

# Simulation-based Selective Opening CCA Security for PKE from Key Encapsulation Mechanisms <sup>\*</sup>

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**Abstract.** We study simulation-based, selective opening security against chosen-ciphertext attacks (SIM-SO-CCA security) for public key encryption (PKE). In a selective opening, chosen-ciphertext attack (SO-CCA), an adversary has access to a decryption oracle, sees a vector of ciphertexts, adaptively chooses to open some of them, and obtains the corresponding plaintexts and random coins used in the creation of the ciphertexts. The SIM-SO-CCA notion captures the security of unopened ciphertexts with respect to probabilistic polynomial-time (ppt) SO-CCA adversaries in a semantic way: what a ppt SO-CCA adversary can compute can also be simulated by a ppt simulator with access only to the opened messages. Building on techniques used to achieve weak deniable encryption and non-committing encryption, Fehr *et al.* (Eurocrypt 2010) presented an approach to constructing SIM-SO-CCA secure PKE from extended hash proof systems (EHPs), collision-resistant hash functions and an information-theoretic primitive called Cross Authentication Codes (XACs). We generalize their approach by introducing a special type of Key Encapsulation Mechanism (KEM) and using it to build SIM-SO-CCA secure PKE. We investigate what properties are needed from the KEM to achieve SIM-SO-CCA security. We also give three instantiations of our construction. The first uses hash proof systems, the second relies on the  $n$ -Linear assumption, and the third uses indistinguishability obfuscation ( $i\mathcal{O}$ ) in combination with extracting, puncturable Pseudo-Random Functions in a similar way to Sahai and Waters (STOC 2014). Our results establish the existence of SIM-SO-CCA secure PKE assuming only the existence of one-way functions and  $i\mathcal{O}$ . This result further highlights the simplicity and power of  $i\mathcal{O}$  in constructing different cryptographic primitives.

## 1 Introduction

Selective Opening Attacks (SOAs) concern a multi-user scenario, where an adversary adaptively corrupts a set of users to get their secret state information. In the case of public key encryption (PKE), we assume that several senders send ciphertexts encrypting possibly correlated messages to a receiver. The SOA adversary is able to (adaptively) corrupt some senders, exposing their messages and also the random coins used to generate their ciphertexts. Security against selective opening attacks (SOA security) considers whether the uncorrupted ciphertexts remain secure.

There are two ways of formalizing SOA security: indistinguishability-based (IND-SO) and simulation-based (SIM-SO). According to whether the adversary is able to access to a decryption oracle during its attack, SOA security is further classified into IND-SO-CPA, IND-SO-CCA, SIM-SO-CPA and SIM-SO-CCA. In the formalization of SOAs, we allow a probabilistic polynomial-time (ppt) adversary to get the public key, a vector of challenge ciphertexts, and to adaptively corrupt (open) some ciphertexts to obtain opened plaintexts and random coins (and also access to a decryption oracle in the case of SO-CCA). The IND-SO security notions require that the real messages (used to generate the challenge ciphertexts) and re-sampled messages conditioned on the opened messages are computationally indistinguishable to an SOA adversary. Here we have to assume that the joint message distributions are *efficiently conditionally re-samplable* after the opened messages are exposed. On the other hand, the SIM-SO security notions have no such limitations. They require that what a probabilistic polynomial-time (ppt) SOA adversary can compute from the information it has learned can be simulated by a ppt simulator only knowing the opened plaintexts. SIM-SO security seems to be stronger than IND-SO security and significantly harder to achieve. We note the existence of a stronger IND-SO security notion, namely full IND-SO security, which imposes no limitation on the joint message distributions. However, there is no PKE achieving full IND-SO-CPA security yet. The relations among SIM-SO security, IND-SO security, and traditional IND-CPA/CCA security were explored in [5, 17].

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Lossy encryption [3] has shown itself to be a very useful tool in achieving IND-SO-CPA security. Different approaches to achieving IND-SO-CCA security include the use of lossy trapdoor functions [22], All-But- $N$  [14], and All-But-Many lossy trapdoor functions [15]. The basic idea is to make sure that only challenge ciphertexts are lossy encryptions, while ciphertexts queried by the adversary are normal encryptions. If there exists an efficient opener which can open a lossy encryption to an encryption of an arbitrary message, then an IND-SO-CCA secure PKE can also be shown to be SIM-SO-CCA secure. However, it seems that, to date, only a single, DCR-based PKE scheme [15] is known to have this property.

In [12], Fehr *et al.* proposed a black-box PKE construction to achieve SIM-SO-CCA security based on an Extended Hash Proof System (EHPS) associated with a subset membership problem, a collision-resistant hash function and a new information-theoretic primitive called Cross-Authentication Code (XAC). As pointed in [18, 19], a stronger property of XACs is needed to make the security proof rigorous.

## 1.1 Our Contributions

We generalize the black-box PKE construction of Fehr *et al.* [12] by using a special kind of key encapsulation mechanism (KEM) in combination with a strengthened XAC. Essentially, the KEM replaces the EHPS component in [12], opening up a new set of construction possibilities. In more detail:

- We characterise the properties needed of a KEM for our PKE construction to be SIM-SO-CCA secure. At a high level, these properties are that the KEM should have efficiently samplable and explainable (ESE) ciphertext and key spaces; *tailored* decapsulation; and *tailored*, constrained chosen-ciphertext (tCCCA) security. Here tailored decapsulation roughly means that the valid ciphertexts output by the KEM are sparse in the ciphertext space, while tCCCA security is an extension of the CCA security notion of [16]. If a KEM has all three properties, then we say that it is a *tailored KEM*.
- We show three constructions for tailored KEMs, including one based on hash proof systems (HPS) [8], a specific KEM from the  $n$ -Linear assumption [16] (but different from the HPS-based one) and one constructed from indistinguishability Obfuscation ( $i\mathcal{O}$ ) in combination with an extracting puncturable Pseudo-Random Function (PRF) [23]. Consequently, we obtain PKEs of three different types, all enjoying SIM-SO-CCA security. Thus, by adopting the KEM viewpoint, we significantly enlarge the scope of Fehr *et al.*'s construction.
- Since our PKE construction does not rely on collision-resistant hash functions, we immediately obtain the following results:
  - PKE with SIM-SO-CCA security from HPS and strengthened XACs (as compared to the PKE construction of [12] using EHPS, a strong XAC, and a collision-resistant hash function).
  - PKE with SIM-SO-CCA security from the  $n$ -Linear assumption in a way that differs from our HPS-based construction.
  - PKE with SIM-SO-CCA security assuming only the existence of  $i\mathcal{O}$  and one-way functions.

## 1.2 Ingredients of Our Main Construction

We follow the outline provided by the black-box PKE construction of Fehr *et al.* [12]. Observing that the EHPS used in [12] can actually be viewed as a KEM, our construction can be considered as a generalization of their result. We first outline the properties of KEMs and XACs needed for our result, before describing the construction and its security analysis at a high level.

The KEM component in our construction needs to be “tailored” with the following properties:

- (1) **Efficiently samplable and explainable (ESE) domains.** The key space  $\mathcal{K}$  and ciphertext space  $\mathcal{C}$  of the KEM should both be ESE domains. (Meaning that, given a randomised sampling algorithm  $\text{SampleD}$  for  $\mathcal{D}$ , there exists an efficient algorithm,  $\text{SampleD}^{-1}(\mathcal{D}, \cdot)$ , with the property that, given element  $d$  from a domain  $\mathcal{D}$  as input,  $\text{SampleD}^{-1}(\mathcal{D}, \cdot)$  outputs value  $R$  such that  $d$  can be “explained” as having been sampled using  $R$ , i.e.,  $d = \text{SampleD}(\mathcal{D}; R)$ .)
- (2) **Tailored decapsulation.** The valid ciphertexts output by the encapsulation algorithm constitute only a (small) subset of ciphertext space  $\mathcal{C}$ . When the input is a ciphertext randomly chosen from  $\mathcal{C}$ , the decapsulation will either output  $\perp$  with overwhelming probability or output a key that is almost uniformly distributed over  $\mathcal{K}$ .

**(3) Tailored, constrained CCA (tCCCA) security.** The output of the encapsulation algorithm is computationally indistinguishable from  $(K^R, \psi^R)$ , a pair of key and ciphertext randomly chosen from  $\mathcal{K} \times \mathcal{C}$ , for any ppt adversary, even if the adversary has access to a constrained decryption oracle. The adversary is allowed to make queries of the form  $(\psi, P(\cdot))$  to the constrained decryption oracle, where  $\psi$  is an element of  $\mathcal{C}$  and  $P(\cdot)$  is a ppt predicate, such that  $P(\cdot) : \mathcal{K} \rightarrow \{0, 1\}$  evaluates to 1 only for a negligible fraction of keys. The constrained decryption oracle will provide the decapsulated  $K$  to the adversary if only if  $P(K) = 1$ .

We will also need a strengthened XAC definition. A strengthened  $\ell$ -XAC is a collection of algorithms  $\text{XAC} = (\text{XGen}, \text{XAuth}, \text{XVer})$  having the following properties:

**Authentication and Verification.** Algorithm  $\text{XAuth}$  computes a tag  $T \leftarrow \text{XAuth}(K_1, \dots, K_\ell)$  from  $\ell$  inputs (which will be random keys in our construction). Any  $K_i$  used in generating the tag  $T$  almost always satisfies  $\text{XVer}(K_i, T) = 1$ .

**Security against impersonation/substitution attacks.** Security against impersonation attacks means that, given a tag  $T$ , a randomly chosen key  $K$  will almost always fail verification with this specific tag, i.e.,  $\text{XVer}(K, T) = 0$ . A substitution attack considers an (all-powerful) adversary who obtains a tag  $T = \text{XAuth}(K_1, \dots, K_\ell)$  and tries to forge a tag  $T' \neq T$  such that  $\text{XVer}(K_i, T') = 1$ , where  $K_i$  is one of the keys used in computing  $T$ . Security against substitution attacks requires that, if  $K_i$  is randomly chosen, then any adversary succeeds in outputting  $T'$  with  $T' \neq T$  and  $\text{XVer}(K_i, T') = 1$  with negligible probability, even if it is given  $T$  and all keys except  $K_i$  as input.

**Strongness and semi-uniqueness.** Strongness says that when  $K_i$  is randomly chosen, then  $K_i$ , given  $(K_j)_{j \in [\ell], j \neq i}$  and the tag  $T = \text{XAuth}(K_1, \dots, K_\ell)$ , is re-samplable with the correct probability distribution. That is to say, there exists a ppt algorithm  $\text{ReSample}((K_j)_{j \in [\ell], j \neq i}, T)$  such that  $\text{ReSample}$  outputs a key  $\hat{K}_i$  that is statistically indistinguishable from  $K_i$ , even given  $(K_j)_{j \in [\ell], j \neq i}$  and  $T = \text{XAuth}(K_1, \dots, K_\ell)$ . Semi-uniqueness says that it is possible to parse a key  $K$  as  $(K_x, K_y) \in \mathcal{K}_x \times \mathcal{K}_y$  for some sets  $\mathcal{K}_x, \mathcal{K}_y$ , and for every  $K_x \in \mathcal{K}_x$  and a tag  $T$ , there is at most one  $K_y \in \mathcal{K}_y$  such that  $(K_x, K_y)$  satisfies  $\text{XVer}((K_x, K_y), T) = 1$ .

### 1.3 Overview of Our Main Construction

Given a tailored KEM  $\text{KEM}$  and a strengthened  $(\ell + s)$ -XAC  $\text{XAC}$ , our construction of a PKE scheme  $\text{PKE}$  is as follows. (See Figure 4 for full details.)

- The public key of  $\text{PKE}$  is the public key  $pk_{kem}$  of  $\text{KEM}$ , an injective function  $F$  with domain  $\mathcal{C}^\ell$  and range  $(\mathcal{K}_y)^s$ , and a vector of values  $(K_{x_1}, \dots, K_{x_s}) \in (\mathcal{K}_x)^s$ . The secret key of  $\text{PKE}$  is  $sk_{kem}$ , the secret key of  $\text{KEM}$ .
- The encryption operates in a bitwise mode. Let the  $\ell$ -bit message be  $m_1 || \dots || m_\ell$ .
  - When  $m_i = 1$ , we set  $(K_i, \psi_i) \leftarrow \text{KEM.Encap}(pk_{kem})$ .
  - When  $m_i = 0$ , we choose  $(K_i, \psi_i)$  randomly from  $\mathcal{K} \times \mathcal{C}$ .
  - After encrypting  $\ell$  bits, we compute  $F(\psi_1, \dots, \psi_\ell)$  to get  $(K_{y_1}, \dots, K_{y_s})$ , and construct  $s$  extra keys  $K_{\ell+j} = (K_{x_j}, K_{y_j})$  for  $j = 1, \dots, s$ . All  $\ell + s$  keys are then used to compute a tag  $T = \text{XAuth}(K_1, \dots, K_{\ell+s})$ .
  - Finally, the  $\text{PKE}$  ciphertext is  $C = (\psi_1, \dots, \psi_\ell, T)$ .
- The decryption also operates in a bitwise fashion. Omitting some crucial details, we first recompute  $(K_{y_1}, \dots, K_{y_s})$  using  $F$  and  $(\psi_1, \dots, \psi_\ell)$ , reconstruct  $K_{\ell+j}$  for  $j = 1, \dots, s$ , and then verify the correctness of  $T$  using each  $K_{\ell+j}$  in turn. Assuming this step passes, for each  $i$ , we compute  $K_i \leftarrow \text{KEM.Decap}(sk_{kem}, \psi_i)$ , and set the recovered message bit as the output of  $\text{XVer}(K_i, T)$ . (When  $K_i = \perp$ , we set  $\text{XVer}(K_i, T) = 0$ ).

Now, in the above decryption procedure, a KEM decapsulation error occurs whenever  $m_i = 0$ . However,  $\psi_i$  is random in this case, and the tailored decapsulation makes sure that the output of  $\text{KEM.Decap}(sk_{kem}, \psi_i)$  is either  $\perp$  or a random key  $K$ ; in either case,  $\text{XVer}(K_i, T)$  is 0 except with negligible probability because of the security of  $\text{XAC}$  against impersonation attacks.

## 1.4 SIM-SO-CCA Security of Our Main Construction

We follow the techniques of non-committing and deniable encryption [7, 6, 10, 20] and try to create equivocal ciphertexts that not only can be opened arbitrarily but that are also computationally indistinguishable from real ciphertexts. In our construction, the equivocal ciphertexts are in fact encryptions of ones. Note that tCCCA security of KEM ensures that  $(K, \psi) \approx_c (K^R, \psi^R)$ , where  $(K, \psi)$  is the output of  $\text{KEM.Encap}(pk_{kem})$  and  $(K^R, \psi^R)$  is randomly chosen from  $\mathcal{K} \times \mathcal{C}$ . On the other hand, both  $\mathcal{K}$  and  $\mathcal{C}$  are ESE. Therefore,  $(K, \psi)$  encrypting 1 can always be explained as a random pair  $(K^R, \psi^R)$  encrypting 0 by exposing the randomness output from  $\text{SampleK}^{-1}(\mathcal{K}, K)$  and  $\text{SampleC}^{-1}(\mathcal{C}, \psi)$ .

However, this is not sufficient in the SO-CCA setting since the adversary is able to query its decryption oracle and perform corruptions, and it might then be easy for the adversary to distinguish an encryption of ones and an encryption of a real message. For example, consider an adversary that is given a ciphertext  $C = (\psi_1, \psi_2, \dots, \psi_\ell, T)$ , where  $C$  is either an encryption of ones but opened as zeros with re-explained randomness, or an encryption of zeros being opened honestly. In fact, opened randomness exposes all  $K_i$ 's to the adversary. Then the adversary can generate a different ciphertext  $C' = (\psi_1, \psi_2, \dots, \psi_\ell, T')$  as follows. A new tag  $T'$  ( $T' \neq T$ ) is computed as  $T' := \text{XAuth}(K'_1, K_2, \dots, K_{\ell+s})$ , where  $K'_1$  is randomly chosen and all other  $K_i$ 's ( $2 \leq i \leq \ell+s$ ) are the same as in  $T$ . The decryption of  $C'$  will be  $(0, 1, \dots, 1)$  if  $C$  is an encryption of ones but  $(0, 0, \dots, 0)$  if  $C$  is an encryption of zeros! The problem is that the opened randomness discloses  $K_i$  and that gives too much information to the adversary, especially when  $(K_i, \psi_i)$  encodes 0. To solve this problem, we have to use a different method to open  $K_i$  so that the adversary obtains no extra information about  $K_i$  when  $(K_i, \psi_i)$  encodes 0: first, we use algorithm  $\text{ReSample}$  of  $\text{XAC}$  to resample  $K_i$  to obtain a statistically indistinguishable  $\hat{K}_i$ ; then we call  $\text{SampleK}^{-1}(\mathcal{K}, \hat{K}_i)$  and  $\text{SampleC}^{-1}(\mathcal{C}, \psi_i)$  to open  $(\hat{K}_i, \psi_i)$  to an encryption of 0. Now an encryption of ones, say  $C = (\psi_1, \psi_2, \dots, \psi_\ell, T)$ , is able to play the role of an equivocal ciphertext, due to the tCCCA security of KEM and the security of XAC.

Consequently, we can build a simulator  $\mathcal{S}$  with respect to an adversary  $\mathcal{A}$  to prove SIM-SO-CCA security:  $\mathcal{S}$  simulates the real environment for  $\mathcal{A}$  by generating public and private keys, and uses the private key to answer  $\mathcal{A}$ 's decryption queries;  $\mathcal{S}$  creates  $n$  challenge ciphertexts all of which are encryptions of ones; when  $\mathcal{A}$  makes a corruption query concerning a challenge ciphertext  $C$ ,  $\mathcal{S}$  can open  $C$  bit-by-bit according to the real message. If the bit  $m_i$  is 1, it opens  $(K_i, \psi_i)$  honestly, otherwise it opens  $(K_i, \psi_i)$  to 0 by using  $\text{ReSample}$ ,  $\text{SampleC}^{-1}$  and  $\text{SampleK}^{-1}$ .

## 1.5 Related Work

The SOA security notion was first formally proposed by Dwork *et al.* [11]. SIM-SO-CPA and IND-SO-CPA notions were given by Bellare *et al.* [3]. The relations among SOA security notions and traditional IND-CPA security were investigated in [5, 17]. Bellare *et al.* [4] proposed the first SIM-SO-CPA secure Identity-Based Encryption (IBE), while also adopting the non-committing technique and weak deniable encryption. Lai *et al.* [21] proposed the first construction for SIM-SO-CCA secure IBE from a so-called extractable IBE, a collision-resistant hash function, and a strengthened XAC. Recently, Sahai and Waters [23] introduced the puncturable programming technique and employed puncturable PRFs and Indistinguishability Obfuscation ( $i\mathcal{O}$ ) to obtain a variety of cryptographic primitives including deniable encryption with IND-CPA security, PKE with IND-CPA and IND-CCA security, KEM with IND-CCA security, injective trapdoor functions, etc. It should be noted that any IND-CPA secure deniable encryption with ESE ciphertext space implies a PKE with SIM-SO-CPA security. Therefore, the deniable encryption scheme in [23] that is based on a puncturable PRF and  $i\mathcal{O}$  implicitly already gives us a SIM-SO-CPA secure PKE. Our result establishes that SIM-SO-CCA security is achievable from puncturable PRFs and  $i\mathcal{O}$  as well, albeit via the combination of an IND-CCA secure KEM and a strengthened XAC.

## 2 Preliminaries

We use  $s_1, \dots, s_t \leftarrow S$  to denote picking elements  $s_1, \dots, s_t$  uniformly from set  $S$ . Let  $|S|$  denote the size of set  $S$ . Let  $[n]$  denote the set  $\{1, \dots, n\}$ . Let  $s_1 \| s_2 \| \dots$  denotes the concatenation of strings. For a probabilistic polynomial-time (ppt) algorithm  $A$ , we denote  $y \leftarrow A(x; R)$  the process of running  $A$  on input  $x$  with randomness  $R$ , and assigning  $y$  as the result. Let  $\mathcal{R}_A$  denote the randomness space of  $A$ , and  $y \leftarrow A(x)$  denote  $y \leftarrow A(x; R)$  with  $R$  chosen from  $\mathcal{R}_A$  uniformly at random. Let  $U_n$  denote the uniform distribution over  $\{0, 1\}^n$ . A function  $f(\kappa)$  is *negligible*, denoted by  $\text{neg}(\kappa)$ , if for every  $c > 0$  there exists a  $\kappa_c$  such that

$f(\kappa) < 1/\kappa^c$  for all  $\kappa > \kappa_c$ . Let  $\approx_c$  (resp.  $\approx_s$ ) denote computational (resp. statistical) indistinguishability between two ensembles of random variables.

We use boldface letters for vectors. For a vector  $\mathbf{m}$  of finite dimension, let  $|\mathbf{m}|$  denote the length of the vector. For a set  $I = \{i_1, i_2, \dots, i_{|I|}\} \subseteq [|\mathbf{m}|]$ , we define  $\mathbf{m}[I] := (\mathbf{m}[i_1], \mathbf{m}[i_2], \dots, \mathbf{m}[i_{|I|}])$ .

## 2.1 Public Key Encryption

A public key encryption (PKE) scheme is made up of three ppt algorithms:

$\text{KeyGen}(1^\kappa)$  takes as input the security parameter  $\kappa$ , and outputs a public key and a secret key  $(pk, sk)$ .  
 $\text{Enc}(pk, M)$  takes as input the public key  $pk$  and a message  $M$  and outputs a ciphertext  $C$ .  
 $\text{Dec}(sk, C)$  takes as input the secret key  $sk$  and a ciphertext  $C$  and outputs either a message  $M$  or a failure symbol  $\perp$ .

The correctness of a PKE scheme is relaxed to allow a negligible decryption error  $\epsilon(\kappa)$ . That is,  $\text{Dec}(sk, \text{Enc}(pk, M)) = M$  holds with probability at least  $1 - \epsilon(\kappa)$  for all  $(pk, sk) \leftarrow \text{KeyGen}(1^\kappa)$ , where the probability is taken over the coins used in encryption.

Let  $\mathbf{m}$  and  $\mathbf{r}$  be two vectors of dimension  $n := n(\kappa)$ . Define  $\text{Enc}(pk, \mathbf{m}; \mathbf{r}) := (\text{Enc}(pk, \mathbf{m}[1]; \mathbf{r}[1]), \dots, \text{Enc}(pk, \mathbf{m}[n]; \mathbf{r}[n]))$ . Here  $\mathbf{r}[i]$  is the fresh randomness used for the encryption of  $\mathbf{m}[i]$  for  $i \in [n]$ .

## 2.2 Simulation-based, Selective Opening CCA Security of PKE

We review the simulation-based definition of security for PKE against selective opening, chosen-ciphertext adversaries from [12]. Let  $\mathcal{M}$  denote an  $n$ -message sampler, which on input string  $\alpha \in \{0, 1\}^*$  outputs an  $n$ -vector  $\mathbf{m} = (\mathbf{m}[1], \dots, \mathbf{m}[n])$  of messages. Let  $R$  be any ppt algorithm outputting a single bit.

**Definition 1 (SIM-SO-CCA Security)** *A PKE scheme  $\text{PKE} = (\text{KeyGen}, \text{Enc}, \text{Dec})$  is simulation-based, selective opening, chosen-ciphertext secure (SIM-SO-CCA secure) if for every ppt  $n$ -message sampler  $\mathcal{M}$ , every ppt relation  $R$ , every restricted, stateful ppt adversary  $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2, \mathcal{A}_3)$ , there is a stateful ppt simulator  $\mathcal{S} = (\mathcal{S}_1, \mathcal{S}_2, \mathcal{S}_3)$  such that  $\text{Adv}_{\text{PKE}, \mathcal{A}, \mathcal{S}, n, \mathcal{M}, R}^{\text{so-cca}}(\kappa)$  is negligible, where*

$$\text{Adv}_{\text{PKE}, \mathcal{A}, \mathcal{S}, n, \mathcal{M}, R}^{\text{so-cca}}(\kappa) = \left| \Pr \left[ \text{Exp}_{\text{PKE}, \mathcal{A}, n, \mathcal{M}, R}^{\text{so-cca-real}}(\kappa) = 1 \right] - \Pr \left[ \text{Exp}_{\text{PKE}, \mathcal{S}, n, \mathcal{M}, R}^{\text{so-cca-ideal}}(\kappa) = 1 \right] \right|$$

and experiments  $\text{Exp}_{\text{PKE}, \mathcal{A}, n, \mathcal{M}, R}^{\text{so-cca-real}}(\kappa)$  and  $\text{Exp}_{\text{PKE}, \mathcal{S}, n, \mathcal{M}, R}^{\text{so-cca-ideal}}(\kappa)$  are defined in Figure 1. Here the restriction on  $\mathcal{A}$  is that  $\mathcal{A}_2, \mathcal{A}_3$  are not allowed to query the decryption oracle  $\text{Dec}(\cdot)$  with any challenge ciphertext  $\mathbf{c}[i] \in \mathbf{c}$ .

$\text{Exp}_{\text{PKE}, \mathcal{A}, n, \mathcal{M}, R}^{\text{so-cca-real}}(\kappa)$ :	$\text{Exp}_{\text{PKE}, \mathcal{S}, n, \mathcal{M}, R}^{\text{so-cca-ideal}}(\kappa)$ :
$(pk, sk) \leftarrow \text{KeyGen}(1^\kappa)$	$(\alpha, s_1) \leftarrow \mathcal{S}_1(1^\kappa)$
$(\alpha, a_1) \leftarrow \mathcal{A}_1^{\text{Dec}(\cdot)}(pk)$	$\mathbf{m} \leftarrow \mathcal{M}(\alpha)$
$\mathbf{m} \leftarrow \mathcal{M}(\alpha), \mathbf{r} \leftarrow \text{coins}$	$(I, s_2) \leftarrow \mathcal{S}_2(s_1, (1^{ \mathbf{m}[i] })_{i \in [n]})$
$\mathbf{c} \leftarrow \text{Enc}(pk, \mathbf{m}; \mathbf{r})$	$\text{out}_{\mathcal{S}} \leftarrow \mathcal{S}_3(s_2, \mathbf{m}[I])$
$(I, a_2) \leftarrow \mathcal{A}_2^{\text{Dec}_{\mathbf{c}}(\cdot)}(a_1, \mathbf{c})$	return $R(\mathbf{m}, I, \text{out}_{\mathcal{S}})$
$\text{out}_{\mathcal{A}} \leftarrow \mathcal{A}_3^{\text{Dec}_{\mathbf{c}}(\cdot)}(a_2, \mathbf{m}[I], \mathbf{r}[I])$	
return $R(\mathbf{m}, I, \text{out}_{\mathcal{A}})$	

**Fig. 1.** Experiments used in the definition of SIM-SO-CCA security of PKE.

## 2.3 Key Encapsulation Mechanisms

A Key Encapsulation Mechanism (KEM) consists of three ppt algorithms (KEM.Kg, KEM.Enc, KEM.Dec). Let  $\mathcal{K}$  be the key space associated with KEM.

$\text{KEM.Kg}(1^\kappa)$  takes as input a security parameter  $\kappa$  and outputs public/secret key pair  $(pk, sk)$ .

$\text{KEM.Encap}(pk)$  takes as input the public key  $pk$  and outputs a key  $K$  and a ciphertext (or encapsulation)  $\psi$ .

$\text{KEM.Decap}(sk, \psi)$  takes as input the secret key  $sk$  and a ciphertext  $\psi$ , and outputs either a key  $K$  or a failure symbol  $\perp$ .

The correctness condition on a KEM  $\text{KEM}$  is that  $\text{KEM.Decap}(sk, \psi) = K$  holds for all  $\kappa \in \mathbb{N}$ , all  $(pk, sk) \leftarrow \text{KEM.Kg}(1^\kappa)$ , and all  $(K, \psi) \leftarrow \text{KEM.Encap}(pk)$ .

## 2.4 Efficiently Samplable and Explainable (ESE) Domain

A domain  $\mathcal{D}$  is said to be *efficiently samplable and explainable* (ESE) [12] if associated with  $\mathcal{D}$  are the following two ppt algorithms:

$\text{Sample}(\mathcal{D}; R)$ : On input (a description of) domain  $\mathcal{D}$  and random coins  $R \leftarrow \mathcal{R}_{\text{Sample}}$ , this algorithm outputs an element that is uniformly distributed over  $\mathcal{D}$ .

$\text{Sample}^{-1}(\mathcal{D}, x)$ : On input (a description of) domain  $\mathcal{D}$  and *any*  $x \in \mathcal{D}$ , this algorithm outputs  $R$  that is uniformly distributed over the set  $\{R \in \mathcal{R}_{\text{Sample}} \mid \text{Sample}(\mathcal{D}; R) = x\}$ .

Clearly  $\mathcal{D} = \{0, 1\}^\kappa$  is ESE with  $R = \text{Sample}(\mathcal{D}; R) = \text{Sample}^{-1}(\mathcal{D}, R)$ . It was shown by Damgård and Nielsen in [10] that any dense subset of an efficiently samplable domain is ESE as long as the dense subset admits an efficient membership test. Hence, for example,  $\mathbb{Z}_{N^s}^*$  for a RSA modulus  $N$  is ESE.

## 2.5 Cross-Authentication Codes (XACs)

Cross-Authentication Codes (XACs) were first proposed by Fehr et al. in [12] and later adapted to strong XACs in [19] and strengthened XACs in [21].

**Definition 2 (*L*-Cross-Authentication Code [12].)** For  $L \in \mathbb{N}$ , an *L*-cross-authentication code XAC consists of three ppt algorithms ( $X\text{Gen}$ ,  $X\text{Auth}$ ,  $X\text{Ver}$ ) and two associated spaces, the key space  $\mathcal{XK}$  and the tag space  $\mathcal{XT}$  (whose sizes are related to the security parameter  $\kappa$ ). The key generation algorithm  $X\text{Gen}(1^\kappa)$  outputs a uniformly random key  $K \in \mathcal{XK}$ , the authentication algorithm  $X\text{Auth}(K_1, \dots, K_L)$  computes a tag  $T \in \mathcal{XT}$ , and the verification algorithm  $X\text{Ver}(K, T)$  outputs a decision bit.

**Correctness.** For all  $i \in [L]$ , the probability

$$\text{fail}_{\text{XAC}}(\kappa) := \Pr[X\text{Ver}(K_i, X\text{Auth}(K_1, \dots, K_L)) \neq 1],$$

is negligible, where the probability is taken over  $K_1, \dots, K_L \leftarrow \mathcal{XK}$ .<sup>3</sup>

**Security against impersonation and substitution attacks.** Define:

$$\text{Adv}_{\text{XAC}}^{\text{imp}}(\kappa) := \max_{T'} \Pr[X\text{Ver}(K, T') = 1 \mid K \leftarrow \mathcal{XK}]$$

where the max is over all  $T' \in \mathcal{XT}$ , and

$$\text{Adv}_{\text{XAC}}^{\text{sub}}(\kappa) := \max_{i, K_{\neq i}, F} \Pr \left[ \begin{array}{c} T' \neq T \wedge \\ X\text{Ver}(K_i, T') = 1 \end{array} \middle| \begin{array}{c} K_i \leftarrow \mathcal{XK}, \\ T := X\text{Auth}(K_1, \dots, K_L), \\ T' \leftarrow F(T) \end{array} \right]$$

where the max is over all  $i \in [L]$ , all  $K_{\neq i} = (K_j)_{j \in [L], j \neq i} \in \mathcal{XK}^{L-1}$  and all (possibly randomized) functions  $F : \mathcal{XT} \rightarrow \mathcal{XT}$ . Then  $\text{Adv}_{\text{XAC}}^{\text{imp}}(\kappa)$  and  $\text{Adv}_{\text{XAC}}^{\text{sub}}(\kappa)$  are both negligible.

**Definition 3 (Strong and semi-unique XACs.)** An *L*-cross-authentication code XAC is strong and semi-unique if it has the following two properties:

<sup>3</sup> When  $\mathcal{XK} = \mathcal{K}_x \times \mathcal{K}_y$ , the perfect correctness can be relaxed to allow a negligible failure probability  $\text{fail}_{\text{XAC}}(\kappa)$  when only the first components of  $(K_i)_{i \in [L]}$  are randomly chosen.

**Strongness [19]:** Suppose there exists a ppt algorithm  $\text{ReSamp}$ , which takes as input  $(K_j)_{j \neq i}$  and  $T$ , with  $K_1, \dots, K_L \leftarrow \text{XGen}(1^\kappa)$  and  $T \leftarrow \text{XAuth}(K_1, \dots, K_L)$ , and which outputs  $\hat{K}_i$ . We write  $\hat{K}_i \leftarrow \text{ReSamp}(K_{\neq i}, T)$ . Suppose the statistical distance between  $\hat{K}_i$  output by  $\text{ReSamp}$  and  $K_i$ , conditioned on  $(K_{\neq i}, T)$ , is bounded by  $\delta(\kappa)$ , i.e.,

$$\frac{1}{2} \cdot \sum_{k \in \mathcal{XK}} \left| \Pr[\hat{K}_i = k \mid (K_{\neq i}, T)] - \Pr[K_i = k \mid (K_{\neq i}, T)] \right| \leq \delta(\kappa).$$

Then the code  $\text{XAC}$  is said to be  $\delta(\kappa)$ -strong;  $\text{XAC}$  is said to be strong if there exists ppt algorithm  $\text{ReSamp}$  such that  $\delta(\kappa)$  is negligible in  $\kappa$ .

**Semi-Uniqueness [21]:** The code  $\text{XAC}$  is said to be semi-unique if  $\mathcal{XK}$  can be written as  $\mathcal{K}_x \times \mathcal{K}_y$  and if, given  $T \in \mathcal{XT}$  and  $K_x \in \mathcal{K}_x$ , there exists at most one  $K_y \in \mathcal{K}_y$  such that  $\text{XVer}((K_x, K_y), T) = 1$ .

For completeness, we include below the construction of  $\ell$ -cross-authentication codes proposed by Fehr *et al.* [12]. That it is also strong and semi-unique was shown in [21].

- $\mathcal{XK} = \mathcal{K}_a \times \mathcal{K}_b = \mathbb{F}_q^2$  and  $\mathcal{XT} = \mathbb{F}_q^\ell \cup \{\perp\}$ .
- $(a, b) \leftarrow \text{XGen}$ , where  $(a, b) \leftarrow \mathbb{F}_q^2$ .
- $T \leftarrow \text{XAuth}((a_1, b_1), \dots, (a_\ell, b_\ell))$ . Let matrix  $\mathbf{A} \in \mathbb{F}_q^{\ell \times \ell}$  consists of rows  $(1, a_i, a_i^2, \dots, a_i^{\ell-1})$  for  $i \in [\ell]$ . Let  $\mathbf{B}$  be a column vector consisting of  $b_1, \dots, b_\ell \in \mathbb{F}_q$ . If  $\mathbf{AT} = \mathbf{B}$  has no solution or more than one solution, set  $T := \perp$ . Otherwise  $\mathbf{A}$  is a Vandermonde matrix. Let tag  $T = (T_0, \dots, T_{\ell-1})$  be the column vector that is uniquely determined by solving the linear system  $\mathbf{AT} = \mathbf{B}$ .
- Define  $T(x) = T_0 + T_1x + \dots + T_{\ell-1}x^{\ell-1} \in \mathbb{F}_q[x]$  with  $T = (T_0, \dots, T_{\ell-1})$ .  $\text{XVer}((a, b), T)$  outputs 1 if and only if  $T \neq \perp$  and  $T(a) = b$ .
- $(a, b) \leftarrow \text{ReSamp}((a_j, b_j)_{j \neq i, j \in [\ell]}, T)$ . Choose  $a \leftarrow \mathbb{F}_q$  such that  $a \neq a_j$  ( $1 \leq j \leq \ell, j \neq i$ ) and compute  $b := T(a)$ . Conditioned on  $T = \text{XAuth}((a_1, b_1), \dots, (a_\ell, b_\ell))$  ( $T \neq \perp$ ) and  $(a_j, b_j)_{j \neq i}$ , both of  $(a, b)$  and  $(a_i, b_i)$  are uniformly distributed over the same support.
- Any  $a \in \mathbb{F}_q$  uniquely determines  $b := T(a) = T_0 + T_1a + \dots + T_{\ell-1}a^{\ell-1}$  such that  $\text{XVer}((a, b), T) = 1$ .

### 3 KEM Tailored for Construction of PKE with SIM-SO-CCA Security

We describe the properties that are required of a KEM to build SIM-SO-CCA secure PKE; the construction itself is given in the next section.

#### 3.1 Valid Ciphertext Indistinguishability (VCI) of KEMs

Suppose  $\text{KEM} = (\text{KEM.Kg}, \text{KEM.Encap}, \text{KEM.Decap})$  is associated with an efficiently recognizable ciphertext space  $\mathcal{C}$ . For fixed  $\kappa$ , let  $\Psi \subset \mathcal{C}$  denote the set of possible key encapsulations output by  $\text{KEM.Encap}$ , so  $\Psi = \{\psi : \psi \leftarrow \text{KEM.Encap}(pk; r), (pk, sk) \leftarrow \text{KEM.Kg}(1^\kappa), r \leftarrow \text{Coins}\}$ . The set  $\Psi$  is called the *valid ciphertext set* (for  $\kappa$ ).

**Definition 4 (Valid Ciphertext Indistinguishability)** Let  $\text{KEM}$  be a KEM with valid ciphertext set  $\Psi$  and ciphertext space  $\mathcal{C}$ . Define the advantage of an adversary  $\mathcal{A}$  in the experiment depicted in Figure 2 to be

$$\text{Adv}_{\text{KEM}, \mathcal{A}}^{\text{VCI}}(\kappa) := \left| \Pr \left[ \text{Exp}_{\text{KEM}, \mathcal{A}}^{\text{VCI-0}}(\kappa) = 1 \right] - \Pr \left[ \text{Exp}_{\text{KEM}, \mathcal{A}}^{\text{VCI-1}}(\kappa) = 1 \right] \right|.$$

Then  $\text{KEM}$  is said to be Valid Ciphertext Indistinguishable (VCI) if for all ppt adversaries  $\mathcal{A}$ ,  $\text{Adv}_{\text{KEM}, \mathcal{A}}^{\text{VCI}}(\kappa)$  is negligible.

#### 3.2 Tailored KEMs

To be of service in our construction of SIM-SO-CCA secure PKE, we need a KEM that is tailored to have the following three properties, as explained in the introduction: (1) the key space  $\mathcal{K}$  and ciphertext space  $\mathcal{C}$  of the KEM should both be ESE domains; (2) the valid ciphertexts output by the encapsulation algorithm constitute only a small subset of ciphertext space  $\mathcal{C}$ , and the decryption of a random ciphertext results in failure or a random key; (3) the KEM has tailored, constrained CCA security. We define the last of these three properties next.

$\text{Exp}_{\text{KEM}, \mathcal{A}}^{\text{VCI-}b}(\kappa) :$ $(pk, sk) \leftarrow \text{KEM.Kg}(1^\kappa)$ $\psi_0^* \leftarrow \mathcal{C}, (K^*, \psi_1^*) \leftarrow \text{KEM.Encap}(pk)$ $b' \leftarrow \mathcal{A}^{\widetilde{\text{Decap}}_{\neq \psi^*}(\cdot)}(pk, \psi_b^*)$ Return( $b'$ )	$\widetilde{\text{Decap}}_{\neq \psi^*}(P, \psi)$ If $\psi = \psi^*$ return ( $\perp$ ) $K \leftarrow \text{KEM.Decap}(sk, \psi)$ If $P(K) = 0$ return ( $\perp$ ); Else return ( $K$ )
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**Fig. 2.** Experiment for defining Valid Ciphertext Indistinguishability of KEMs. Here  $\widetilde{\text{Decap}}_{\neq \psi^*}(P, \psi)$  denotes a constrained decryption oracle, taking as input predicate  $P(\cdot)$  and encapsulation  $\psi$ .

$\text{Exp}_{\text{KEM}, \mathcal{A}}^{t\text{CCCA-}b}(\kappa) :$ $(pk, sk) \leftarrow \text{KEM.Kg}(1^\kappa)$ $K_0^* \leftarrow \mathcal{K}, \psi_0^* \leftarrow \mathcal{C}$ $(K_1^*, \psi_1^*) \leftarrow \text{KEM.Encap}(pk)$ $b' \leftarrow \mathcal{A}^{\widetilde{\text{Decap}}_{\neq \psi^*}(\cdot)}(pk, K_b^*, \psi_b^*)$ Return( $b'$ )	$\widetilde{\text{Decap}}_{\neq \psi^*}(P, \psi)$ If $\psi = \psi^*$ return ( $\perp$ ) $K \leftarrow \text{KEM.Decap}(sk, \psi)$ If $P(K) = 0$ return ( $\perp$ ); Else return ( $K$ )
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**Fig. 3.** Experiment for defining IND-tCCCA security of KEMs. Here  $\widetilde{\text{Decap}}_{\neq \psi^*}(P, \psi)$  denotes a constrained decryption oracle, taking as input predicate  $P(\cdot)$  and encapsulation  $\psi$ . Predicate  $P(\cdot)$  may vary in different queries.

**Definition 5 (IND-tCCCA Security for KEMs)** Let  $\text{KEM}$  be a KEM with ciphertext space  $\mathcal{C}$  and valid ciphertext set  $\Psi$ , let  $\mathcal{A}$  be a ppt adversary, and consider the experiment  $\text{Exp}_{\text{KEM}, \mathcal{A}}^{t\text{CCCA-}b}(\kappa)$  defined in Figure 3. Define the advantage  $\text{Adv}_{\text{KEM}, \mathcal{A}}^{t\text{CCCA}}(\kappa)$  of  $\mathcal{A}$  by:

$$\text{Adv}_{\text{KEM}, \mathcal{A}}^{t\text{CCCA}}(\kappa) := \left| \Pr \left[ \text{Exp}_{\text{KEM}, \mathcal{A}}^{t\text{CCCA-}0}(\kappa) = 1 \right] - \Pr \left[ \text{Exp}_{\text{KEM}, \mathcal{A}}^{t\text{CCCA-}1}(\kappa) = 1 \right] \right|.$$

Then  $\text{KEM}$  is said to be secure against tailored, constrained chosen ciphertext attacks (IND-tCCCA secure) if for all ppt adversaries  $\mathcal{A}$  with negligible uncertainty  $\text{uncert}_{\mathcal{A}}(\kappa)$  (in  $\kappa$ ), the advantage  $\text{Adv}_{\text{KEM}, \mathcal{A}}^{t\text{CCCA}}(\kappa)$  is also negligible in  $\kappa$ . Here, the uncertainty of  $\mathcal{A}$  is defined as  $\text{uncert}_{\mathcal{A}}(\kappa) := \frac{1}{q_d} \sum_{i=1}^{q_d} \Pr [P_i(K) = 1]$ , which measures the average fraction of keys for which the evaluation of predicate  $P_i(\cdot)$  is equal to 1 in the tCCCA experiment, where  $P_i$  denotes the predicate used in the  $i$ -th query by  $\mathcal{A}$ , and  $q_d$  the number of decapsulation queries made by  $\mathcal{A}$ .

Constrained CCA (CCCA) security for PKE was introduced in [16] as a strictly weaker notion than IND-CCA security. The formal definition of IND-CCA security is included in Appendix A. The main difference between IND-CCCA security and our newly defined IND-tCCCA security is that, in the IND-CCCA definition, the adversary is given a pair  $(K_b^*, \psi^*)$  where  $\psi^*$  is always a correct encapsulation of  $K_1^*$ , while in the IND-tCCCA definition, the adversary is given a pair  $(K_b^*, \psi_b^*)$  where, when  $b = 0$ ,  $\psi_b^*$  is just a random element of  $\mathcal{C}$  and, when  $b = 1$ ,  $\psi_b^*$  is a correct encapsulation of  $K_b^*$ . However, IND-CCCA security and VCI together imply IND-tCCCA security for KEMs:

**Lemma 1.** Suppose that  $\text{KEM}$  is a KEM having an efficiently recognizable ciphertext space  $\mathcal{C}$ . If  $\text{KEM}$  is both IND-CCCA secure and VCI then it is also IND-tCCCA secure.

*Proof.* Recall that the VCI and CCCA experiments are almost the same except for the construction of the adversary's challenge. Let  $(K^{(R)}, \psi^{(R)})$  be chosen from  $\mathcal{C} \times \mathcal{K}$  uniformly at random. Let  $(K, \psi)$  be the output of  $\text{KEM.Encap}$  in the CCCA experiment. IND-CCCA security implies  $(K, \psi) \approx_c (K^{(R)}, \psi)$ . The VCI property implies that  $\psi \approx_c \psi^{(R)}$ , hence  $(K^{(R)}, \psi) \approx_c (K^{(R)}, \psi^{(R)})$  when  $K^{(R)}$  is chosen uniformly and independently of everything else. Finally,  $(K, \psi) \approx_c (K^{(R)}, \psi^{(R)})$  follows from transitivity.  $\square$

**Tailored Decapsulation.** We also tailor the functionality of our KEMs' decapsulation algorithms to suit our PKE construction.

**Definition 6 (Tailored Decapsulation)** Suppose  $\text{KEM} = (\text{KEM.Kg}, \text{KEM.Encap}, \text{KEM.Decap})$  is a KEM. Then  $\text{KEM}$  is said to have tailored decapsulation if there exists a negligible function  $\eta(\kappa)$  such that for all  $(pk, sk)$  output by  $\text{KEM.Kg}(1^\kappa)$ , one or the other of the following two cases pertains:



<p><b>KeyGen</b>(<math>1^\kappa</math>) :</p> <p><math>(pk_{kem}, sk_{kem}) \leftarrow \text{KEM.Kg}(1^\kappa)</math>  <math>K_{x_1}, \dots, K_{x_s} \leftarrow \mathcal{K}_x</math>  <math>pk = (pk_{kem}, (K_{x_j})_{j \in [s]}, F)</math>  <math>sk = (sk_{kem}, pk)</math>.  Return(<math>pk, sk</math>)</p> <p><b>Enc</b>(<math>pk, m_1    \dots    m_\ell</math>) :</p> <p>Parse <math>pk</math> as <math>(pk_{kem}, (K_{x_j})_{j \in [s]}, F)</math>  For <math>i = 1</math> to <math>\ell</math>    If <math>m_i = 1</math>      <math>(K_i, \psi_i) \leftarrow \text{KEM.Encap}(pk_{kem})</math>    Else <math>\psi_i \leftarrow \mathcal{C}; K_i \leftarrow \mathcal{K}</math>  <math>(K_{y_1}, \dots, K_{y_s}) \leftarrow F(\psi_1, \dots, \psi_\ell)</math>  For <math>j = 1</math> to <math>s</math>    <math>K_{\ell+j} \leftarrow (K_{x_j}, K_{y_j})</math>  <math>T \leftarrow \text{XAuth}(K_1, \dots, K_{\ell+s})</math>  Return <math>(\psi_1, \dots, \psi_\ell, T)</math></p>	<p><b>Dec</b>(<math>sk, C</math>) :</p> <p>Parse <math>C</math> as <math>(\psi_1, \dots, \psi_\ell, T)</math>  For <math>i = 1</math> to <math>\ell</math>    <math>m'_i \leftarrow 0</math>  <math>(K'_{y_1}, \dots, K'_{y_s}) \leftarrow F(\psi_1, \dots, \psi_\ell)</math>  For <math>j = 1</math> to <math>s</math>    <math>K'_{\ell+j} \leftarrow (K_{x_j}, K'_{y_j})</math>  If <math>\bigwedge_{j=1}^s \text{XVer}(K'_{\ell+j}, T) = 1</math>    For <math>i = 1</math> to <math>\ell</math>      <math>K'_i \leftarrow \text{KEM.Decap}(sk_{kem}, \psi_i)</math>      If <math>K'_i = \perp</math>, then <math>m'_i \leftarrow 0</math>      Else <math>m'_i \leftarrow \text{XVer}(K'_i, T)</math>  Return(<math>m'_1    m'_2    \dots    m'_\ell</math>)</p>
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**Fig. 4.** Construction of PKE scheme PKE from tailored KEM and  $(\ell + s)$ -XAC.

- $\text{KEM.Decap}$  rejects a random  $\psi' \in \mathcal{C}$ , except with negligible probability, i.e.,

$$\Pr[\text{KEM.Decap}(sk_{kem}, \psi') \neq \perp \mid \psi' \leftarrow \mathcal{C}] \leq \eta(\kappa).$$

- $\text{KEM.Decap}$  outputs  $\eta(\kappa)$ -uniform keys on input a random element from  $\mathcal{C}$ . That is, the statistical distance between the output and a uniform distribution on  $\mathcal{K}$  is bounded by  $\eta(\kappa)$ :

$$\frac{1}{2} \sum_{k \in \mathcal{K}} \left| \Pr[\text{KEM.Decap}(sk_{kem}, \psi') = k \mid \psi' \leftarrow \mathcal{C}] - \frac{1}{|\mathcal{K}|} \right| \leq \eta(\kappa).$$

**Remark.** The former case implies that valid ciphertexts are sparse in the whole ciphertext space, i.e.,  $|\mathcal{V}|/|\mathcal{C}|$  is negligible. In the latter case, VCI (when VCI holds for all  $(pk, sk) \leftarrow \text{KEM.Kg}(1^\kappa)$ ) alone might imply IND-tCCCA security of KEM, since the decapsulated key is uniquely determined by the secret key and the ciphertext (be it valid or invalid).

## 4 Construction of PKE with SIM-SO-CCA Security from Tailored KEMs

Let  $\text{KEM} = (\text{KEM.Kg}, \text{KEM.Encap}, \text{KEM.Decap})$  be a KEM with valid ciphertext set  $\Psi$ , efficiently recognizable ciphertext space  $\mathcal{C}$ , and key space  $\mathcal{K} = \mathcal{K}_x \times \mathcal{K}_y$ . We further assume that:

- (1)  $\text{KEM.Decap}$  has tailored functionality as per Definition 6 (this will be used for the correctness of our PKE construction);
- (2) KEM is IND-tCCCA secure (this will be used in the SIM-SO-CCA security proof of the PKE construction).
- (3) Both the key space  $\mathcal{K}$  and the ciphertext space  $\mathcal{C}$  of KEM are efficiently samplable and explainable domains, with algorithms  $(\text{SampleK}, \text{SampleK}^{-1})$  and  $(\text{SampleC}, \text{SampleC}^{-1})$  (these algorithms are also used in the security analysis).

We refer to a KEM possessing all three properties above as being a *tailored* KEM.

Let  $F : \mathcal{C}^\ell \rightarrow (\mathcal{K}_y)^s$  be an injective function (such functions are easily constructed using, for example, encodings from  $\mathcal{C}$  to bit-strings and from bit-strings to  $\mathcal{K}_y$ , provided  $s$  is sufficiently large). Let  $\text{XAC} = (\text{XGen}, \text{XAuth}, \text{XVer})$  be a  $\delta(\kappa)$ -strong and semi-unique  $(\ell + s)$ -XAC with tag space  $\mathcal{XT}$  and key space  $\mathcal{XK}$ ; suppose also that  $\mathcal{XK} = \mathcal{K} = \mathcal{K}_x \times \mathcal{K}_y$ . Our main construction of PKE scheme  $\text{PKE} = (\text{KeyGen}, \text{Enc}, \text{Dec})$  with message space  $\{0, 1\}^\ell$  is shown in Figure 4.

Note that in the decryption, if  $\text{XVer}(K'_{\ell+j}, T) = 1$  for all  $j \in [s]$ , then the recovered bit  $m'_i$  equals 0 if and only if the decapsulated key  $K'_i$  equals  $\perp$  or  $\text{XVer}(K'_i, T) = 0$ .

<p><b>KeyGen'</b>(<math>1^\kappa</math>) :</p> <p><math>(pk_{kem}, sk_{kem}) \leftarrow \text{KEM.Kg}(1^\kappa)</math>  <math>K_x \leftarrow \mathcal{K}_x, H \leftarrow \text{HGen}(1^\kappa)</math>.  <math>pk = (pk_{kem}, K_x, H)</math>  <math>sk = (sk_{kem}, pk)</math>  Return(<math>pk, sk</math>)</p> <p><b>Enc'</b>(<math>pk, m_1    \dots    m_\ell</math>) :</p> <p>For <math>i = 1</math> to <math>\ell</math>    If <math>m_i = 1</math>      <math>(K_i, \psi_i) \leftarrow \text{KEM.Encap}(pk_{kem})</math>    Else <math>\psi_i \leftarrow \mathcal{C}; K_i \leftarrow \mathcal{K}</math>  <math>K_y \leftarrow H(\psi_1, \dots, \psi_\ell)</math>  <math>K_{\ell+1} \leftarrow (K_x, K_y)</math>  <math>T = \text{XAuth}(K_1, \dots, K_{\ell+1})</math>  Return (<math>\psi_1, \dots, \psi_\ell, T</math>)</p>	<p><b>Dec'</b>(<math>sk, C</math>) :</p> <p><math>C = (\psi_1, \dots, \psi_\ell, T)</math>  For <math>i = 1</math> to <math>\ell</math> <math>m'_i \leftarrow 0</math>  <math>K'_y \leftarrow H(\psi_1, \dots, \psi_\ell)</math>  <math>K'_{\ell+1} \leftarrow (K_x, K'_y)</math>  If <math>\text{XVer}(K'_{\ell+1}, T) = 1</math>    For <math>i = 1</math> to <math>\ell</math>      <math>K'_i \leftarrow \text{KEM.Decap}(sk_{kem}, \psi_i)</math>      If <math>K'_i = \perp</math>, then <math>m'_i \leftarrow 0</math>      Else <math>m'_i \leftarrow \text{XVer}(K'_i, T)</math>  Return(<math>m'_1    m'_2    \dots    m'_\ell</math>)</p>
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**Fig. 5.** Construction of PKE scheme PKE' from tailored KEM,  $(\ell + 1)$ -XAC and CR hash function.

*Correctness.* Encryption and decryption are performed in bitwise fashion. Suppose  $m_i = 1$ . Then  $(K_i, \psi_i)$  are the encapsulated key and corresponding valid encapsulation; by the correctness of KEM and XAC, the decryption algorithm outputs  $m'_i = 1$ , except with negligible probability  $\text{fail}_{\text{XAC}}$ . Suppose  $m_i = 0$ . Then  $K_i$  and  $\psi_i$  are chosen independently and uniformly at random from  $\mathcal{K}$  and  $\mathcal{C}$ , respectively. It follows that the tag  $T$  is independent of  $\psi_i$ . Now, during the decryption of the  $i$ -th bit, according to the tailored property of  $\text{KEM.Decap}$ ,  $K'_i$  is either  $\perp$  (and thus  $m'_i = 0$ ) with probability at least  $1 - \eta(\kappa)$ , or  $K'_i$  is  $\eta(\kappa)$ -close to being uniformly distributed on  $\mathcal{K}$ . In the latter case, it holds that  $m'_i = 0$  except with probability  $\eta(\kappa) + \text{Adv}_{\text{XAC}}^{\text{imp}}(\kappa)$  due to the  $\eta(\kappa)$ -uniformity of the key and the security of XAC against impersonation attack. Consequently, decryption correctly undoes encryption except with probability at most  $\ell \cdot \max\{\text{fail}_{\text{XAC}}(\kappa), \text{Adv}_{\text{XAC}}^{\text{imp}}(\kappa) + \eta(\kappa)\}$ , which is negligible.

**Lemma 2.** *PKE scheme PKE in Figure 4 has the property that, if two distinct ciphertexts  $C, \hat{C}$  both pass the verification step  $\bigwedge_{j=1}^s \text{XVer}(K_{\ell+j}, T) = 1$  during decryption, then they must have different tags  $T \neq \hat{T}$ .*

*Proof.* The proof is by contradiction and relies on the injectivity of  $F$ . Let  $C = (\psi_1, \dots, \psi_\ell, T)$  and  $\hat{C} = (\hat{\psi}_1, \dots, \hat{\psi}_\ell, \hat{T})$  be two different ciphertexts. Let  $(K_{y_1}, \dots, K_{y_s}) = F(\psi_1, \dots, \psi_\ell)$  and  $(\hat{K}_{y_1}, \dots, \hat{K}_{y_s}) = F(\hat{\psi}_1, \dots, \hat{\psi}_\ell)$ . Suppose  $\bigwedge_{j=1}^s \text{XVer}((K_{x_j}, K_{y_j}), T) = \bigwedge_{j=1}^s \text{XVer}((\hat{K}_{x_j}, \hat{K}_{y_j}), \hat{T}) = 1$ . If  $T = \hat{T}$ , then  $C \neq \hat{C}$  implies  $(\psi_1, \dots, \psi_\ell) \neq (\hat{\psi}_1, \dots, \hat{\psi}_\ell)$ , which further implies  $K_{y_j} \neq \hat{K}_{y_j}$  for some  $j \in [s]$ , by the injectivity of  $F$ . On the other hand, we know that  $\text{XVer}((K_{x_j}, K_{y_j}), T) = 1$  and  $\text{XVer}((K_{x_j}, \hat{K}_{y_j}), \hat{T}) = 1$ ; the semi-unique property of XAC now implies that  $K_{y_j} = \hat{K}_{y_j}$ , a contradiction.  $\square$

The SIM-SO-CCA security of PKE will rely on Lemma 2, which in turn relies on the injectivity of  $F$ . The size of  $F$ 's domain is closely related to parameter  $s$ : generally the parameter  $s$  will be linear in  $\ell$ . Since we need a  $(\ell + s)$ -XAC in the construction, the size of public key will be linear in  $\ell$ . The size of tag  $T$  in the ciphertext will also grow linearly in  $s$  and therefore in  $\ell$ . To further decrease the size of public key and tags in our PKE construction, we can employ a collision-resistant (CR) hash function  $H = (\text{HGen}, \text{HEval})$  mapping  $\mathcal{C}^\ell$  to  $\mathcal{K}_y$  instead of the injective function  $F$  (see Appendix B for definitions). Then an  $(\ell + 1)$ -XAC is sufficient for the construction, and this results in more compact public keys and tags, but requires an additional cryptographic assumption. When  $F$  is replaced by a CR hash function  $H$ , Lemma 2 still holds but relies on the existence of collision-resistant (CR) hash function mapping  $\mathcal{C}^\ell$  to a smaller range. We omit the obvious modifications to our PKE construction needed to accommodate the use of CR hash functions. The construction using CR hash functions is given in Figure 5.

**Theorem 1** *Suppose KEM is a tailored KEM, and the  $(\ell + s)$ -cross-authentication code XAC is  $\delta(\kappa)$ -strong, semi-unique, and secure against impersonation and substitution attacks. Then the PKE scheme PKE constructed in Figure 4 is SIM-SO-CCA secure. More precisely, for every ppt adversary  $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2, \mathcal{A}_3)$  against PKE in the SIM-SO-CCA real experiment that makes at most  $q_d$  decryption queries, for every ppt  $n$ -message sampler  $\mathcal{M}$ , and every ppt relation  $R$ , we can construct a stateful ppt simulator  $\mathcal{S} = (\mathcal{S}_1, \mathcal{S}_2, \mathcal{S}_3)$*

for the ideal experiment, and a ppt adversary  $\mathcal{B}$  against the IND-tCCA security of KEM, such that:

$$\text{Adv}_{\text{PKE}, \mathcal{A}, \mathcal{S}, n, \mathcal{M}, R}^{\text{so-cca}}(\kappa) \leq n\ell \cdot \text{Adv}_{\text{KEM}, \mathcal{B}}^{\text{tcca}}(\kappa) + n\ell^2 q_d \cdot \left( \text{Adv}_{\text{XAC}}^{\text{sub}}(\kappa) + \text{Adv}_{\text{XAC}}^{\text{imp}}(\kappa) + \eta(\kappa) \right) + n\ell \cdot \delta(\kappa).$$

First, we give a high level overview of the proof before presenting our formal proof. We construct a ppt simulator  $\mathcal{S}$  as follows. (See Figure 6.)

- $\mathcal{S}$  generates a public/private key pair and provides the public key to  $\mathcal{A}$ .
- $\mathcal{S}$  answers  $\mathcal{A}$ 's decryption queries using the private key.
- $\mathcal{S}$  prepares for  $\mathcal{A}$  a vector of  $n$  challenge ciphertexts, each ciphertext encrypting  $\ell$  ones.
- When  $\mathcal{A}$  decides to corrupt a subset of the challenge ciphertexts,  $\mathcal{S}$  obtains the messages corresponding to the corrupted ciphertexts and opens the corrupted ciphertexts bit-by-bit according to the messages. If bit  $m_i$  should be opened to 1,  $\mathcal{S}$  reveals to  $\mathcal{A}$  the original randomness used by  $\text{KEM.Encap}$  to generate  $(K_i, \psi_i)$ . If bit  $m_i$  should be opened to 0,  $\mathcal{S}$  first explains  $\psi_i$  with randomness output by  $\text{SampleC}^{-1}(\mathcal{C}, \psi_i)$  (as if  $\psi_i$  were randomly chosen). Then  $\mathcal{S}$  uses algorithm  $\text{ReSample}$  of XAC to resample  $K_i$  to get  $\hat{K}_i$ , and explains  $\hat{K}_i$  with randomness output by  $\text{SampleC}^{-1}(\mathcal{K}, \hat{K}_i)$  (as if  $\hat{K}_i$  was randomly chosen).
- $\mathcal{S}$  finally outputs whatever  $\mathcal{A}$  outputs.

The essence of the SIM-SO-CCA security proof is then to show that encryptions of 1's are computationally indistinguishable from encryptions of real messages, even if the adversary can see the opened (real) messages and the randomness of a corrupted subset of the challenge ciphertexts of his/her choice, and have access to the decryption oracle. This is done with a hybrid argument running from Game 0 to Game  $n\ell$ . In Game  $k$  the first  $k$  bits of messages are 1's and are opened as  $\mathcal{S}$  does while the last  $n\ell - k$  bits come from the real messages and are opened honestly. The proof shows that Games  $k$  and  $k - 1$  are indistinguishable using the tCCA security of the tailored KEM and the security properties of the strengthened XAC.

If the  $k$ -th bit of the messages is 1, Games  $k$  and  $k - 1$  are identical. Otherwise, a tailored KEM adversary  $\mathcal{B}$  can be constructed to simulate Game  $k$  or  $k + 1$  for adversary  $\mathcal{A}$ .  $\mathcal{B}$  is provided with a public key  $pk_{k_{\text{em}}}$ , a challenge  $(K^*, \psi^*)$  and a constrained decryption oracle, and is going to tell whether  $(K^*, \psi^*)$  is an output of  $\text{KEM.Encap}(pk_{k_{\text{em}}})$  or a random pair.  $\mathcal{B}$  can generate a public key for  $\mathcal{A}$ . When preparing the vector of challenge ciphertexts,  $\mathcal{B}$  will encrypt the first  $k - 1$  bits from the real messages, use  $(K^*, \psi^*)$  as the encryption of the  $k$ -th bit, and encrypt  $n\ell - k$  ones for the remaining bits. If  $(K^*, \psi^*)$  is an output of  $\text{KEM.Encap}(pk_{k_{\text{em}}})$ , the challenge vector of ciphertexts is just that in Game  $k$ , otherwise it is just that in Game  $k - 1$ . Finally, to answer  $\mathcal{A}$ 's decryption query  $C = (\psi_1, \dots, \psi_\ell, T)$ ,  $\mathcal{B}$  can query  $(\psi_i, \text{XVer}(\cdot, T))$  (note that  $\text{XVer}(\cdot, T)$  is a predicate) to his own constrained decryption oracle if  $\psi^* \neq \psi_i$ ;  $\mathcal{B}$  then replies to  $\mathcal{A}$  with decrypted bit 0 iff  $\mathcal{B}$  gets  $\perp$  from its own oracle. The decryption is correct because  $\mathcal{B}$ 's oracle outputs  $\perp$  iff the decapsulated key is  $K_i = \perp$  or  $\text{XVer}(K_i, T) = 0$ . If  $\psi^* = \psi_i$ ,  $\mathcal{B}$  is not allowed to query his own oracle, but can instead respond to  $\mathcal{A}$  with the output of  $\text{XVer}(K^*, T)$  as the decrypted bit. This decryption is also correct with overwhelming probability for the following reasons: (1) If  $K^*$  is the encapsulated key of  $\psi^*$ , then  $\text{XVer}(K^*, T) = 1$  and decryption is correct. (2) If  $(K^*, \psi^*)$  is a random pair, then all the information leaked about  $K^*$  is just the very tag  $T^*$  that is computed by  $K^*$  during the generation of some challenge ciphertext. The semi-uniqueness of XAC guarantees that  $T \neq T^*$ , and the adversary's corruption only reveals information about a re-sampled  $\hat{K}^*$ . The security of XAC against substitution attacks shows that even if  $\mathcal{A}$  knows  $T^*$  and all keys other than  $K^*$ , then  $\mathcal{A}$  forges a different tag  $T$  such that  $\text{XVer}(K^*, T) = 1$  with negligible probability. Therefore,  $\mathcal{B}$  will almost always respond to  $\mathcal{A}$  with bit 0, which is the correct answer.

*Proof.* (Proof of Theorem 1.) For every ppt  $n$ -message sampler  $\mathcal{M}$ , every ppt relation  $R$ , every stateful ppt adversary  $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2, \mathcal{A}_3)$ , we construct a stateful ppt simulator  $\mathcal{S} = (\mathcal{S}_1, \mathcal{S}_2, \mathcal{S}_3)$  as follows (see also Figure 6).

- $\mathcal{S}_1$  calls  $\text{KeyGen}$  to obtain  $(pk, sk)$ . Then it calls  $\mathcal{A}_1^{\text{Dec}(\cdot)}(pk)$  to obtain  $\alpha$  and the state  $a_1$ . Note that  $\mathcal{S}_1$  possesses  $sk$  and is able to provide a decryption oracle to  $\mathcal{A}_1$ . The view of  $\mathcal{A}_1$  is exactly the same as that in  $\text{Exp}_{\text{PKE}, \mathcal{A}, n, \mathcal{M}, R}^{\text{so-cca-real}}(\kappa)$ .
- Without the knowledge of  $\mathbf{m} = (\mathbf{m}[1], \dots, \mathbf{m}[n])$ , which is the output of  $\mathcal{M}(\alpha)$ ,  $\mathcal{S}_2$  generates the challenge ciphertext vector  $\mathbf{c} = (\mathbf{c}[1], \dots, \mathbf{c}[n])$  with each  $\mathbf{c}[i]$  being an encryption of  $\ell$  ones, i.e.,

$$\mathbf{c}[i] = \text{Enc}(pk, 1^\ell; \mathbf{r}[i]).$$

Then  $\mathcal{S}_2$  calls  $\mathcal{A}_2^{\text{Dec}_{\neq \mathbf{c}}(\cdot)}(a_1, \mathbf{c})$  to get the corruption set  $I$  and state  $a_2$ . Recall that in 'real' experiment  $\text{Exp}_{\text{PKE}, \mathcal{A}, n, \mathcal{M}, R}^{\text{so-cca-real}}(\kappa)$ ,  $\mathcal{A}_2$  receives encryptions of real messages  $\mathbf{m}$ .

$\mathcal{S}_1(1^\kappa):$ $(pk, sk) \leftarrow \text{KeyGen}(1^\kappa)$ $(\alpha, a_1) \leftarrow \mathcal{A}_1^{\text{Dec}(\cdot)}(pk)$ $s_1 \leftarrow (pk, sk, a_1)$ Return( $\alpha, s_1$ )	$\mathcal{S}_2(s_1, (1^{ \mathbf{m}[i] })_{i \in [n]}):$ $pk = (pk_{kem}, K_x, H)$ $sk = (sk_{kem}, pk)$ For $i = 1$ to $n$ For $j = 1$ to $\ell$ $(K_j^{(i)}, \psi_j^{(i)}) \leftarrow \text{KEM.Enc}(pk_{kem}; r_{ij})$ $(K_{y_1}^{(i)}, \dots, K_{y_s}^{(i)}) \leftarrow F(\psi_1^{(i)}, \dots, \psi_\ell^{(i)})$ For $j = 1$ to $s$ $K_{\ell+j}^{(i)} \leftarrow (K_{x_j}, K_{y_j}^{(i)})$ $T^{(i)} \leftarrow \text{XAuth}(K_1^{(i)}, \dots, K_{\ell+s}^{(i)})$ $\mathbf{c}[i] \leftarrow (\psi_1^{(i)}, \dots, \psi_\ell^{(i)}, T^{(i)})$ $\mathbf{r}[i] \leftarrow (r_{i1}, r_{i2}, \dots, r_{i\ell})$ $\mathbf{k}[i] \leftarrow (K_1^{(i)}, \dots, K_{\ell+s}^{(i)})$ $\mathbf{c} \leftarrow (\mathbf{c}[1], \mathbf{c}[2], \dots, \mathbf{c}[n])$ $\mathbf{r} \leftarrow (\mathbf{r}[1], \mathbf{r}[2], \dots, \mathbf{r}[n])$ $\mathbf{k} \leftarrow (\mathbf{k}[1], \mathbf{k}[2], \dots, \mathbf{k}[n])$ $(I, a_2) \leftarrow \mathcal{A}_2^{\text{Dec} \notin \mathcal{C}(\cdot)}(a_1, \mathbf{c})$ $s_2 \leftarrow (a_2, \mathbf{c}, \mathbf{r}, \mathbf{k}, s_1)$ Return( $I, s_2$ )	$\mathcal{S}_3(s_2, \mathbf{m}[I]):$ For $i \in I$ do $\mathbf{c}[i] = (\psi_1^{(i)}, \dots, \psi_\ell^{(i)}, T^{(i)})$ $\mathbf{r}[i] = (r_{i1}, r_{i2}, \dots, r_{i\ell})$ $\mathbf{k}[i] = (K_1^{(i)}, \dots, K_{\ell+s}^{(i)})$ For $j = 1$ to $\ell$ If $\mathbf{m}[i][j] = 1$ $\hat{r}_{ij} \leftarrow r_{ij}$ ; Else $r_c \leftarrow \text{SampleC}^{-1}(\mathcal{C}, \psi_j^{(i)})$ $\hat{K}_j^{(i)} \leftarrow \text{ReSamp}(j, (K_v^{(i)})_{v \neq j}, T^{(i)})$ $r_k \leftarrow \text{SampleK}^{-1}(\mathcal{K}, \hat{K}_j^{(i)})$ $\hat{r}_{ij} \leftarrow (r_c, r_k)$ $K_j^{(i)} \leftarrow \hat{K}_j^{(i)}$ $\hat{\mathbf{r}}[i] \leftarrow (\hat{r}_{i1}, \hat{r}_{i2}, \dots, \hat{r}_{i\ell})$ $out_{\mathcal{A}} \leftarrow \mathcal{A}_3^{\text{Dec} \notin \mathcal{C}(\cdot)}(a_2, \mathbf{m}[I], \hat{\mathbf{r}}[I])$ Return( $out_{\mathcal{A}}$ )
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**Fig. 6.** Construction of simulator  $\mathcal{S}$  for  $\text{Exp}_{\text{PKE}, \mathcal{S}, n, \mathcal{M}, R}^{\text{so-cca-ideal}}(\kappa)$ .

- $\mathcal{S}_3$  opens the challenge ciphertext vector  $(\mathbf{c}[i])_{i \in I}$ , where  $\mathbf{c}[i] = (\psi_1^{(i)}, \dots, \psi_\ell^{(i)}, T^{(i)}) = \text{Enc}(pk, 1^\ell; \mathbf{r}[i])$ , according to the corrupted set of messages  $(\mathbf{m}[i])_{i \in I}$ .
  - If  $\mathbf{m}[i][j] = 1$ ,  $\mathcal{S}_3$  opens with the original random coins;
  - If  $\mathbf{m}[i][j] = 0$ ,  $\mathcal{S}_3$  utilizes  $\text{SampleC}^{-1}$  to recover a properly distributed randomness for  $\psi_j^{(i)}$ , and  $\text{ReSamp}$  to re-sample  $\hat{K}_j^{(i)}$  so as to hide the real key  $K_j^{(i)}$ , and then uses  $\text{SampleK}^{-1}$  to recover a properly distributed randomness for  $\hat{K}_j^{(i)}$ .

Finally,  $\mathcal{S}_3$  collects the newly opened randomness  $\hat{\mathbf{r}}[I]$  and calls  $\mathcal{A}_3^{\text{Dec} \notin \mathcal{C}(\cdot)}(a_2, \mathbf{m}[I], \hat{\mathbf{r}}[I])$  to get the output  $out_{\mathcal{A}}$  as its own output.

The differences between the real and ideal experiments lie in two points. The first is how the challenge ciphertext vector is generated. We formalize it by an algorithm **Generate-c**; the second is how the corrupted ciphertexts are opened, which is formalized by an algorithm **Open-c**. These algorithms are shown in Figure 7. In the proof, we focus on these two points, and proceed with a series of games starting with Game  $G_{-2}$  and ending with Game  $G_{n\ell}$ , with adjacent games being proven to be computationally indistinguishable. The full set of games are illustrated in Figures 7 and 9.

**Game  $G_{-2}$ :** This is the original real experiment  $\text{Exp}_{\text{PKE}, \mathcal{A}, n, \mathcal{M}, R}^{\text{so-cca-real}}(\kappa)$ . Here we elaborate how the challenge ciphertext vector  $\mathbf{c}$  is generated with algorithm **Generate-c** and how the randomness is opened with algorithm **Open-c**. Let  $\mathbf{m} \leftarrow \mathcal{M}(\alpha)$ . In fact,  $\mathbf{c}$  is the output of  $\text{Enc}(pk, \mathbf{m}; \mathbf{r})$  and  $\mathbf{c}[I]$  is opened with the real randomness  $\mathbf{r}[I]$  that was used to compute  $\mathbf{c}[I]$ . So

$$\Pr \left[ \text{Exp}_{\text{PKE}, \mathcal{A}, n, \mathcal{M}, R}^{\text{so-cca-real}}(\kappa) = 1 \right] = \Pr [G_{-2} = 1]. \quad (1)$$

**Game  $G_{-1}$ :** This game is the same as Game  $G_{-2}$  except for the way  $\mathbf{c}[I]$  is opened. If  $\mathbf{m}[i][j] = 0$  for  $i \in I, j \in [\ell]$ , we use  $\text{SampleC}^{-1}$  and  $\text{SampleK}^{-1}$  to obtain the randomness in **Open-c**. Then the original  $r_{ij}$  used to encrypt  $\mathbf{m}[i][j] = 0$  does not need to be kept any more and we can set  $r_{ij}$  to be  $\perp$  in **Generate-c**. This change makes no difference because both  $\mathcal{K}$  and  $\mathcal{C}$  are *efficiently samplable and explainable*. Hence

$$\Pr [G_{-1} = 1] = \Pr [G_{-2} = 1]. \quad (2)$$

**Game  $G_0$ :** This game is the same as Game  $G_{-1}$  except for the way  $\mathbf{c}[I]$  is opened. In case of  $\mathbf{m}[i][j] = 0$  for  $i \in I, j \in [\ell]$ , we do not use the original key  $K_j^{(i)}$ , which was used to create  $T^{(i)}$ , as the input of  $\text{SampleK}^{-1}$ . Instead, we use the  $\text{ReSamp}$  algorithm to re-sample a key  $\hat{K}_j^{(i)}$ , i.e.,  $\hat{K}_j^{(i)} \leftarrow \text{ReSamp}(j, (K_v^{(i)})_{v \neq j}, T^{(i)})$ .

<b>Exp<sub>PKE, A, n, M, R</sub>(<math>\kappa</math>):</b> $(pk, sk) \leftarrow \text{KeyGen}(1^\kappa)$ $(\alpha, a_1) \leftarrow \mathcal{A}_1^{\text{Dec}(\cdot)}(pk)$ $\mathbf{m} \leftarrow \mathcal{M}(\alpha)$ $(\mathbf{c}, \mathbf{r}, \mathbf{k}) \leftarrow \text{Generate-c}(pk, \mathbf{m})$ $(I, a_2) \leftarrow \mathcal{A}_2^{\text{Dec} \neq \text{c}(\cdot)}(a_1, \mathbf{c})$ $\mathbf{r}[I] \leftarrow \text{Open-c}(I, \mathbf{m}, \mathbf{c}, \mathbf{r}, \mathbf{k})$ $out_{\mathcal{A}} \leftarrow \mathcal{A}_3^{\text{Dec} \neq \text{c}(\cdot)}(a_2, \mathbf{m}[I], \mathbf{r}[I])$ return $R(\mathbf{m}, I, out_{\mathcal{A}})$	<b>Open-c(<math>I, \mathbf{m}, \mathbf{c}, \mathbf{r}, \mathbf{k}</math>)</b> <span style="float: right;"><math>G_{-2}</math></span> Return( $\mathbf{r}[I]$ )
<b>Generate-c(<math>pk, \mathbf{m}</math>)</b> <span style="float: right;"><math>G_{-2}</math> <math>G_{-1}, G_0</math></span> $pk = (pk_{kem}, (K_{x_j})_{j \in [s]}, F)$ For $i = 1$ to $n$ For $j = 1$ to $\ell$ If $\mathbf{m}[i][j] = 1$ $(K_j^{(i)}, \psi_j^{(i)}) \leftarrow \text{KEM.Enc}(pk_{kem}; r_{ij})$ Else $K_j^{(i)} \leftarrow \text{SampleK}(\mathcal{K}; r'_{ij})$ $\psi_j^{(i)} \leftarrow \text{SampleC}(\mathcal{C}; r''_{ij})$ $r_{ij} \leftarrow (r'_{ij}, r''_{ij}); \boxed{r_{ij} \leftarrow \perp}$ $(K_{y_1}^{(i)}, \dots, K_{y_s}^{(i)}) \leftarrow F(\psi_1^{(i)}, \dots, \psi_\ell^{(i)})$ For $j = 1$ to $s$ $K_{\ell+j}^{(i)} \leftarrow (K_{x_j}, K_{y_j}^{(i)})$ $T^{(i)} \leftarrow \text{XAuth}(K_1^{(i)}, \dots, K_{\ell+s}^{(i)})$ $\mathbf{c}[i] \leftarrow (\psi_1^{(i)}, \dots, \psi_\ell^{(i)}, T^{(i)})$ $\mathbf{r}[i] \leftarrow (r_{i1}, r_{i2}, \dots, r_{i\ell})$ $\mathbf{k}[i] \leftarrow (K_1^{(i)}, \dots, K_{\ell+s}^{(i)})$ $\mathbf{c} \leftarrow (\mathbf{c}[1], \mathbf{c}[2], \dots, \mathbf{c}[n]); \mathbf{r} \leftarrow (\mathbf{r}[1], \mathbf{r}[2], \dots, \mathbf{r}[n])$ $\mathbf{k} \leftarrow (\mathbf{k}[1], \mathbf{k}[2], \dots, \mathbf{k}[n])$ Return( $\mathbf{c}, \mathbf{r}, \mathbf{k}$ )	<b>Open-c(<math>I, \mathbf{m}, \mathbf{c}, \mathbf{r}, \mathbf{k}</math>)</b> <span style="float: right;"><math>G_{-1}, G_0 \sim G_{n\ell}</math></span> For $i \in I$ do $\mathbf{c}[i] = (\psi_1^{(i)}, \dots, \psi_\ell^{(i)}, T^{(i)})$ $\mathbf{r}[i] = (r_{i1}, r_{i2}, \dots, r_{i\ell})$ $\mathbf{k}[i] \leftarrow (K_1^{(i)}, \dots, K_{\ell+s}^{(i)})$ For $j = 1$ to $\ell$ If $\mathbf{m}[i][j] = 1$ $\hat{r}_{ij} \leftarrow r_{ij}$ ; Else $\hat{K}_j^{(i)} \leftarrow \text{ReSamp}(j, (K_v^{(i)})_{v \neq j}, T^{(i)}); K_j^{(i)} \leftarrow \hat{K}_j^{(i)}$ $r'_{ij} \leftarrow \text{SampleK}^{-1}(\mathcal{K}, K_j^{(i)})$ $r''_{ij} \leftarrow \text{SampleC}^{-1}(\mathcal{C}, \psi_j^{(i)})$ $\hat{r}_{ij} \leftarrow (r'_{ij}, r''_{ij})$ $\hat{\mathbf{r}}[i] \leftarrow (\hat{r}_{i1}, \hat{r}_{i2}, \dots, \hat{r}_{i\ell})$ Return( $\hat{\mathbf{r}}[I]$ )
<b>Dec<sub>c</sub>(<math>sk, (\psi_1, \dots, \psi_\ell, T)</math>)</b> <span style="float: right;"><math>G_{-2} \sim G_{n\ell}</math></span> If $(\psi_1, \dots, \psi_\ell, T) \in \mathcal{C}$ Return( $\perp$ ) $sk = (sk_{kem}, pk)$ For $i = 1$ to $\ell$ $m'_i \leftarrow 0$ $(K'_{y_1}, \dots, K'_{y_s}) \leftarrow F(\psi_1, \dots, \psi_\ell)$ For $j = 1$ to $s$ $K'_{\ell+j} \leftarrow (K_{x_j}, K_{y_j})$ If $\bigwedge_{j=1}^s \text{XVer}(K'_{\ell+j}, T) = 1$ For $i = 1$ to $\ell$ $K'_i \leftarrow \text{KEM.Decap}(sk_{kem}, \psi_i)$ If $K'_i = \perp$ , then $m'_i \leftarrow 0$ Else $m'_i \leftarrow \text{XVer}(K'_i, T)$ Return( $m'_1    m'_2    \dots    m'_\ell$ )	

Fig. 7. Games

The new key  $\hat{K}_j^{(i)}$  is statistically  $\delta(\kappa)$ -close to the real key  $K_j^{(i)}$ , due to the  $\delta(\kappa)$ -strongness of XAC. A union bound gives

$$|\Pr[G_0 = 1] - \Pr[G_{-1} = 1]| \leq (n\ell)\delta(\kappa). \quad (3)$$

**Game  $G_k$ :** Here  $k \in \{0, 1, \dots, n\ell\}$ . Let  $k = (u-1)\ell + w$  with  $u \in [n]$  and  $w \in \{0, 1, \dots, \ell-1\}$ , then each  $k$  determines a unique pair  $(u, w)$ . The difference between  $G_k$  and  $G_0$  is how the challenge ciphertext vector  $\mathbf{c}$  is created in **Generate-c**. There are in total  $n\ell$  bits in  $\mathbf{m}$ ; denote them by  $b_1, b_2, \dots, b_{n\ell}$ . In this game,  $\mathbf{c}$  can be thought of as an encryption of bit-sequence  $1^k b_{k+1}, \dots, b_{n\ell}$ . More precisely, as in Figure 9,

- $(\mathbf{c}[i])_{i \in [u-1]}$  are all encryptions of ones, i.e.,  $\mathbf{c}[i] = \text{Enc}(pk, 1^\ell; \mathbf{r}[i])$  for  $i \in [u-1]$ ;
- For the  $u$ -th ciphertext  $\mathbf{c}_u = (\psi_1^{(u)}, \dots, \psi_\ell^{(u)}, T^{(u)})$ , we have  $(K_j^{(u)}, \psi_j^{(u)}) \leftarrow \text{KEM.Enc}(pk_{kem}; r_{uj})$  for  $j \in \{0, 1, \dots, w\}$ . The remaining  $(K_j^{(u)}, \psi_j^{(u)})$  for  $j \in \{w+1, \dots, \ell-1\}$  are generated according to the value of  $\mathbf{m}[u][j]$  as in Game  $G_0$ . Finally  $T^{(u)}$  is computed as  $T^{(u)} \leftarrow \text{XAuth}(K_1^{(u)}, \dots, K_{\ell+s}^{(u)})$ .
- The remaining ciphertexts  $\mathbf{c}[i]$  for  $i \in \{u+1, \dots, n\}$  are generated according to the value of  $\mathbf{m}[i]$  as in Game  $G_0$ .

When  $k = n\ell$ , the game is exactly the environment provided by our  $\mathcal{S}$  (as described above) in the ideal experiment. All we have to prove is that Game  $G_0$  and  $G_{n\ell}$  are computationally indistinguishable. Further we have:

$$\Pr[G_{n\ell} = 1] - \Pr[G_0 = 1] = \sum_{k=1}^{n\ell} \Pr[G_k = 0] - \Pr[G_{k-1} = 1]. \quad (4)$$

$\mathcal{B}^{\widetilde{\text{Decap}}, \text{LR}}(pk_{kem})$ : $u \leftarrow [n]; w \leftarrow \{0, 1, \dots, \ell - 1\}$ $k \leftarrow (u - 1)\ell + w$ $K_{x_1}, \dots, K_{x_s} \leftarrow \mathcal{K}_x$ $pk = (pk_{kem}, (K_{x_j})_{j \in [s]}, F)$ $K^* \leftarrow \perp; \psi^* \leftarrow \perp$ $(\alpha, a_1) \leftarrow \mathcal{A}_1^{\text{B}, \text{Dec}(\cdot)}(pk)$ $\mathbf{m} \leftarrow \mathcal{M}(\alpha)$ $(\mathbf{c}, \mathbf{r}, \mathbf{k}) \leftarrow \mathbf{G}_k.\text{Generate-c}(pk, \mathbf{m})$ If $\mathbf{m}[u][w] = 0$ $(K^*, \psi^*) \leftarrow \text{LR}(pk_{kem}, b^*)$ $K_w^{(u)} \leftarrow K^*; \psi_w^{(u)} \leftarrow \psi^*$ $(K_{y_1}^{(u)}, \dots, K_{y_s}^{(u)}) \leftarrow F(\psi_1^{(u)}, \dots, \psi_\ell^{(u)})$ For $j = 1$ to $s$ $K_{\ell+j}^{(u)} \leftarrow (K_{x_j}, K_{y_j}^{(u)})$ $T^{(u)} \leftarrow \text{XAuth}(K_1^{(u)}, \dots, K_{\ell+s}^{(u)})$ $\mathbf{c}[u] \leftarrow (\psi_1^{(u)}, \dots, \psi_\ell^{(u)}, T^{(u)})$ $(I, a_2) \leftarrow \mathcal{A}_2^{\text{B}, \text{Dec}_{\neq \mathbf{c}}(\cdot)}(a_1, \mathbf{c})$ $\mathbf{r}[I] \leftarrow \mathbf{G}_k.\text{Open-c}(I, \mathbf{m}, \mathbf{c}, \mathbf{r}, \mathbf{k})$ $out_{\mathcal{A}} \leftarrow \mathcal{A}_3^{\text{B}, \text{Dec}_{\neq \mathbf{c}}(\cdot)}(a_2, \mathbf{m}[I], \mathbf{r}[I])$ return $R(\mathbf{m}, I, out_{\mathcal{A}})$	
$\mathbf{B}.\text{Dec}(\psi_1, \dots, \psi_\ell, T)$ :	$\mathbf{B}.\text{Dec}_{\neq \mathbf{c}}(\psi_1, \dots, \psi_\ell, T)$ :
If $(\psi_1, \dots, \psi_\ell, T) \in \mathbf{c}$ Return( $\perp$ ) For $i = 1$ to $\ell$ $m'_i \leftarrow 0$ $(K'_{y_1}, \dots, K'_{y_s}) \leftarrow F(\psi_1, \dots, \psi_\ell)$ For $j = 1$ to $s$ $K'_{\ell+j} \leftarrow (K_{x_j}, K'_{y_j})$ If $\bigwedge_{j=1}^s \text{XVer}(K'_{\ell+j}, T) = 1$ For $i = 1$ to $\ell$ If $\psi_i = \psi^*$ then $m'_i \leftarrow \text{XVer}(K^*, T)$ Else $K'_i \leftarrow \widetilde{\text{Decap}}(\text{XVer}(\cdot, T), \psi_i)$ if $K_i = \perp$ , then $m'_i \leftarrow 0$ $m'_i \leftarrow \text{XVer}(K'_i, T)$ Return( $m'_1    m'_2    \dots    m'_\ell$ )	
$\widetilde{\text{Decap}}(\text{XVer}(\cdot, T), \psi)$ $K \leftarrow \text{KEM}.\text{Decap}(sk, \psi)$ If $K = \perp$ or $\text{XVer}(K, T) = 0$ return ( $\perp$ ); Otherwise return ( $K$ )	

**Fig. 8.** Algorithm  $\mathcal{B}$ , where  $\mathbf{G}_k.\text{Generate-c}(pk, \mathbf{m})$  (resp.  $\mathbf{G}_k.\text{Open-c}$ ) is algorithm **Generate-c** (resp. **Open-c**) in game  $G_k$ .

$\text{Generate-c}(pk, \mathbf{m})$ <span style="float: right;"><math>G_k = G_{(u,w)}, k \in [n\ell]</math></span> $pk = (pk_{kem}, (\overline{K}_{x_j})_{j \in [s]}, F)$ For $i = 1$ to $u - 1$ For $j = 1$ to $\ell$ $(K_j^{(i)}, \psi_j^{(i)}) \leftarrow \text{KEM}.\text{Enc}(pk_{kem}; r_{ij});$ $(K_{y_1}^{(i)}, \dots, K_{y_s}^{(i)}) \leftarrow F(\psi_1^{(i)}, \dots, \psi_\ell^{(i)})$ For $j = 1$ to $s$ $K_{\ell+j}^{(i)} \leftarrow (K_{x_j}, K_{y_j}^{(i)})$ $T^{(i)} \leftarrow \text{XAuth}(K_1^{(i)}, \dots, K_{\ell+s}^{(i)})$ $\mathbf{c}[i] \leftarrow (\psi_1^{(i)}, \dots, \psi_\ell^{(i)}, T^{(i)})$ $\mathbf{r}[i] \leftarrow (r_{i1}, r_{i2}, \dots, r_{i\ell}); \mathbf{k}[i] \leftarrow (K_1^{(i)}, \dots, K_{\ell+s}^{(i)})$ For $i = u$ do For $j = 1$ to $w$ $(K_j^{(u)}, \psi_j^{(u)}) \leftarrow \text{KEM}.\text{Enc}(pk_{kem}; r_{uj});$ For $j = w + 1$ to $\ell$ If $\mathbf{m}[u][j] = 1$ $(K_j^{(u)}, \psi_j^{(u)}) \leftarrow \text{KEM}.\text{Enc}(pk_{kem}; r_{uj})$ Else $K_j^{(u)} \leftarrow \text{SampleK}(\mathcal{K}; r_{uj});$ $\psi_j^{(u)} \leftarrow \text{SampleC}(\mathcal{C}; r_{uj}'' ); r_{uj} \leftarrow \perp$ $(K_{y_1}^{(u)}, \dots, K_{y_s}^{(u)}) \leftarrow F(\psi_1^{(u)}, \dots, \psi_\ell^{(u)})$ For $j = 1$ to $s$ $K_{\ell+j}^{(u)} \leftarrow (K_{x_j}, K_{y_j}^{(u)})$ $T^{(u)} \leftarrow \text{XAuth}(K_1^{(u)}, \dots, K_{\ell+s}^{(u)})$ $\mathbf{c}[u] \leftarrow (\psi_1^{(u)}, \dots, \psi_\ell^{(u)}, T^{(u)})$ $\mathbf{r}[u] \leftarrow (r_{u1}, r_{u2}, \dots, r_{u\ell})$ $\mathbf{k}[u] \leftarrow (K_1^{(u)}, \dots, K_{\ell+s}^{(u)})$ For $i = u + 1$ to $n$ For $j = 1$ to $\ell$ If $\mathbf{m}[i][j] = 1$ $(K_j^{(i)}, \psi_j^{(i)}) \leftarrow \text{KEM}.\text{Enc}(pk_{kem}; r_{ij})$ Else $K_j^{(i)} \leftarrow \text{SampleK}(\mathcal{K}; r_{ij}');$ $\psi_j^{(i)} \leftarrow \text{SampleC}(\mathcal{C}; r_{ij}'' ); r_{ij} \leftarrow \perp$ $(K_{y_1}^{(i)}, \dots, K_{y_s}^{(i)}) \leftarrow F(\psi_1^{(i)}, \dots, \psi_\ell^{(i)})$ For $j = 1$ to $s$ $K_{\ell+j}^{(i)} \leftarrow (K_{x_j}, K_{y_j}^{(i)})$ $T^{(i)} \leftarrow \text{XAuth}(K_1^{(i)}, \dots, K_{\ell+s}^{(i)})$ $\mathbf{c}[i] \leftarrow (\psi_1^{(i)}, \dots, \psi_\ell^{(i)}, T^{(i)})$ $\mathbf{r}[i] \leftarrow (r_{i1}, r_{i2}, \dots, r_{i\ell}); \mathbf{k}[i] \leftarrow (K_1^{(i)}, \dots, K_{\ell+s}^{(i)})$ $\mathbf{c} \leftarrow (\mathbf{c}[1], \mathbf{c}[2], \dots, \mathbf{c}[n]); \mathbf{r} \leftarrow (\mathbf{r}[1], \mathbf{r}[2], \dots, \mathbf{r}[n])$ $\mathbf{k} \leftarrow (\mathbf{k}[1], \mathbf{k}[2], \dots, \mathbf{k}[n])$ Return( $\mathbf{c}, \mathbf{r}, \mathbf{k}$ )
---

**Fig. 9.** Generation of challenge ciphertext vector in  $G_k$ .

Let  $E_j^{(k)}$  be the event that  $\mathbf{m}[u][w] = 0$  with  $k = (u - 1)\ell + w$  in Game  $G_j$ . Since  $E_j^{(k)}$  is independent of  $j$ , we have  $\Pr[E_0^{(k)}] = \Pr[E_1^{(k)}] = \dots = \Pr[E_{n\ell}^{(k)}]$ . Therefore, we use  $E^{(k)}$  to denote the event that  $\mathbf{m}[u][w] = 1$  for any game. If  $E^{(k)}$  does not happen, i.e.,  $\mathbf{m}[u][w] = 1$  where  $k = (u - 1)\ell + w$ , then the two adjacent Games  $G_k$  and  $G_{k-1}$  are identical. Now:

$$\begin{aligned}
 \Pr[G_k = 1] - \Pr[G_{k-1} = 1] &= (\Pr[G_k = 1 \mid E^{(k)}] \Pr[E^{(k)}] + \Pr[G_k = 1 \mid \neg E^{(k)}] \Pr[\neg E^{(k)}]) \\
 &\quad - (\Pr[G_{k-1} = 1 \mid E^{(k)}] \Pr[E^{(k)}] + \Pr[G_{k-1} = 1 \mid \neg E^{(k)}] \Pr[\neg E^{(k)}]) \\
 &= \left( \Pr[G_k = 1 \mid E^{(k)}] - \Pr[G_{k-1} = 1 \mid E^{(k)}] \right) \Pr[E^{(k)}].
 \end{aligned}$$

Then we have

$$\begin{aligned}
\Pr[G_{n\ell} = 1] - \Pr[G_0 = 1] &= \sum_{k=1}^{n\ell} \Pr[G_k = 0] - \Pr[G_{k-1} = 1] \\
&= \sum_{k=1}^{n\ell} \left( \Pr[G_k = 1 \mid E^{(k)}] - \Pr[G_{k-1} = 1 \mid E^{(k)}] \right) \Pr[E^{(k)}] \\
&= \sum_{k=1}^{n\ell} \left( \Pr[G_k = 1, E^{(k)}] - \Pr[G_{k-1} = 1, E^{(k)}] \right) \tag{5}
\end{aligned}$$

The computational indistinguishability of  $G_0$  and  $G_{n\ell}$  will be proved based on the IND-tCCCA security of KEM, and the security of XAC. Let  $\mathcal{B}$  be an adversary in the IND-tCCCA experiment  $\text{Exp}_{\text{KEM}, \mathcal{B}}^{\text{tccca}-b}(\kappa)$ . Now  $\mathcal{B}$  is provided with a KEM public key  $pk_{kem}$ , and has access to a constrained decapsulation oracle  $\widetilde{\text{Decap}}_{\neq \psi^*}(\cdot)$  and (effectively) a one-time encapsulation oracle  $\text{LR}(pk_{kem}, b)$ . If  $b = 1$ ,  $\text{LR}(pk_{kem}, b)$  will call  $\text{KEM.Encap}(pk_{kem})$  to output  $(K^*, \psi^*)$ , a key and a valid ciphertext encapsulating the key. If  $b = 0$ , it chooses  $(K^*, \psi^*)$  from  $\mathcal{K} \times \mathcal{C}$  uniformly at random.  $\mathcal{B}$  is aiming to tell whether  $b = 0$  or  $b = 1$  after obtaining  $(K^*, \psi^*)$  from  $\text{LR}(pk_{kem}, b)$ , while having access to a constrained decapsulation oracle  $\widetilde{\text{Decap}}_{\neq \psi^*}(\cdot)$ . Now  $\mathcal{B}$  can be constructed as follow (see also Figure 8).

- $\mathcal{B}$  randomly chooses  $u \leftarrow [n]$  and  $w \leftarrow \{0, 1, \dots, \ell - 1\}$ . Setting  $k = (u - 1)\ell + w$ , we have that  $k$  is uniformly distributed in  $[n\ell]$ .
- $\mathcal{B}$  chooses  $K_{x_1}, \dots, K_{x_s} \leftarrow \mathcal{K}_x$ , an injective function  $F : \mathcal{C}^\ell \rightarrow (\mathcal{K}_y)^s$ , and sets  $pk = (pk_{kem}, (K_{x_j})_{j \in [s]}, F)$ .
- $\mathcal{B}$  provides  $\mathcal{A}_1$  with a decryption oracle, denoted by  $\mathcal{B}.\text{Dec}(\cdot)$ , with the help of its own constrained decapsulation oracle  $\widetilde{\text{Decap}}(\cdot)$ . Given a query  $C = (\psi_1, \dots, \psi_\ell, T)$ , algorithm  $\mathcal{B}.\text{Dec}$  verifies  $\bigwedge_{j=1}^s \text{XVer}((K_{x_j}, K'_{y_j}), T) = 1$ , where  $(K'_{y_1}, \dots, K'_{y_s}) = F(\psi_1, \dots, \psi_\ell)$ . If the verification fails,  $\mathcal{B}$  returns  $\ell$  zeros. Otherwise, the decryption goes bitwise as follows.  $\mathcal{B}$  queries its oracle  $\widetilde{\text{Decap}}$  with  $(P, \psi_i)$ , where the predicate  $P$  defined as

$$P(K) := \begin{cases} 1 & \text{if } \text{XVer}(K, T) = 1 \\ 0 & \text{if } \text{XVer}(K, T) = 0 \text{ or } K = \perp \end{cases}$$

If the output of  $\widetilde{\text{Decap}}(\cdot)$  is  $\perp$ ,  $\mathcal{B}$  sets  $m'_i = 0$ , otherwise  $\mathcal{B}$  sets  $m'_i = 1$ . Finally  $\mathcal{B}$  returns  $(m'_1, \dots, m'_\ell)$ . Note that  $\mathcal{B}.\text{Dec}(\cdot)$  functions exactly the same as the decryption oracle  $\text{Dec}(\cdot)$ .

- $\mathcal{B}$  runs  $\mathcal{A}_1^{\text{B.Dec}(\cdot)}(pk)$  to get  $(\alpha, a_1)$  and uses  $\alpha$  to sample a message vector  $\mathbf{m}$ . There are in total  $n\ell$  bits in  $\mathbf{m}$ , denoted  $b_1, b_2, \dots, b_{n\ell}$ . Then  $\mathcal{B}$  computes  $\mathbf{c}$  as follows.
  - if  $\mathbf{m}[u][w] = 1$ ,  $\mathcal{B}$  will encrypt the  $n$  messages consisting of bit sequence  $1^k b_{k+1}, \dots, b_{n\ell}$  to get  $\mathbf{c}$ . In this case,  $\mathbf{c}$  is exactly the output of algorithm **Generate-c** in Game  $G_k$ .
  - if  $\mathbf{m}[u][w] = 0$ ,  $\mathcal{B}$  will obtain  $\mathbf{c}$  by encrypting  $n$  messages which constitute bit sequence  $1^{k-1} b b_{k+1}, \dots, b_{n\ell}$  (notice the presence of  $\mathcal{B}$ 's challenge bit  $b$  here). To do this,  $\mathcal{B}$  queries  $\text{LR}(pk_{kem}, b)$  to get  $(K^*, \psi^*)$  and sets  $K_w^{(u)} := K^*$  and  $\psi_w^{(u)} := \psi^*$  during the generation of  $\mathbf{c}[u]$ . In this way,  $\mathcal{B}$  embeds his own challenge in the creation of  $\mathbf{c}$ . If  $(K^*, \psi^*)$  is the output of  $\text{LR}(pk_{kem}, 1)$ , then  $\mathcal{B}$  implicitly sets  $b_k = b = 1$ . Otherwise  $b_k = b = 0$ . Consequently, the challenge ciphertext vector  $\mathbf{c}$  is just the output of algorithm **Generate-c** in Game  $G_{k-1}$  or  $G_k$  depending on  $b = 0$  or  $b = 1$ .
- $\mathcal{B}$  uses  $\mathcal{B}.\text{Dec}_{\neq \mathbf{c}}(\cdot)$  to answer decryption queries from  $\mathcal{A}_2$ . Algorithm  $\mathcal{B}.\text{Dec}_{\neq \mathbf{c}}(\cdot)$  is almost the same as  $\mathcal{B}.\text{Dec}(\cdot)$  (see Figure 8) but with the restriction that the queried ciphertext  $C = (\psi_1, \dots, \psi_\ell, T)$  does not appear in  $\mathbf{c}$ . Recall that  $\mathcal{B}$  has its own constrained decryption oracle  $\widetilde{\text{Decap}}_{\neq \mathbf{c}}(sk_{kem}, \cdot)$ . This enables  $\mathcal{B}$  to simulate the decryption oracle for  $\mathcal{A}_2$  so long as  $\psi^* \notin \{\psi_1, \dots, \psi_\ell\}$ . However, if  $\psi^* \in \{\psi_1, \dots, \psi_\ell\}$ , say  $\psi_j = \psi^*$ , then  $\mathcal{B}$  computes  $m'_j = \text{XVer}(K^*, T)$  instead. Next it runs  $\mathcal{A}_2^{\text{B.Dec}_{\neq \mathbf{c}}(\cdot)}(a_1, \mathbf{c})$  to obtain the corruption set  $I$  and state  $a_2$ .
- $\mathcal{B}$  calls the **Open-c** algorithm in Game  $G_k$  (which is the same algorithm in all games  $G_0 - G_{n\ell}$ ) to obtain the randomness  $\mathbf{r}[I]$ .
- $\mathcal{B}$  runs  $\mathcal{A}_3^{\text{B.Dec}_{\neq \mathbf{c}}(\cdot)}(a_2, \mathbf{m}[I], \mathbf{r}[I])$  to get the output  $out_{\mathcal{A}}$ .
- $\mathcal{B}$  returns  $R(\mathbf{m}, I, out_{\mathcal{A}})$ .

Let  $\mathcal{B}_k$  denote algorithm  $\mathcal{B}$  running in Game  $G_k$  for a fixed  $k$ . (Recall that  $k = (u - 1)\ell + w$  was chosen from  $[n\ell]$  uniformly at random by  $\mathcal{B}$  at the start of its simulation.) We will show that  $\mathcal{B}_k$  almost perfectly

simulates Game  $G_k$  or  $G_{k-1}$  for  $\mathcal{A}$  depending on the value of  $\mathbf{m}[u][w]$  and  $\mathcal{B}$ 's own challenge  $(K^*, \psi^*)$ . Observe that  $\mathcal{B}.\text{Dec}(\cdot)$  functions exactly the same as  $\text{Dec}(\cdot)$ . If  $\mathbf{m}[u][w] = 1$ ,  $\mathcal{B}_k$  is exactly the same as  $G_k$ . If  $\mathbf{m}[u][w] = 0$ ,  $\mathcal{B}_k$  will be the same as  $G_k$  when  $b = 1$  and  $G_{k-1}$  when  $b = 0$ , as long as  $\mathcal{B}.\text{Dec}_{\neq \mathbf{c}}(\cdot)$  can perfectly simulate the decryption oracle  $\text{Dec}_{\neq \mathbf{c}}(\cdot)$ . However, there is a difference between  $\mathcal{B}.\text{Dec}_{\neq \mathbf{c}}(\cdot)$  and  $\text{Dec}_{\neq \mathbf{c}}(\cdot)$ . Below we show that this difference has negligible influence due to the security properties of XAC.

**Claim 1** *Suppose that  $\mathbf{m}[u][w] = 0$  in  $\mathcal{B}_k$ , where  $k = (u-1)n + w$ . If  $b = 1$ , then the output of algorithm  $\mathcal{B}.\text{Dec}_{\neq \mathbf{c}}(\cdot)$  is exactly the same as that of decryption algorithm  $\text{Dec}_{\neq \mathbf{c}}(\cdot)$ . If  $b = 0$ , algorithm  $\mathcal{B}.\text{Dec}_{\neq \mathbf{c}}(\cdot)$  functions identically to decryption algorithm  $\text{Dec}_{\neq \mathbf{c}}(\cdot)$  except with probability at most*

$$(\ell q_d) \left( \eta(\kappa) + \text{Adv}_{\text{XAC}}^{\text{imp}}(\kappa) + \text{Adv}_{\text{XAC}}^{\text{sub}}(\kappa) \right).$$

*Proof.* Let  $C = (\psi_1, \dots, \psi_\ell, T)$  denote a decryption query from  $\mathcal{A}$ . Clearly  $C \notin \mathbf{c}$ . Let  $(K'_{y_1}, \dots, K'_{y_s}) = F(\psi_1, \dots, \psi_\ell)$ . If  $\bigwedge_{j=1}^s \text{XVer}((K_{x_j}, K'_{y_j}), T) = 0$ , both algorithms  $\text{Dec}_{\neq \mathbf{c}}(\cdot)$  and  $\mathcal{B}.\text{Dec}_{\neq \mathbf{c}}(\cdot)$  will output  $(m'_1, \dots, m'_\ell) = (0, \dots, 0)$ . From now on, we assume that  $\bigwedge_{j=1}^s \text{XVer}((K_{x_j}, K'_{y_j}), T) = 1$ . Then it follows that  $T \neq T^{(i)}$  for all  $i \in [n]$  by Lemma 2.

Next both algorithms  $\text{Dec}_{\neq \mathbf{c}}(\cdot)$  and  $\mathcal{B}.\text{Dec}_{\neq \mathbf{c}}(\cdot)$  will decrypt  $C = (\psi_1, \dots, \psi_\ell, T)$  bitwise to recover the message  $(m'_1, \dots, m'_\ell)$ . Recall that  $\text{Dec}_{\neq \mathbf{c}}(\cdot)$  sets  $m'_i = 0$  iff the decapsulated key  $\text{KEM}.\text{Decap}(\psi_i) = \perp$  or  $\text{XVer}(\text{KEM}.\text{Decap}(\psi_i), T) = 0$ . As for how  $\mathcal{B}.\text{Dec}_{\neq \mathbf{c}}(\cdot)$  determines  $m'_i$ , we consider two cases.

- Case  $\psi^* \neq \psi_i$ :  $\mathcal{B}.\text{Dec}_{\neq \mathbf{c}}(\cdot)$  will query the constrained decapsulation oracle  $\widetilde{\text{Decap}}_{\neq \mathbf{c}}(\cdot)$  with  $(P(\cdot) = \text{XVer}(\cdot, T), \psi_i)$ .  $\mathcal{B}.\text{Dec}_{\neq \mathbf{c}}$  outputs the bit  $m'_i = 0$  iff  $\widetilde{\text{Decap}}_{\neq \mathbf{c}}$  returns  $\perp$ , and this happens when  $\text{KEM}.\text{Decap}(\psi_i)$  outputs  $\perp$  or  $\text{XVer}(\text{KEM}.\text{Decap}(\psi_i), T) = 0$ . Hence, in this case,  $\mathcal{B}.\text{Dec}_{\neq \mathbf{c}}(\cdot)$  functions exactly as  $\text{Dec}_{\neq \mathbf{c}}(\cdot)$  does.
- Case  $\psi^* = \psi_i$ ,  $\mathcal{B}.\text{Dec}_{\neq \mathbf{c}}(\cdot)$  simply computes  $m'_i := \text{XVer}(K^*, T)$ . We further consider two sub-cases.
  - Case  $b = 1$ : Then  $(K^*, \psi^*)$  is the output of  $\mathcal{B}$ 's encryption oracle  $\mathbf{LR}(pk_{kem}, 1)$ , thus  $K^*$  is the encapsulated key of  $\psi^*$ , i.e.  $K^* = \text{KEM}.\text{Decap}(sk_{kem}, \psi^*)$ . Consequently, the decrypted bit  $m'_i := \text{XVer}(K^*, T)$  is exactly the same as the bit  $m'_i$  output by  $\text{Dec}_{\neq \mathbf{c}}(\cdot)$ .
  - Case  $b = 0$ : Then  $(K^*, \psi^*)$  is the output of  $\mathcal{B}$ 's encryption oracle  $\mathbf{LR}(pk_{kem}, 0)$ . Hence both  $K^*$  and  $\psi^*$  are uniformly chosen. In this case,  $\text{Dec}_{\neq \mathbf{c}}(\cdot)$  outputs  $m'_i = 0$  except with probability  $\eta(\kappa) + \text{Adv}_{\text{XAC}}^{\text{imp}}(\kappa)$ , according to the decryption correctness of PKE.

On the other hand,  $\mathcal{B}.\text{Dec}_{\neq \mathbf{c}}(\cdot)$  outputs  $m'_i := \text{XVer}(K^*, T)$ . Recall that  $T \neq T^{(u)}$ , so  $\text{XVer}(K^*, T) = 0$  except with probability  $\text{Adv}_{\text{XAC}}^{\text{sub}}(\kappa)$ , due to the security of XAC against substitution attacks. More precisely,  $K^*$  is uniformly chosen from  $\mathcal{K}$ , and all information leaked about  $K^*$  is given by  $T^{(u)} = \text{XAuth}(K_1^{(u)}, \dots, K_w^{(u)} = K^*, \dots, K_\ell^{(u)})$  in the challenge ciphertext  $\mathbf{c}[u]$  and  $\hat{K}_w^{(u)} = \text{ReSamp}(w, (K_v^{(u)})_{v \in [\ell+s], v \neq w}, T^{(u)})$  by the **Open-c** algorithm. However, algorithm **Open-c** did not expose the real key  $K_w^{(u)} = K^*$  and what it revealed is a fake key  $\hat{K}_w^{(u)}$ . The fake key is the output of **ReSamp** on input  $(K_v^{(u)})_{v \neq w}$  and  $T^{(u)}$ , hence is redundant and does not leak more information beyond  $(K_v^{(u)})_{v \neq w}$  and  $T^{(u)}$ . Consequently, the security of XAC against substitution attacks shows that  $\text{XVer}(K^*, T) = 0$  except with probability  $\text{Adv}_{\text{XAC}}^{\text{sub}}(\kappa)$ , even conditioned on the knowledge of  $T^{(u)}$  and  $(K_v^{(u)})_{v \neq w}$ .

Hence  $\mathcal{B}.\text{Dec}_{\neq \mathbf{c}}(\cdot)$  outputs an  $m'_i$  different from that output by  $\text{Dec}_{\neq \mathbf{c}}(\cdot)$  with probability at most  $\eta(\kappa) + \text{Adv}_{\text{XAC}}^{\text{imp}}(\kappa) + \text{Adv}_{\text{XAC}}^{\text{sub}}(\kappa)$ . The lemma follows by applying the union bound, since the plaintext has  $\ell$  bits and there are  $q_d$  decryption queries. (End of proof of Claim 1.)

Let  $\text{Err}_k$  denote the event that there exists a decryption query  $C$  to which  $\mathcal{B}.\text{Dec}_{\neq \mathbf{c}}(\cdot)$  replies differently from  $\text{Dec}_{\neq \mathbf{c}}(\cdot)$  in  $\mathcal{B}_k$ . According to the proof of Claim 1, we know that  $\text{Err}_k$  occurs only when  $b = 0$  and  $\mathbf{m}[u][w] = 0$ ; moreover

$$\Pr[\text{Err}_k \mid \mathbf{m}[u][w] = 0, b = 0] \leq (\ell q_d) \left( \eta(\kappa) + \text{Adv}_{\text{XAC}}^{\text{imp}}(\kappa) + \text{Adv}_{\text{XAC}}^{\text{sub}}(\kappa) \right). \quad (6)$$



Let  $\mathcal{B} = b'$  (respectively  $\mathcal{B}_k = b'$ ) denote that algorithm  $\mathcal{B}$  (respectively  $\mathcal{B}_k$ ) returns bit  $b'$ . Let  $E^{(k)}$  denote the event that  $\mathbf{m}[u][w] = 0$  with  $k = (u-1)\ell + w$  in  $\mathcal{B}_k$ . The advantage of  $\mathcal{B}$  is given by

$$\begin{aligned} \text{Adv}_{\text{KEM}, \mathcal{B}}^{\text{tccca}}(\kappa) &= \Pr[\mathcal{B} = 1 \mid b = 1] - \Pr[\mathcal{B} = 1 \mid b = 0] \\ &= \frac{1}{n\ell} \sum_{k=1}^{n\ell} \left( \Pr[\mathcal{B}_k = 1, E^{(k)} \mid b = 1] - \Pr[\mathcal{B}_k = 1, E^{(k)} \mid b = 0] \right) \end{aligned} \quad (7)$$

$$= \frac{1}{n\ell} \Pr[E^{(k)}] \sum_{k=1}^{n\ell} \left( \Pr[\mathcal{B}_k = 1 \mid E^{(k)}, b = 1] - \Pr[\mathcal{B}_k = 1 \mid E^{(k)}, b = 0] \right) \quad (8)$$

$$\begin{aligned} &= \frac{\Pr[E^{(k)}]}{n\ell} \sum_{k=1}^{n\ell} \left( \Pr[G_k = 1 \mid E^{(k)}] - \Pr[\mathcal{B}_k = 1, \overline{\text{Err}}_k \mid E^{(k)}, b = 0] \right. \\ &\quad \left. - \Pr[\mathcal{B}_k = 1, \text{Err}_k \mid E^{(k)}, b = 0] \right) \end{aligned} \quad (9)$$

$$\begin{aligned} &= \frac{1}{n\ell} \sum_{k=1}^{n\ell} \Pr[G_k = 1, E^{(k)}] \\ &\quad - \frac{1}{n\ell} \Pr[E^{(k)}] \sum_{k=1}^{n\ell} \Pr[\mathcal{B}_k = 1 \mid \overline{\text{Err}}_k, E^{(k)}, b = 0] \cdot \Pr[\overline{\text{Err}}_k \mid E^{(k)}, b = 0] \\ &\quad - \frac{1}{n\ell} \Pr[E^{(k)}] \sum_{k=1}^{n\ell} \Pr[\mathcal{B}_k = 1 \mid \text{Err}_k, E^{(k)}, b = 0] \cdot \Pr[\text{Err}_k \mid E^{(k)}, b = 0] \\ &= \frac{1}{n\ell} \sum_{k=1}^{n\ell} \left( \Pr[G_k = 1, E^{(k)}] - \Pr[G_{k-1} = 1, E^{(k)}] \cdot \Pr[\overline{\text{Err}}_k \mid E^{(k)}, b = 0] \right) \\ &\quad - \frac{1}{n\ell} \sum_{k=1}^{n\ell} \Pr[E^{(k)}] \cdot \Pr[\mathcal{B}_k = 1 \mid \text{Err}_k, E^{(k)}, b = 0] \cdot \Pr[\text{Err}_k \mid E^{(k)}, b = 0] \end{aligned} \quad (10)$$

$$\geq \frac{1}{n\ell} \sum_{k=1}^{n\ell} \left( \Pr[G_k = 1, E^{(k)}] - \Pr[G_{k-1} = 1, E^{(k)}] \right) - \max_k \Pr[\text{Err}_k \mid E^{(k)}, b = 0] \quad (11)$$

Eqn. (7) follows from the fact that the output of  $\mathcal{B}_k$  is independent of  $b$  if  $E^{(k)}$  does not occur. Eqn. (8) follows from that  $E^{(k)}$  and  $b$  are independent of each other. Eqn. (9) follows from the fact that  $\mathcal{B}_k$  perfectly simulates  $G_k$  if  $b = 1$  and  $E^{(k)}$  occurs. Eqn. (10) is due to the fact that  $\mathcal{B}_k$  perfectly simulates  $G_{k-1}$  when  $b = 0$  and  $\text{Err}_k$  does not occur. According to (5), (6) and (11), we have

$$\begin{aligned} \Pr[G_{n\ell} = 1] - \Pr[G_0 = 1] &= \sum_{k=1}^{n\ell} \left( \Pr[G_k = 1, E^{(k)}] - \Pr[G_{k-1} = 1, E^{(k)}] \right) \\ &\leq n\ell \cdot \text{Adv}_{\text{KEM}, \mathcal{B}}^{\text{tccca}}(\kappa) + n\ell^2 q_d \cdot \eta(\kappa) + n\ell^2 q_d \cdot \text{Adv}_{\text{XAC}}^{\text{imp}}(\kappa) + n\ell^2 q_d \cdot \text{Adv}_{\text{XAC}}^{\text{sub}}(\kappa) \end{aligned} \quad (12)$$

Combining (1), (2), (3) and (12), we have

$$\begin{aligned} \text{Adv}_{\text{PKE}, \mathcal{A}, \mathcal{S}, n, \mathcal{M}, R}^{\text{so-cca}}(\kappa) &= |\Pr[G_{n\ell} = 1] - \Pr[G_{-2} = 1]| \\ &\leq |\Pr[G_{n\ell} = 1] - \Pr[G_0 = 1]| + |\Pr[G_0 = 1] - \Pr[G_{-2} = 1]| \\ &\leq n\ell \cdot \text{Adv}_{\text{KEM}, \mathcal{B}}^{\text{tccca}}(\kappa) + n\ell^2 q_d \cdot \text{Adv}_{\text{XAC}}^{\text{sub}}(\kappa) + n\ell^2 q_d \cdot \text{Adv}_{\text{XAC}}^{\text{imp}}(\kappa) + n\ell^2 q_d \cdot \eta(\kappa) + n\ell \cdot \delta(\kappa) \end{aligned} \quad (13)$$

from which the theorem follows.  $\square$

The security of our modified construction using CR hash functions (see Figure 5) is stated in the following theorem, whose proof is similar to that of Theorem 1.

**Theorem 2** *Suppose KEM is a tailored KEM, the  $(\ell+1)$ -cross-authentication code XAC is  $\delta(\kappa)$ -strong, semi-unique, and secure against impersonation and substitution attacks, and  $H$  is collision-resistant. Then the PKE scheme PKE' constructed in Figure 5 is SIM-SO-CCA secure. More precisely, for every ppt adversary  $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2, \mathcal{A}_3)$  against PKE' in the SIM-SO-CCA real experiment that makes at most  $q_d$  decryption*

queries, for every ppt  $n$ -message sampler  $\mathcal{M}$ , and every ppt relation  $R$ , we can construct a stateful ppt simulator  $\mathcal{S} = (\mathcal{S}_1, \mathcal{S}_2, \mathcal{S}_3)$  for the ideal experiment, a ppt adversary  $\mathcal{B}$  against the IND-tCCA security of KEM, and a ppt algorithm  $\mathcal{F}$  against the collision-resistance of  $H$  such that:

$$\begin{aligned} \text{Adv}_{\text{PKE}, \mathcal{A}, \mathcal{S}, n, \mathcal{M}, R}^{\text{so-cca}}(\kappa) &\leq n\ell \cdot \text{Adv}_{\text{KEM}, \mathcal{B}}^{\text{tcca}}(\kappa) + n\ell^2 q_d \cdot \left( \text{Adv}_{\text{XAC}}^{\text{sub}}(\kappa) + \text{Adv}_{\text{XAC}}^{\text{jmp}}(\kappa) + \eta(\kappa) \right) \\ &\quad + n\ell \cdot \delta(\kappa) + \text{Adv}_{H, \mathcal{F}}^{\text{cr}}(\kappa). \end{aligned}$$

## 5 Instantiations

In this section, we explore three different constructions of tailored KEMs, each suitable for the application of Theorems 1 and 2. The first is based on any Strongly Universal<sub>2</sub> hash proof system, the second is a direct construction relying on the  $n$ -Linear Assumption and a target collision-resistant hash function (see Appendix B for definitions), while the third uses indistinguishability obfuscation.

### 5.1 Strongly Universal<sub>2</sub> Hash Proof Systems

We use hash proof systems [8] to build tailored KEMs suitable for application in our main theorem.

Let  $\Psi \subset \mathcal{C}$  be a language. The hardness of the *subset membership problem* for  $\Psi$  with respect to  $\mathcal{C}$  requires that a random element from  $\Psi$  is indistinguishable from a random element from  $\mathcal{C}$ . Let  $\mathcal{K}$  be a set and  $\Lambda_{sk} : \mathcal{C} \rightarrow \mathcal{K}$  be a hash function indexed with  $sk \in \mathcal{SK}$ . Then  $\Lambda_{sk}$  is said to be *projective* if there exists a map  $\mu : \mathcal{SK} \rightarrow \mathcal{PK}$  such that  $\mu(sk) \in \mathcal{PK}$  defines the action of  $\Lambda_{sk}$  on the subset  $\Psi$ ;  $\mu$  is then said to be a *projection* on subset  $\Psi$ .

A hash proof system (HPS) HPS consists of three algorithms (HPS.param, HPS.pub, HPS.priv). The randomized algorithm HPS.param( $1^\kappa$ ) outputs  $\text{params} = (\mathbb{G}, \mathcal{C}, \Psi, \mathcal{PK}, \mathcal{SK}, \Lambda, \mu)$ , where  $\mathbb{G}$  is a group. The secret key  $sk$  is randomly chosen from  $\mathcal{SK}$ , and the public key is computed as  $pk = \mu(sk)$  where  $\mu$  is a projection on  $\Psi$ . Algorithm HPS.Pub( $pk, \psi, w$ ) is given the public key  $pk$ , an element  $\psi \in \Psi$  and its witness  $w$ , and outputs an encapsulated key  $K = \text{HPS.Pub}(pk, \psi, w)$  such that  $K = \Lambda_{sk}(\psi)$ . Algorithm HPS.Priv( $sk, \psi$ ) recovers  $K = \Lambda_{sk}(\psi)$  using  $sk$ .

The *Strongly Universal<sub>2</sub>* ( $SU_2$ ) property of an HPS characterizes the unpredictability of  $\Lambda_{sk}(\psi)$  for  $\psi \in \mathcal{C} \setminus \Psi$ .

**Definition 7** Let  $\text{HPS} = (\text{HPS.param}, \text{HPS.pub}, \text{HPS.priv})$  be a hash proof system. Then HPS is said to be  $SU_2$  if

$$\Pr[\Lambda_{sk}(\psi) = K \mid pk = \mu(sk), \psi', K' = \Lambda_{sk}(\psi')] = 1/|\mathcal{K}|,$$

for all  $pk \in \mathcal{PK}$ , all  $\psi, \psi' \in \mathcal{C} \setminus \Psi$  with  $\psi' \neq \psi$  and all  $K, K' \in \mathcal{K}$ , where the probability is taken over  $sk \leftarrow \mathcal{SK}$ .

Given that HPS is an  $SU_2$  HPS, a KEM KEM can be constructed as shown in Figure 10. The output  $\text{params}$  of HPS.param is used as a set of public parameters implicitly used as input in the algorithms of KEM. Notice that the valid ciphertext set for KEM is  $\Psi$ .

$\text{KEM.Kg}(1^\kappa):$ $sk \leftarrow \mathcal{SK}$ $pk = \mu(sk)$ Return $(pk, sk)$	$\text{KEM.Encap}(pk):$ $\psi \leftarrow \Psi$ with witness $w$ $K \leftarrow \text{HPS.Pub}(pk, \psi, w)$ Return $(K, \psi)$	$\text{KEM.Decap}(sk, \psi):$ $K \leftarrow \text{HPS.Priv}(sk, \psi)$ Return $(K)$
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**Fig. 10.** Construction of a KEM from an  $SU_2$  hash proof system.

**Theorem 3** Let HPS be an  $SU_2$  HPS with  $\text{params} = (\mathbb{G}, \mathcal{C}, \Psi, \mathcal{PK}, \mathcal{SK}, \Lambda, \mu)$ . Suppose the subset membership problem is hard for  $\Psi$  with respect to  $\mathcal{C}$ . Then the KEM KEM constructed from HPS as shown in Figure 10 is IND-tCCA secure. Furthermore, if  $\Psi$  is sparse in  $\mathcal{C}$ , and both  $\mathcal{C}$  and  $\mathcal{K}$  are efficiently samplable and explainable, then KEM is a tailored KEM.

*Proof.* It was already proved in [16] that the  $SU_2$  property and the hardness of the subset membership problem for  $\Psi$  with respect to  $\mathcal{C}$  implies the IND-CCCA security of KEM. On the other hand, public and secret key pairs can be generated independently from  $\mathcal{C}$  and  $\Psi$  and the subset membership problem holds even if the secret key is known to the adversary. More precisely, when an adversary  $\mathcal{B}$  is given  $\psi$  and tries to distinguish whether  $\psi$  is randomly chosen from  $\Psi$  or  $\mathcal{C}$ , it can establish a VCI experiment for a VCI adversary  $\mathcal{A}$  as follows: first call  $(pk, sk) \leftarrow \text{KEM.Kg}(1^\kappa)$  and use  $sk$  to answer decryption queries.  $\mathcal{B}$  gives  $pk$  to  $\mathcal{A}$  and gives  $\psi$  as the challenge ciphertext. Finally  $\mathcal{B}$  outputs whatever  $\mathcal{A}$  returns. It is clear that  $\mathcal{B}$  has the same advantage as  $\mathcal{A}$ . This implies that the VCI property holds for KEM under the hardness of the subset membership problem. Then IND-tCCCA security follows from Lemma 1.

The  $SU_2$  property of HPS implies that

$$\Pr[\text{KEM.Decap}(sk, \psi) = K] = \Pr[\text{HPS.Priv}(sk, \psi) = K] = \frac{1}{|\mathcal{K}|}$$

for all invalid ciphertexts  $\psi \in \mathcal{C} \setminus \Psi$ , all  $K \in \mathcal{K}$ , and all  $pk = \mu(sk)$ , where the probability is taken over  $sk \leftarrow \mathcal{SK}$ . Then

$$\begin{aligned} \Pr[\text{KEM.Decap}(sk, \psi) = K \mid \psi \leftarrow \mathcal{C}] &= \Pr[\text{KEM.Decap}(sk, \psi) = K \mid \psi \in \Psi] \cdot \frac{|\Psi|}{|\mathcal{C}|} \\ &+ \Pr[\text{KEM.Decap}(sk, \psi) = K \mid \psi \in \mathcal{C} \setminus \Psi] \cdot \left(1 - \frac{|\Psi|}{|\mathcal{C}|}\right) \\ &= \Pr[\text{KEM.Decap}(sk, \psi) = K \mid \psi \in \Psi] \cdot \frac{|\Psi|}{|\mathcal{C}|} + \frac{1}{|\mathcal{K}|} \cdot \left(1 - \frac{|\Psi|}{|\mathcal{C}|}\right) \leq \frac{|\Psi|}{|\mathcal{C}|} + \frac{1}{|\mathcal{K}|}. \end{aligned}$$

Noting that  $\Pr[\text{KEM.Decap}(sk, \psi) = K \mid \psi \in \Psi]$  lies between 0 and 1, it follows that the statistical distance between  $\text{KEM.Decap}(sk, \psi)$  (when  $\psi$  is uniformly selected from  $\mathcal{C}$ ) and the uniform distribution is at most  $|\Psi|/|\mathcal{C}|$ , which is negligible due to the sparseness of  $\Psi$ . This establishes that  $\text{KEM.Decap}$  has tailored functionality.

Finally, KEM is a tailored KEM because it has samplable and explainable domains  $\mathcal{C}$  and  $\mathcal{K}$ , it has IND-tCCCA security, and  $\text{KEM.Decap}$  has tailored functionality.  $\square$

*Remark 1.* As pointed out in [12], both DDH-based and DCR-based HPS could have samplable and explainable platform groups. For example, we can choose the subgroup of order  $q$  in  $\mathbb{Z}_p^*$  (with  $p = 2q + 1$ ) as the DDH group, and choose  $\mathbb{Z}_{N^2}^*$  as the DCR group.

## 5.2 Tailored KEM Based on $n$ -Linear Assumption

Let  $\mathcal{G}(1^\kappa)$  be a group generator, that is, a ppt algorithm which outputs  $(\mathbb{G}, g, p)$  where  $\mathbb{G}$  is a group of prime order  $p$  (having  $\kappa$  bits) and  $g$  a generator of  $\mathbb{G}$ .

**Definition 8** *The  $n$ -Linear Assumption for  $\mathcal{G}(1^\kappa)$  states that for all ppt adversaries  $\mathcal{B}$ , the advantage of  $\mathcal{B}$  defined below is negligible.*

$$\begin{aligned} \text{Adv}_{\mathcal{B}}^{n\text{-lin}}(\kappa) &:= \left| \Pr \left[ \mathcal{B}(g_1, \dots, g_n, g_1^{r_1}, \dots, g_n^{r_n}, h, h^{\sum_{i=1}^n r_i}) = 1 \right] \right. \\ &\quad \left. - \Pr \left[ \mathcal{B}(g_1, \dots, g_n, g_1^{r_1}, \dots, g_n^{r_n}, h, h^z) = 1 \right] \right|, \end{aligned}$$

where  $(\mathbb{G}, g, p) \leftarrow \mathcal{G}(1^\kappa)$ ,  $(g_i)_{i \in [n]}, h \leftarrow \mathbb{G}$  and  $(r_i)_{i \in [n]}, z \leftarrow \mathbb{Z}_p$ .

In [16], Hofheiz and Kiltz presented a KEM based on the  $n$ -Linear Assumption for a group generator  $\mathcal{G}(1^\kappa)$  and a target collision-resistant hash function, and proved its IND-CCCA security. We replicate the algorithms of this KEM in Figure 11. Note that this construction does not fall into the category of HPS-based KEMs.

**Lemma 3.** *If the  $n$ -Linear Assumption holds for  $\mathcal{G}(1^\kappa)$ , and TCR is target collision-resistant, then the Hofheiz-Kiltz KEM in Fig. 11 is IND-tCCCA secure.*

*Proof.* In view of the results of [16] and Lemma 1, we need only prove that the KEM in Figure 11 has the VCI property.

Given an adversary  $\mathcal{A}$  winning the VCI experiment with non-negligible probability, we can construct a ppt algorithm  $\mathcal{B}$  solving the  $n$ -Linear problem with help of  $\mathcal{A}$  with non-negligible probability. Let  $(g_1, \dots, g_n, g_1^{r_1}, \dots, g_n^{r_n}, h, K^*)$  be a challenge instance from the  $n$ -Linear problem, where  $K^* = h^{\sum_{i=1}^n r_i}$  or  $K^*$  is a random element from  $\mathbb{G}$ . Here,  $\mathcal{B}$  simulates the VCI experiment for  $\mathcal{A}$  using its input  $(g_1^{r_1}, \dots, g_n^{r_n}, h, K^*)$ .

KEM.Kg( $1^\kappa$ ):	KEM.Encap( $pk$ ):	KEM.Decap( $sk, \psi$ ):
$b \leftarrow \mathbb{Z}_p; h \leftarrow g^b$ For $i = 1$ to $n$ $\alpha_i, \alpha_i, \beta_i \leftarrow \mathbb{Z}_p$ $g_i \leftarrow g^{\alpha_i}; \omega_i = a_i^{-1}b$ $u_i \leftarrow g_i^{\alpha_i}; v_i \leftarrow g_i^{\beta_i}$ $pk = (h, (g_i, u_i, v_i)_{i \in [n]})$ $sk \leftarrow ((\alpha_i, \beta_i, \omega_i)_{i \in [n]}, pk)$ Return ( $pk, sk$ )	For $i = 1$ to $n$ $r_i \leftarrow \mathbb{Z}_p; c_i \leftarrow g_i^{r_i}$ $t = \text{TCR}(c_1, \dots, c_n)$ $\pi \leftarrow \prod_{i=1}^n (u_i^t v_i)^{r_i}$ $K \leftarrow h^{r_1 + \dots + r_n}$ $\psi \leftarrow (c_1, \dots, c_n, \pi)$ Return( $K, \psi$ )	For $i = 1$ to $n$ Check if $c_i \in \mathbb{G}$ $t = \text{TCR}(c_1, \dots, c_n)$ If $\prod_{i=1}^n c_i^{\alpha_i t + \beta_i} \neq \pi$ Return ( $\perp$ ) $K \leftarrow \prod_{i=1}^n c_i^{\omega_i}$ Return( $K$ )

**Fig. 11.** KEM from  $n$ -Linear Assumption [16].

–  $\mathcal{B}$  chooses  $(x_i, y_i)_{i \in [n]}, z, z' \leftarrow \mathbb{Z}_p^*$ , and computes  $u_i = g_i^{x_i} h^z$  and  $v_i = g_i^{y_i} h^{z'}$  for  $i \in [n]$ .  $\mathcal{B}$  sets  $pk = ((g_i, u_i, v_i)_{i \in [n]}, h)$ . All the elements in  $pk$  is randomly distributed, as in the real VCI experiment. Here  $\mathcal{B}$  implicitly sets  $sk = ((\alpha_i, \beta_i, \omega_i)_{i \in [n]}, pk)$  with  $\alpha_i = x_i + \omega_i z, \beta_i = y_i + \omega_i z'$  and  $\omega_i = \log_{g_i} h$ .

–  $\mathcal{B}$  computes the challenge ciphertext  $\psi^* = (c_1^*, \dots, c_n^*, \pi^*)$  for  $\mathcal{A}$ , where  $c_i^* := g_i^{r_i^*}$  for  $i \in [n]$ ,  $t^* = \text{TCR}(c_1^*, \dots, c_n^*)$  and  $\pi^* = (K^*)^{zt^* + z'} \prod_{i=1}^n (c_i^*)^{x_i t^* + y_i}$ .

- If  $K^* = h^{\sum_{i=1}^n r_i^*}$ , we have  $\pi^* = \prod_{i=1}^n (u_i^{t^*} v_i)^{r_i^*}$ . Hence  $\psi^*$  is just a valid ciphertext output by the KEM's encapsulation algorithm with randomness  $(r_i)_{i \in [n]}$ .
- If  $K^*$  is random, then  $\pi^*$  is also random, so that  $\psi^*$  is uniformly distributed in  $\mathcal{C} = \mathbb{G}^n$ .

–  $\mathcal{B}$  uses  $((x_i, y_i)_{i \in [n]}, z, z')$  to answer  $\mathcal{A}$ 's constrained decryption queries  $(P, \psi)$ . Let  $\psi = (c_1, \dots, c_n, \pi)$ . We have that  $t = \text{TCR}(c_1, \dots, c_n) \neq t^*$  due to the target-collision resistance of TCR.  $\mathcal{B}$  computes

$$K = \left( \frac{\pi}{\prod_{i=1}^n c_i^{x_i t + y_i}} \right)^{1/(zt + z')}. \text{ If } P(K) = 1 \text{ then } \mathcal{B} \text{ returns } K; \text{ otherwise } \mathcal{B} \text{ returns } \perp.$$

- If  $\psi$  is consistent, i.e.,  $\psi$  satisfies  $\prod_{i=1}^n c_i^{\alpha_i t + \beta_i} = \pi$ , then  $\pi = h^{(zt + z') \sum_{i=1}^n r_i'}$ .  $\prod_{i=1}^n c_i^{x_i t + y_i}$ , where  $t = \text{TCR}(c_1, \dots, c_n)$  and  $r_i' = \log_{g_i} c_i$ . Then  $K = h^{\sum_{i=1}^n r_i'}$  is exactly the encapsulated key. Thus the correct  $K$  is returned to  $\mathcal{A}$  when  $P(K) = 1$ .
- If  $\psi$  is NOT consistent, then  $\pi \neq \prod_{i=1}^n c_i^{\alpha_i t + \beta_i}$ . Let  $\beta = \log_g \pi, \omega = \log_g h, a_i = \log_g g_i$ , and  $r_i' = \log_{g_i} c_i$ . Then  $\gamma := \beta - \sum_{i=1}^n a_i r_i' (\alpha_i t + \beta_i) \neq 0$ . Consequently,  $\log_g K = \gamma / (zt + z') + \omega \sum_{i=1}^n r_i'$ . The following  $2n + 2$  equations in  $2n + 2$  unknowns  $((x_i, y_i)_{i \in [n]}, z, z')$  are linearly independent, as long as  $t \neq t^*$ , which is guaranteed by the target-collision resistance of TCR:

$$\begin{aligned} \log_g u_i &= a_i x_i + \omega z \quad i = 1, 2, \dots, n \\ \log_g v_i &= a_i y_i + \omega z' \quad i = 1, 2, \dots, n \\ \log_g \pi^* &= \sum_{i=1}^n a_i r_i (t^* x_i + y_i) + (\log_g K^*) \cdot (t^* z + z') \\ \gamma \left( \log_g K - \omega \sum_{i=1}^n r_i' \right)^{-1} &= zt + z'. \end{aligned}$$

This establishes that  $zt + z'$  is uniformly distributed over  $\mathbb{Z}_p$ . Therefore,  $\log_g K$  is uniformly distributed over  $\mathbb{Z}_p$  and the predicate  $P$  satisfies  $P(K) = 0$  except with negligible probability. As a result,  $\psi$  will be correctly rejected (due to the failed predicate) except with negligible probability.

Hence,  $\mathcal{B}$  provides an almost perfect decryption oracle to  $\mathcal{A}$  as long as  $t \neq t^*$ , for all queried encapsulations  $\psi \neq \psi^*$ .

– Eventually,  $\mathcal{B}$  returns what  $\mathcal{A}$  returns.

Finally,  $\mathcal{A}$ 's non-negligible advantage in the VCI game is converted into  $\mathcal{B}$ 's non-negligible advantage in breaking the  $n$ -Linear Assumption.  $\square$

**Theorem 4.** *Suppose that the  $n$ -Linear Assumption holds for  $\mathcal{G}(1^\kappa)$ , and TCR is target collision-resistant. If groups  $\mathbb{G}$  output by  $\mathcal{G}(1^\kappa)$  are samplable and explainable, then the KEM in Figure 11 is a tailored KEM.*

*Proof.* We note that the ciphertext space  $\mathcal{C}$  equals  $\mathbb{G}^{n+1}$  and the encapsulated key space  $\mathcal{K}$  equals  $\mathbb{G}$ . If group  $\mathbb{G}$  is samplable and explainable, so are  $\mathcal{C}$  and  $\mathcal{K}$ .

Next, we have  $|\mathcal{C}| = p^{n+1}$ . For a valid ciphertext  $\psi = (c_1, \dots, c_n, \pi)$ , we note that  $\pi$  is uniquely determined by  $c_1, \dots, c_n$  and  $pk$ . Therefore, the valid ciphertext set  $|\Psi|$  has size  $p^n$ . Consequently, a random ciphertext

<b>KEM.Kg(<math>1^\kappa</math>):</b> $k \leftarrow \text{PGen}(1^\kappa)$ $pk \leftarrow i\mathcal{O}(\text{Encap}(k, \cdot))$ $sk \leftarrow k$ Return ( $pk, sk$ )	<b>KEM.Encap(<math>pk</math>):</b> $r \leftarrow \{0, 1\}^\kappa$ $(K, \psi) \leftarrow i\mathcal{O}(\text{Encap}(k, r))$ Return ( $K, \psi$ )	<b>KEM.Decap(<math>sk, \psi</math>):</b> $k \leftarrow sk$ $K \leftarrow \text{PEval}(k, \psi)$ Return ( $K$ )	<b>Encap(<math>k, r</math>):</b> $\psi \leftarrow \text{PRG}(r)$ $K \leftarrow \text{PEval}(k, \psi)$ Return ( $K, \psi$ )
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**Fig. 13.** Program of Encap.

**Fig. 12.** Sahai-Waters KEM from  $i\mathcal{O}$  and Puncturable PRF [23].

from  $\mathbb{G}^{n+1}$  passes the verification test  $\pi = \prod_{i=1}^n c_i^{\alpha_i t + \beta_i}$  in the decapsulation algorithm with negligible probability  $1/p$ . Therefore, the decapsulation algorithm has tailored functionality.

Together with Lemma 3, it follows that the KEM in Figure 11 is a tailored, and therefore suitable for the application of Theorem 1.

### 5.3 Tailored KEM Based on Indistinguishability Obfuscation and Puncturable PRF

Background definitions for this construction can be found in [23] and Appendix C.

Sahai and Waters [23] gave a KEM construction from an indistinguishability obfuscator ( $i\mathcal{O}$ ) and a puncturable PRF, as shown in Figure 12. Their construction makes use of a Pseudo-Random Generator (PRG)  $\text{PRG} : \{0, 1\}^\kappa \rightarrow \{0, 1\}^{2\kappa}$  and a puncturable PRF family  $\text{PRF} = (\text{PGen}, \text{PEval}, \text{Punc})$  whose functions map  $\{0, 1\}^{2\kappa}$  to  $\{0, 1\}^\kappa$ . We assume that (descriptions of) PRG and PRF are implicitly part of the inputs to KEM.Kg, KEM.Decap in Figure 12 and Encap in Figure 13.

The ciphertext space of the KEM is  $\mathcal{C} = \{0, 1\}^{2\kappa}$ , the valid ciphertext set is  $\Psi = \{\psi \mid \psi = \text{PRG}(r); r \in \{0, 1\}^\kappa\}$ , and the key space is  $\mathcal{K} = \{0, 1\}^\kappa$ . Obviously, both of  $\mathcal{C}$  and  $\mathcal{K}$  are efficiently samplable and explainable with  $\text{SampleC}^{-1}(\mathcal{C}, \psi) := \psi$  and  $\text{SampleK}^{-1}(\mathcal{K}, K) := K$ .

**Lemma 5** *If  $i\mathcal{O}(\cdot)$  is an indistinguishability obfuscator for  $P/\text{poly}$ , PRG is a secure PRG, and PRF is a puncturable PRF, then the Sahai-Waters KEM in Figure 12 is IND-tCCCA secure.*

*Proof.* In [23], the Sahai-Waters KEM was proved to be IND-CCA secure, so it is obviously IND-CCCA secure.

Next we prove the VCI property, based on the security of PRG. If there is a ppt adversary  $\mathcal{A}$  that can distinguish a random ciphertext from a random valid ciphertext with non-negligible probability, then we can construct a ppt algorithm  $\mathcal{B}$  that breaks the security of PRG. Suppose  $\mathcal{B}$  is given an element  $\psi^*$  and tries to decide whether  $\psi^*$  is the output of PRG or a randomly chosen element from  $\mathcal{C}$ .  $\mathcal{B}$  will simulate a VCI experiment for  $\mathcal{A}$ . It first chooses a puncturable PRF PRF and calls KEM.Kg( $1^\kappa$ ) to generate  $(pk, sk)$ . The public key  $pk$  is given to  $\mathcal{A}$ . Then  $\mathcal{B}$  gives  $\psi^*$  as the challenge encapsulation to  $\mathcal{A}$ . Using the secret key  $sk$  and algorithm PEval,  $\mathcal{B}$  is able to provide a (constrained) decryption oracle for  $\mathcal{A}$ . Finally,  $\mathcal{B}$  outputs whatever  $\mathcal{A}$  outputs. Then it is easy to see that  $\mathcal{A}$ 's non-negligible advantage in the VCI security game results in a non-negligible advantage for  $\mathcal{B}$  in breaking the security of PRG.

The IND-CCCA security and VCI property in combination with Lemma 1 establish that the Sahai-Waters KEM in Figure 12 has IND-tCCCA security.  $\square$

Extracting puncturable PRFs are a strengthening of puncturable PRFs introduced in [23]; essentially, an extracting puncturable PRF acts as a strong extractor on its inputs.

**Definition 9 (Extracting puncturable PRF)** *Let  $\epsilon(\cdot)$  and  $h_{\min}(\cdot)$  be functions. A puncturable PRF family  $\text{PRF} = (\text{PGen}, \text{PEval}, \text{Punc})$  mapping  $\{0, 1\}^{\ell_1(\kappa)}$  to  $\{0, 1\}^{\ell_2(\kappa)}$  is said to be extracting with error  $\epsilon(\kappa)$  for min-entropy function  $h_{\min}(\kappa)$  if for all  $\kappa \in \mathbb{N}$  and for all random variables  $X$  on  $\{0, 1\}^{\ell_1(\kappa)}$  with min-entropy greater than  $h_{\min}(\kappa)$ , the statistical distance between  $(k, \text{PEval}(k, X))$  and  $(k, U_{\ell_2(\kappa)})$  is at most  $\epsilon(\kappa)$ , where  $k \leftarrow \text{PGen}(1^\kappa)$  and  $U_{\ell_2(\kappa)}$  denotes the uniform distribution over  $\{0, 1\}^{\ell_2(\kappa)}$ . The family PRF is said to be extracting puncturable if the error  $\epsilon(\kappa)$  is negligible (for some choice of function  $h_{\min}$ ).*

The existence of extracting puncturable PRFs is implied by the existence of one-way functions, as was proved in [23]:

**Lemma 4.** [23] *Assume that one-way functions exist. Then for all efficiently computable functions  $\ell_1(\kappa)$ ,  $\ell_2(\kappa)$ ,  $e(\kappa)$  and  $h_{\min}(\kappa)$  such that  $\ell_1(\kappa) \geq h_{\min}(\kappa) \geq \ell_2(\kappa) + 2e(\kappa) + 2$ , there exists an extracting puncturable PRF family  $\text{PRF} = (\text{PGen}, \text{PEval}, \text{Punc})$  mapping  $\{0, 1\}^{\ell_1(\kappa)}$  to  $\{0, 1\}^{\ell_2(\kappa)}$  with error function  $\epsilon(\kappa) = 2^{-e(\kappa)}$  and min-entropy function  $h_{\min}(\kappa)$ .*

**Lemma 6** *If PRF is an extracting puncturable PRF obtained from Lemma 4, then the decapsulation algorithm  $\text{KEM.Decap}$  of the Sahai-Waters KEM in Figure 12 has tailored functionality.*

*Proof.* We show that the output of  $\text{PRF}(sk, \psi)$  is statistically close to the uniform distribution on  $\{0, 1\}^\kappa$  so long as  $\psi$  is chosen from  $\mathcal{C}$  uniformly at random, and the puncturable PRF satisfies the bounds in Lemma 4.

Recall that PRF maps  $2\kappa$  bits to  $\kappa$  bits. When  $\psi$  is randomly chosen from  $\{0, 1\}^{2\kappa}$ , the min-entropy of  $\psi$  is  $2\kappa$ . According to Lemma 4, the statistical distance between  $(k, \text{PEval}(k, \psi))$  and  $(k, U_\kappa)$  is upper-bounded by  $2^{-(\kappa/2-1)}$ , where  $k \leftarrow \text{PGen}(1^\kappa)$  and  $U_\kappa$  is the uniform distribution over  $\{0, 1\}^\kappa$ . Hence,  $\text{KEM.Decap}$  has  $2^{-(\kappa/2-1)}$ -tailored functionality.  $\square$

**Theorem 7.** *If  $i\mathcal{O}(\cdot)$  is an indistinguishability obfuscator for  $P/\text{poly}$ , PRG is a secure PRG, and PRF is an extracting puncturable PRF, then the Sahai-Waters KEM in Figure 12 is a tailored KEM.*

*Proof.* The fact that the KEM in Figure 12 is a tailored KEM follows immediately from Lemma 5, Lemma 6 and the fact that  $\mathcal{C} = \{0, 1\}^{2\kappa}$  and  $\mathcal{K} = \{0, 1\}^\kappa$  are efficiently samplable and explainable.

The existence of one-way functions implies the existence of PRGs and extracting puncturable PRFs. Hence the existence of one-way functions and  $i\mathcal{O}$  implies the existence of a tailored KEM by the above theorem. Such a tailored KEM can further be used to build a PKE scheme encrypting  $\ell$  bits at a time with the help of an information-theoretically secure  $(\ell + s)$ -XAC (for suitable parameter  $s$ ), by following the construction in Figure 4; the SIM-SO-CCA security of the PKE scheme follows from Theorem 1. Thus we obtain the following corollary:

**Corollary 8** *Suppose one-way functions and indistinguishability obfuscation for  $P/\text{poly}$  exist. Then there exists a PKE scheme with SIM-SO-CCA security.*

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## A CCCA security for KEMs

A chosen-ciphertext attack against a KEM [9] allows an adversary to query a decapsulation oracle  $\text{Decap}(\cdot)$ , (which runs  $\text{KEM.Decap}$ ) on any input  $\psi \neq \psi^*$  where  $(K', \psi^*)$  is given to the adversary,  $\psi^*$  is an encapsulation of  $K^*$  and  $K'$  is either  $K^*$  or random. In a relaxed version of chosen-ciphertext attacks [16], the adversary is allowed to query a *constrained* decapsulation oracle  $\widetilde{\text{Decap}}(\cdot)$  only if it has some *a priori* knowledge about the decapsulated key, which can be characterized by a ppt Boolean predicate  $P : \mathcal{K} \cup \{\perp\} \rightarrow \{0, 1\}$  (for

simplicity, we put  $\perp$  into the domain and define  $P(\perp) = 0$ ). The constrained decapsulation oracle  $\widetilde{\text{Decap}}(\cdot)$  replies with  $K$  for a query  $\psi$  if and only if  $P(K) = 1$  where  $K \leftarrow \text{KEM.Decap}(sk, \psi)$ .

Constrained CCA security for a KEM ( $\text{KEM.Kg}, \text{KEM.Enc}, \text{KEM.Dec}$ ) is formalised using the security experiment shown in Figure 14.

$\text{Exp}_{\text{KEM}, \mathcal{A}}^{\text{ccca}-b}(\kappa) :$ $(pk, sk) \leftarrow \text{KEM.Kg}(1^\kappa)$ $K_0^* \leftarrow \mathcal{K}, (K_1^*, \psi^*) \leftarrow \text{KEM.Encap}(pk)$ $b \leftarrow \widetilde{\text{Decap}}_{\neq \psi^*}(\cdot)(pk, K_b^*, \psi^*)$ $\text{Return}(b)$	$\widetilde{\text{Decap}}_{\neq \psi^*}(P, \psi)$ $\text{If } \psi = \psi^* \text{ return } (\perp)$ $K \leftarrow \text{KEM.Decap}(sk, \psi)$ $\text{If } P(K) = 0 \text{ return } (\perp);$ $\text{Else return } (K)$
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**Fig. 14.** CCCA security experiment for KEMs.

The advantage of an adversary  $\mathcal{A}$  in the CCCA experiment is defined as

$$\text{Adv}_{\text{KEM}, \mathcal{A}}^{\text{ccca}}(\kappa) := \left| \Pr \left[ \text{Exp}_{\text{KEM}, \mathcal{A}}^{\text{ccca}-0}(\kappa) = 1 \right] - \Pr \left[ \text{Exp}_{\text{KEM}, \mathcal{A}}^{\text{ccca}-1}(\kappa) = 1 \right] \right|.$$

Now let  $q_d$  denote the number of decapsulation queries made by  $\mathcal{A}$ . The *uncertainty* of  $\mathcal{A}$  measures the average fraction of keys for which the evaluation of predicate  $P_i(\cdot)$  is equal to 1 in the CCCA experiment, where  $P_i$  is the predicate used in the  $i$ -th query by  $\mathcal{A}$ , i.e.,

$$\text{uncert}_{\mathcal{A}}(\kappa) := \frac{1}{q_d} \sum_{i=1}^{q_d} \Pr [P_i(K) = 1].$$

**Definition 10 (IND-CCCA Security [16])** *The KEM  $\text{KEM}$  is secure against constrained chosen ciphertext attacks (IND-CCCA secure) if for all ppt adversaries  $\mathcal{A}$  such that  $\text{uncert}_{\mathcal{A}}(\kappa)$  is negligible (in  $\kappa$ ), the advantage  $\text{Adv}_{\text{KEM}, \mathcal{A}}^{\text{ccca}}(\kappa)$  is also negligible in  $\kappa$ .*

## B (Target) Collision Resistant Hash Functions

**Definition 11 (CR hash function)** *Let  $H$  be a family of hash functions associated with two ppt algorithms ( $H\text{Gen}, H\text{Eval}$ ). On input a security parameter  $\kappa$ ,  $H\text{Gen}$  outputs a (description of a) function  $H : \mathcal{D}_\kappa \rightarrow \mathcal{R}_\kappa$ . Algorithm  $H\text{Eval}$ , on input an element  $x \in \mathcal{D}_\kappa$ , outputs  $H(x) \in \mathcal{R}_\kappa$ . Then  $H$  is collision-resistant if for all ppt algorithms  $\mathcal{B}$  the following advantage is negligible in  $\kappa$ :*

$$\text{Adv}_{H, \mathcal{B}}^{\text{cr}}(\kappa) := \Pr [x' \neq x \wedge H(x') = H(x) \mid H \leftarrow H\text{Gen}(1^\kappa), (x, x') \leftarrow \mathcal{B}(H)].$$

**Definition 12 (TCR hash function)** *Let  $H$  be a family of hash functions associated with two ppt algorithms ( $H\text{Gen}, H\text{Eval}$ ). The security notion of target collision-resistance (TCR) requires that for all ppt algorithms  $\mathcal{B}$  the following advantage is negligible in  $\kappa$ :*

$$\text{Adv}_{H, \mathcal{B}}^{\text{tcr}}(\kappa) := \Pr [x' \neq x \wedge H(x') = H(x) \mid H \leftarrow H\text{Gen}(1^\kappa), x \leftarrow \mathcal{D}_\kappa, x' \leftarrow \mathcal{B}(H, x)].$$

## C Indistinguishability Obfuscation ( $i\mathcal{O}$ ) and Puncturable PRFs

*Indistinguishability Obfuscation ( $i\mathcal{O}$ )* was proposed in [1, 2]. Informally, it requires that the obfuscations of two distinct but equal-sized programs, which implement identical functionality, are computationally indistinguishable. The first candidate construction for  $i\mathcal{O}$  for polynomial-sized circuits was given in [13]. We follow the formal definitions in [23].

**Definition 13 (Indistinguishability Obfuscation ( $i\mathcal{O}$ ) for  $P/\text{poly}$  [23])** *Let  $\kappa$  be a security parameter. A uniform ppt machine  $i\mathcal{O}$  is called an indistinguishability obfuscator for a circuit class  $\text{Circ}_\kappa$  if the following two conditions are satisfied.*



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$\text{Exp}_{\text{PRF},\mathcal{A}}^{\text{punc}-b}(\kappa)$  :

$(\sigma, S) \leftarrow \mathcal{A}_1(1^\kappa); k \leftarrow \text{PGen}(1^\kappa); k_S \leftarrow \text{Punc}(k, S)$   
 $y_1 = \text{PEval}(k, S); y_0 \leftarrow \{0, 1\}^{|\mathcal{S}|\ell_1(\kappa)}; b' \leftarrow \mathcal{A}_2(\sigma, k_S, S, y_b)$   
 Return( $b'$ )

---

**Fig. 15.** Experiment for defining punctured PRF pseudo-randomness.

- **Identical functionality:** For all  $\kappa \in \mathbb{N}$ , for all  $C \in \text{Circ}_\kappa$ , on input  $(1^\kappa, C)$ ,  $i\mathcal{O}$  outputs equal-sized circuits all having functionality identical to  $C$ , i.e, for all inputs  $x$ ,

$$\Pr[C'(x) = C(x) \mid C' \leftarrow i\mathcal{O}(1^\kappa, C)] = 1.$$

- **Indistinguishability:** For all (not necessarily uniform) ppt stateful adversaries  $\mathcal{D} = (\mathcal{D}_1, \mathcal{D}_2)$ , there exists a negligible function  $\varepsilon(\kappa)$  satisfying the following condition. Suppose that for all  $x$

$$\Pr[C_0(x) = C_1(x) \mid (C_0, C_1, \sigma) \leftarrow \mathcal{D}_1(1^\kappa)] > 1 - \varepsilon(\kappa).$$

Then

$$\begin{aligned} & \left| \Pr[\mathcal{D}_2(\sigma, i\mathcal{O}(\kappa, C_0)) = 1 \mid (C_0, C_1, \sigma) \leftarrow \mathcal{D}_1(1^\kappa)] \right. \\ & \left. - \Pr[\mathcal{D}_2(\sigma, i\mathcal{O}(\kappa, C_1)) = 1 \mid (C_0, C_1, \sigma) \leftarrow \mathcal{D}_1(1^\kappa)] \right| \leq \varepsilon(\kappa). \end{aligned}$$

A uniform ppt machine  $i\mathcal{O}$  is called an *indistinguishability obfuscator* for  $P/\text{poly}$  if  $i\mathcal{O}$  is an *indistinguishability obfuscator* for the class of circuits of size at most  $\kappa$ .

**Definition 14 (Puncturable PRF)** A puncturable family of PRFs  $\text{PRF}$  is indexed by  $k$ , parameterized by computable functions  $(\ell_1(\kappa), \ell_2(\kappa))$ , and consists of three ppt algorithms  $(\text{PGen}, \text{PEval}, \text{Punc})$ . Algorithm  $\text{PGen}(1^\kappa)$  samples a random key  $k$ ,  $\text{PEval}(k, x)$  computes the function  $\text{PRF}_k(\cdot)$ , which maps  $\{0, 1\}^{\ell_1(\kappa)}$  to  $\{0, 1\}^{\ell_2(\kappa)}$ , and  $\text{PPunc}(k, S)$  takes as input the index  $k$  and a polynomial-sized set  $S$  of inputs and outputs a punctured key  $k_S$ . The following conditions are satisfied:

- **Functionality preserved under puncturing.** For all ppt adversaries  $\mathcal{A}$ , for all polynomial-sized subsets  $S \subseteq \{0, 1\}^{\ell_1(\kappa)}$  given by  $S \leftarrow \mathcal{A}(1^\kappa)$ , for all  $x \in \{0, 1\}^{\ell_1(\kappa)} \setminus S$ , we have that

$$\Pr[\text{PEval}(k, x) = \text{PEval}(k_S, x) \mid k \leftarrow \text{PGen}(1^\kappa), k_S \leftarrow \text{Punc}(k, S)] = 1.$$

- **Pseudo-random at punctured points.** For all stateful ppt adversaries  $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2)$ , define the advantage of  $\mathcal{A}$  against the punctured pseudo-randomness of  $\text{PRF}$  to be

$$\text{Adv}_{\text{PRF},\mathcal{A}}^{\text{punc}}(\kappa) := \left| \Pr \left[ \text{Exp}_{\text{PRF},\mathcal{A}}^{\text{punc}-0}(\kappa) = 1 \right] - \Pr \left[ \text{Exp}_{\text{PRF},\mathcal{A}}^{\text{punc}-1}(\kappa) = 1 \right] \right|$$

where  $\text{Exp}_{\text{PRF},\mathcal{A}}^{\text{punc}-b}(\kappa)$  is defined in Figure 15. Then  $\text{Adv}_{\text{PRF},\mathcal{A}}^{\text{punc}}(\kappa)$  is negligible.