]_{\mathbb{Z}} \bmod P$ . The KDM ciphertexts reveal no more than  $\ell_{\text{LAF}}$  bits of entropy about  $[s_3]_{\mathbb{Z}} \bmod P$ , so that  $[s_3]_{\mathbb{Z}} \bmod P$  has min-entropy at least  $5k$ .

However,  $C$  implies  $2^k$  candidates for  $\delta'$  which, given  $s_1, s_2, s_4$ , in turn determine  $2^k$  candidates for  $[s_3]_{\mathbb{Z}} \bmod P$ . So, assuming  $\neg \text{bad}_{\text{tag}, i, j}$ , the probability that a given  $h$ -inconsistent  $C$  implies “the correct  $[s_3]_{\mathbb{Z}} \bmod P$ ” (which is a prerequisite for non-rejection), is at most  $2^{-4k}$ . The case for  $\delta_2 \neq 0 \bmod N^2$  makes the analogous argument about  $[s_4]_{\mathbb{Z}}$ , and the cases  $\widetilde{\delta}_{j^*} \neq 0 \bmod N^2$  consider  $[s_{j^*}]_{\mathbb{Z}}$ , then using  $\delta_1, \delta_2 = 0 \bmod N^2$ .

**$g$ -inconsistent ciphertexts.** Now assume that  $\delta_1 = \delta_2 = \widetilde{\delta}_1 = \widetilde{\delta}_2 = 0 \bmod N^2$ . Since  $C$  is inconsistent,  $r_1 \neq r_2 \bmod |\mathbb{G}_{\text{rnd}}|$  or  $\widetilde{r}_1 \neq \widetilde{r}_2 \bmod |\mathbb{G}_{\text{rnd}}|$ . We may call such ciphertexts  *$g$ -inconsistent*.

Let us first assume  $r_1 \neq r_2 \bmod |\mathbb{G}_{\text{rnd}}|$ . Recall that  $|\mathbb{G}_{\text{rnd}}| = (P-1)(Q-1)/4$ , where  $P, Q$  are safe primes. Hence, without loss of generality, we can assume that  $r_1 \neq r_2 \bmod (P-1)/2$ , where  $(P-1)/2$  is prime. We claim that the value

$$\begin{aligned} X' &:= \text{dlog}_g \left( G_1^{[s_1]_{\mathbb{Z}} \cdot vk + [s_3]_{\mathbb{Z}}} G_2^{[s_2]_{\mathbb{Z}} \cdot vk + [s_4]_{\mathbb{Z}}} \right) \\ &= \gamma_1 r_1 ([s_1]_{\mathbb{Z}} \cdot vk + [s_3]_{\mathbb{Z}}) + \gamma_2 r_2 ([s_2]_{\mathbb{Z}} \cdot vk + [s_4]_{\mathbb{Z}}) \bmod |\mathbb{G}_{\text{rnd}}| \end{aligned}$$

is (up to a small statistical defect) independently and uniformly random modulo  $(P-1)/2$  from  $A$ 's point of view. Hence, also the value  $g^{X' \cdot N^2}$  from (14) to which  $Z$  is compared is unpredictable for  $A$ . This can be seen as in the discussion after Game 6.i.2, where a similar value  $X$  is seen as essentially uniform. In particular,  $pk$  contains two linear equations (in the  $s_j$ ) that are independent of  $X'$ , and the information about  $X'$  from the KDM challenges is suitably blinded by the  $b_j$ -components of the  $s_j$ . There are two differences to Game 6.i.2: first, the  $r_i$  in our case are adversarially chosen, and so could be equal modulo a factor of  $|\mathbb{G}_{\text{rnd}}|$ . Thus, we can only conclude linear independence modulo  $(P-1)/2$ . Second, if  $\ell = 1$ , then  $A$  additionally receives one  $g$ -inconsistent challenge ciphertext that reveals another linear equation

$$\gamma_1 r_1^* ([s_1]_{\mathbb{Z}} \cdot vk^* + [s_3]_{\mathbb{Z}}) + \gamma_2 r_2^* ([s_2]_{\mathbb{Z}} \cdot vk^* + [s_4]_{\mathbb{Z}}) \bmod |\mathbb{G}_{\text{rnd}}|$$

about the  $s_j$ . However, by our change from Game 5, we may assume  $vk \neq vk^*$ , so that this linear equation is also independent of  $X'$ . Hence,  $X'$  looks (almost) independently uniform modulo  $(P-1)/2$  to  $A$ , so that  $r_1 \neq r_2 \bmod |\mathbb{G}_{\text{rnd}}|$  implies rejection (because of the check (14)) except with probability at most  $\mathbf{O}(2^{-4k})$ .<sup>11</sup>

Let us now assume  $r_1 = r_2 \bmod |\mathbb{G}_{\text{rnd}}|$  but  $\widetilde{r}_1 \neq \widetilde{r}_2 \bmod |\mathbb{G}_{\text{rnd}}|$ . This case is similar to the case  $r_1 \neq r_2 \bmod |\mathbb{G}_{\text{rnd}}|$ , but simpler. A similar analysis as above yields that (the  $\mathbb{G}_{\text{rnd}}$ -component of)  $Z'$  is unpredictable for  $A$ . Specifically,  $C$  will be rejected except with probability  $\mathbf{O}(2^{-4k})$ .

Summarizing, and using a union bound, we obtain that

$$\Pr[\text{bad}_{\text{query}, i, \ell} \mid \neg \text{bad}_{\text{tag}, i, \ell}] \leq q'(k) \cdot 2^{-4k} = \mathbf{O}(2^{-3k})$$

for the number  $q'(k)$  of  $A$ 's decryption queries. Combining with (13) shows the lemma. We stress that in this proof, it may appear that several bounds have been chosen too conservatively. In particular, we arrive at an error bound that is significantly smaller than, e.g.,  $\mathbf{O}(2^{-k})$ . These extra “entropy cushions” are used in the multi-user case.  $\square$

<sup>11</sup>This argument follows in the footsteps of IND-CCA security proofs of PKE schemes based on hash proof systems [17, 18]. Specifically, the knowledge that an adversary receives through  $pk$  and one inconsistent ciphertext in our case is essentially the same as in the analysis of the Kurosawa-Desmedt scheme [26].

### 4.3 Security proof (multi-user case)

**Theorem 4.3.** *Assume the DCR assumption holds in  $\mathbb{Z}_{N^3}$ , the DDH assumption holds in  $\mathbb{G}_{\text{rnd}}$ , LAF is an LAF, Sig is a strongly OT-EUF-CMA secure signature scheme, H is collision-resistant, and (E, D) is a key-unique IND-CPA secure SKE scheme. Then PKE is  $n$ -CIRC-CCA-secure for every polynomial  $n = n(k)$ .*

*Proof sketch.* The proof is very similar to the proof of Theorem 4.1. The way we achieve multi-user KDM security is to have  $n$  “virtual” secret keys  $s^i$  that are set up as

$$s^i = (s_1^i, s_2^i, s_3^i, s_4^i) = (s_1, s_2, s_3, s_4) + (\hat{s}_1^i, \hat{s}_2^i, \hat{s}_3^i, \hat{s}_4^i) \quad (18)$$

(with component-wise addition, also in each sum  $s_j + \hat{s}_j^i$ ) for uniformly chosen  $\hat{s}^i = (\hat{s}_j^i)_{j=1}^4 \leftarrow \mathcal{M}^4$ . Intuitively, the  $\hat{s}^i$  blind a single  $s = (s_j)_{j=1}^4 \in \mathcal{M}^4$  in several instances. While the  $\hat{s}^i$  are all uniform, however, we choose the  $s_j = (a_j, b_j, c_j) \in \mathcal{M}$  with “small” components. Concretely, we pick  $(a_j, b_j, c_j) \leftarrow \mathbb{Z}_{3k} \times \mathbb{Z}_p \times \mathbb{Z}_{N/4}$  and embed  $s_j$  into  $\mathcal{M}$  in the natural way. This choice guarantees that  $[s_j^i]_{\mathbb{Z}} = [s_j]_{\mathbb{Z}} + [\hat{s}_j^i]_{\mathbb{Z}}$  and  $[s_j^i]_{\mathbb{Z}_p^n} = [s_j]_{\mathbb{Z}_p^n} + [\hat{s}_j^i]_{\mathbb{Z}_p^n}$ , except with probability  $\mathbf{O}(2^{-k})$ . Intuitively, the  $\hat{s}^i$  can be known to  $A$  at all times, while we will try to argue that the information  $A$  has about  $s$  is very limited.

We will now go through the proof of Theorem 4.1, and sketch the necessary modifications for the multi-user case. Generally, we assume a setup of keys as in (18) (which guarantees independently uniform  $s^i$ ). **Games 1 to 6.i.2** are as with Theorem 4.1, where the changes apply of course to KDM queries under all public keys. The corresponding reductions to DCR, DDH, the indistinguishability of LAF, and the security of Sig and H apply almost verbatim. The only noteworthy change occurs in the justification of the change from Game 6.i.2.

Here, we have to argue that  $A$  obtains no useful information about the  $s_j^i \bmod |\mathbb{G}_{\text{rnd}}|$  from all public keys  $pk^i$ , all challenge ciphertexts, and all decryption queries. First, each  $pk^i$  reveals (through the corresponding  $u, v$ ) exactly two linear equations

$$\begin{aligned} \gamma_1[s_1^i]_{\mathbb{Z}} + \gamma_2[s_2^i]_{\mathbb{Z}} &= (\gamma_1[s_1]_{\mathbb{Z}} + \gamma_2[s_2]_{\mathbb{Z}}) + (\gamma_1[\hat{s}_1^i]_{\mathbb{Z}} + \gamma_2[\hat{s}_2^i]_{\mathbb{Z}}) \bmod |\mathbb{G}_{\text{rnd}}| \\ \gamma_1[s_3^i]_{\mathbb{Z}} + \gamma_2[s_4^i]_{\mathbb{Z}} &= (\gamma_1[s_3]_{\mathbb{Z}} + \gamma_2[s_4]_{\mathbb{Z}}) + (\gamma_1[\hat{s}_3^i]_{\mathbb{Z}} + \gamma_2[\hat{s}_4^i]_{\mathbb{Z}}) \bmod |\mathbb{G}_{\text{rnd}}| \end{aligned}$$

about  $s = (s_1, s_2, s_3, s_4)$ . These equations only depend on  $\gamma_1[s_1]_{\mathbb{Z}} + \gamma_2[s_2]_{\mathbb{Z}} \bmod |\mathbb{G}_{\text{rnd}}|$  and  $\gamma_1[s_3]_{\mathbb{Z}} + \gamma_2[s_4]_{\mathbb{Z}} \bmod |\mathbb{G}_{\text{rnd}}|$  (but not on other information about the  $s_j$ ), just like in the single-user case. Similarly, since  $[s_j^i]_{\mathbb{Z}_p^n} = [s_j]_{\mathbb{Z}_p^n} + [\hat{s}_j^i]_{\mathbb{Z}_p^n}$ , all challenge ciphertexts depend only on

$$\omega_1 a_j + \omega_2 b_j + \sum_{i=0}^{n-2} \omega_{3+i} c_{j,i} \bmod p$$

(for  $(a_j, b_j, c_{j,0}, \dots, c_{j,n-3}) := [s_j]_{\mathbb{Z}_p^n}$ ) and the  $\hat{s}_j^i$ . This equation is fully blinded by  $b_j \in \mathbb{Z}_p$ . Next, carefully considering the (slightly reduced) entropy in the  $s_j$ , we can prove an analog of Lemma 4.2 for the multi-user case.

As in Lemma 4.2, we will have to argue that the additional linear equation released about the  $s_j$  by the  $(i+1)$ -st challenge ciphertext in Game 6.i.1 does not help in producing  $g$ -inconsistent decryption queries. The corresponding analysis is the same as that in Lemma 4.2, but only considers the  $s_j$ -terms (and not the  $\hat{s}_j^i$ ) in the exponent. Furthermore, because of the reduced entropy, the  $\mathbf{O}(2^{-3k})$  bound from the lemma will become  $\text{poly} \cdot 2^{-k}$ .

Finally, to justify the change from Game 6.i.2, it suffices to note that hence,  $A$ 's view is essentially independent of

$$\gamma_1 r_1 ([s_1]_{\mathbb{Z}} \cdot vk + [s_3]_{\mathbb{Z}}) + \gamma_2 r_2 ([s_2]_{\mathbb{Z}} \cdot vk + [s_4]_{\mathbb{Z}}) \bmod |\mathbb{G}_{\text{rnd}}|,$$

where  $r_1 \neq r_2$  are the  $\mathbb{G}_{\text{rnd}}$ -exponents of the considered  $C_1, C_2$ .

The remaining **Games 6.i.3 to Game 8** are again as with Theorem 4.1, of course again applied to KDM queries under all public keys. The corresponding reductions to DDH, DCR, and the IND-CPA security of (E, D) apply verbatim.  $\square$

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