Non-Committing Identity Based Encryption: Constructions and Applications

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Abstract. A receiver non-committing encryption (RNCE) scheme [Canetti *et al.*, STOC 1996; Canetti *et al.*, TCC 2005] allows one to sample a public key pk and (dummy) ciphertext ct without knowing the message m. Later, when the message is known, one can sample a secret key sk that looks like the secret key corresponding to pk, and decryption of ct produces m. In this work, we study receiver non-committing identity-based encryption (RNC-IBE). We give constructions based on standard assumptions on bilinear groups (prior works [Hiroka *et al.*, ASIACRYPT 2021] require indistinguishability obfuscation).

Our RNC-IBE constructions have important implications for incompressible identity based encryption. This notion was recently introduced by Goyal *et al.*, ITCS 2025. However, there were no constructions for the strongest security definitions in Goyal *et al.*, ITCS 2025. Our RNC-IBE scheme also leads to the first incompressible IBE scheme with optimal ciphertext size, which was another open question in Goyal *et al.*, ITCS 2025.

We also give constructions for relaxed RNC-IBE (where the identity space is polynomial in the security parameter, but the public key is compact) that are based on $\mathsf{DDH}, \mathsf{LWE}$. This leads to a relaxed incompressible IBE scheme with strong security from the same assumptions.

Keywords: non-committing · identity-based encryption · incompressible encryption.

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1 Introduction

Non-committing encryption. Non-committing public-key encryption (NCE), introduced by Canetti, Feige, Goldreich and Naor [17], is a crucial cryptographic tool in the design of adaptively secure multiparty computation protocols [17, 20]. An NCE scheme consists of the following algorithms - $\mathsf{Setup}, \mathsf{Enc}, \mathsf{Dec}$ and $\mathsf{Sim} = (\mathsf{Sim}_1, \mathsf{Sim}_2)$. The syntax for $\mathsf{Setup}, \mathsf{Enc}$ and Dec mirrors that of (standard) public key encryption (PKE) schemes. Additionally, the simulator Sim enables the sampling of ciphertext without knowledge of the underlying message. Once the message is revealed, the simulator can generate the necessary randomness to reconcile the public key and the ciphertext. More formally, security is captured using the real and ideal world framework. In the real-world, the adversary first receives the public key. It then sends a message m, and receives an encryption of m, together with the randomness r_{enc} used for encryption, and the randomness r_{setup} used for sampling the public key. In the ideal world, the simulator Sim_1 produces the public key pk and ciphertext ct^* (together with internal state st which is passed to Sim_2). After the adversary receives the public key, it sends the message m^*. The second-stage simulator Sim_2 receives the message m^* (and the internal state st) samples randomness r_{setup} that can explain pk, and randomness r_{enc} to explain ct^*. The adversary receives $\mathsf{ct}^*, r_{\mathsf{enc}}$ and r_{setup}, and must distinguish between the real-world and ideal-world. Today, we have several constructions of NCE, from a wide range of assumptions [14, 21, 26, 51, 71] as well a good understanding of the barriers [60].

Prior works have also explored weaker notions of non-committing encryption. Receiver NCE (RNCE) [5] is one such relaxation which has garnered significant attention. Here, the simulator Sim_2 only needs to output the secret key corresponding to pk (but does not need to produce the randomness r_{enc} and r_{setup}). Receiver NCE has been used for applications such as designing secure multiparty computation protocols [17, 20], adaptive secure attribute-based encryption for Turing Machines [45], selective opening secure schemes [50] and more. In this work, we will focus on RNCE, with the aim to go beyond public key encryption.

Receiver Non-committing Identity Based Encryption (and beyond). Identity based encryption (IBE) [64] is a powerful generalization of public key encryption, where users can encrypt messages for any identity using a master public key. The master public key, together with a corresponding master secret key, is sampled by the master authority using a Setup algorithm. The master secret key can be used to issue secret keys for every identity using a Keygen algorithm. Using a secret key for identity id, one can decrypt a ciphertext for identity id. Intuitively, security says that even if an adversary has polynomially many secret keys corresponding to identities of its choice, if it does not have a secret key for id^*, then the adversary cannot decrypt an encryption for id^*.

In the non-commmitting setting, identity-based encryption (and more generally, attribute-based encryption) was introduced by Hiroka, Morimae, Nishimaki and Yamakawa [53] in the context of attribute-based quantum encryption

[5] Also referred to as weak NCE [45].

with certified deletion. In a receiver non-committing identity-based encryption (RNC-IBE) scheme, in addition to $(\mathsf{Setup}, \mathsf{Keygen}, \mathsf{Enc}, \mathsf{Dec})$, we have a simulator $\mathsf{Sim} = (\mathsf{Sim}_1, \mathsf{Sim}_2)$ that can produce the master public key, secret keys and challenge ciphertext without knowing the challenge message. Later, when the message is revealed, the simulator can sample a master secret key that is consistent with the master public key, secret keys and the challenge ciphertext. More formally, in the real world, the adversary receives the master public key mpk. Then, it can send polynomially many identities, and receives the secret keys corresponding to these identities. The adversary then sends the challenge identity id^*, together with challenge messages m^*, and receives $\mathsf{ct}^* \leftarrow \mathsf{Enc}(\mathsf{mpk}, \mathsf{id}^*, m^*)$ together with the master secret key msk. In the ideal world, the adversary interacts with a simulator. The simulator first sends the master public key mpk. Then, the adversary receives secret keys for identities of its choice. Finally, in the challenge phase, when the adversary sends the challenge identity id^* and the challenge message m^*, the simulator must first produce ct^* using just id^*. Later, when it receives m^*, it must produce a master secret key msk. The adversary finally receives ct^* and msk, and must distinguish the real and ideal worlds.

While we have several constructions for IBE [1–3, 9–13, 15, 18, 22, 23, 27, 30, 31, 38–40, 43, 49, 55, 57, 65, 66, 70] (and even ABE [4, 5, 7, 41, 42, 45, 46, 62, 68]), our understanding with respect to RNC-IBE and RNC-ABE is very limited. Hiroka *et al.* [53] gave a construction for RNC-ABE using indistinguishability obfuscation. This brings us to the first central question of our work.

Q1. Can we construct RNC-IBE from the same assumptions that give us (regular) IBE?

Besides being a natural question in itself, this would resolve interesting open questions in the landscape of incompressible cryptography, which we discuss next.

Incompressible (Identity-Based) Encryption. The concept of incompressible public key encryption was introduced by Guan, Wichs and Zhandry [47] to address scenarios where the adversary eventually receives the entire secret decryption key, but has limited long-term storage, and as a result, cannot store the entire ciphertext. For S-incompressible security, we require that no adversary should win the following game (with non-negligible probability): the adversary, after receiving the public key pk, sends two challenge messages m_0, m_1, and receives the challenge ciphertext ct^*. It must then compress the ciphertext into a short state st of size at most S bits. After it computes the compressed state st, it receives the secret key sk, and must guess whether m_0 was encrypted or m_1. Guan *et al.* gave two constructions of incompressible PKE : one based on general PKE (with ciphertext size being $(S + |m|) \cdot \mathrm{poly}(\lambda)$), and another based on indistinguishability obfuscation (with ciphertext size $\max(S, |m|) + \mathrm{poly}(\lambda)$).

Later works [44, 48] observed that any RNCE scheme can be used to build an incompressible PKE scheme with ciphertext size $S + |m| + \mathrm{poly}(\lambda)$ as follows. Consider an RNCE scheme and an incompressible SKE scheme with a secret key size of $\mathrm{poly}(\lambda)$ and a ciphertext size $(|m| + S + \mathrm{poly}(\lambda))$ (Dziembowzki [32] gave a

construction of such incompressible SKE scheme, based on one-way functions). In the incompressible PKE scheme, the public and secret keys are identical to those of the RNCE scheme. To encrypt a message m, first generate a fresh secret key $\mathsf{inc.sk}$ for the incompressible SKE scheme. Using this key, encrypt the message to produce an incompressible ciphertext $\mathsf{inc.ct}$. Next, encrypt $\mathsf{inc.sk}$ using the RNCE to obtain $\mathsf{nce.ct}$, and the final ciphertext becomes $\mathsf{ct} := (\mathsf{nce.ct}, \mathsf{inc.ct})$. Note that this scheme is ciphertext-rate preserving, as the size of $\mathsf{nce.ct}$ is $|\mathsf{nce.ct}| = S + |m| + \mathrm{poly}'(\lambda)$.

To argue security, we begin by switching the RNCE to simulation mode, which allows us to freely program the secret key so that $\mathsf{nce.ct}$ can decrypt to any desired value. At this point, the scheme's security relies on the incompressibility of the SKE scheme.

In this work, we focus on incompressible identity based encryption schemes. This primitive, introduced in a recent work by Goyal, Koppula, Rajasree and Verma [44], is a natural generalization of incompressible PKE to the IBE setting. Here again, the syntax is same as that of (regular) IBE. For incompressibility security, however, note that there can be multiple flavors of security. Goyal *et al.* defined two notions of security for incompressible IBE:

- (regular) incompressible security: in this case, the adversary first receives the master public key. Then it can send polynomially many identities, and receives secret keys for these identities. During the challenge phase, the adversary sends a challenge identity id^* together with challenge messages m_0, m_1, and receives the challenge ciphertext, followed by post-challenge secret key queries (similar to the pre-challenge secret key queries). Finally, the adversary must compress the ciphertext into a short state st of size at most S bits. After this, it receives the secret key corresponding to id^*, and must guess whether m_0 was encrypted or m_1.
- strong incompressible security: this game is similar to that of regular incompressible IBE. However, instead of receiving the secret key for id^* at the end, the adversary receives the entire master secret key.

Goyal *et al.* gave constructions for (regular) incompressible IBE, however there were no constructions achieving strong incompressibility! Our first observation is that the connection between incompressible PKE and RNCE also extends to IBE, and as a result, if we construct an RNC-IBE scheme, then that also resolves the following question (left open in [44]).

Q2. Can we construct strongly secure incompressible IBE?

Ciphertext rate of incompressible encryption schemes. An important parameter in the design of incompressible encryption schemes is the size of the ciphertext, as a function of the message size and the adversary's long-term storage bound S. The optimal ciphertext size (ignoring dependence on λ) is $\max(S, |m|)$. Guan *et al.* [47] showed how to construct incompressible PKE schemes with optimal ciphertext size, using indistinguishability obfuscation. Later, Branco, Döttling and Dujmovic [16] showed how to construct incompressible PKE schemes with optimal

ciphertext size using standard assumptions (that is, without using obfuscation). A natural question is whether we can achieve incompressible IBE schemes with optimal ciphertext size. This was also left as an open question by Goyal *et al.* [44].

Q3. Can we construct incompressible IBE schemes with optimal ciphertext size?

1.1 Our results

In this work, we introduce new constructions for RNC-IBE and receiver non-committing identity-based key encapsulation mechanism, based on various standard assumptions. These construction, in turn, give us the first strong incompressible IBE scheme *with* optimal ciphertext size.

Main Results: The first construction is based on the bilinear DDH assumption, utilizing the dual system technique of Waters [67]. Notably, this construction achieves a robust security notion where the adversary obtains the entire randomness used by the setup algorithm (see Remark 1 for further details). Also, the size of the ciphertext is independent of the size of the session key.

Theorem 1. *Assuming the hardness of* SXDH *problem, there exists an adaptively secure RNC-IB-KEM. Additionally, the adversary is allowed to learn the entire randomness of setup, and the size of the ciphertext is independent of the size of the session key.*

Note that Theorem 1 addresses Q1 from above. Additionally, it also resolves questions Q2 and Q3 simultaneously! In fact, the resulting incompressible IBE scheme achieves the strongest possible security, where even the randomness used during setup can be revealed to the adversary. This is obtained by combining the RNC-IB-KEM with the incompressible secret key encryption scheme of [16] (based on the LWE or DCR assumptions). In this incompressible encryption scheme, the size of the ciphertext is $\max(S, |m|) + \mathrm{poly}(\lambda)$. The size of the secret key grows with $|S|$, but since the size of the RNC-IB-KEM's ciphertext is independent of the session-key size, this does not affect the final ciphertext size.

Theorem 2. *Assuming the existence of RNC-IB-KEM such that the size of the ciphertext is independent of the size of the session key, there exists an adaptively secure strongly incompressible IBE scheme with optimal ciphertext size (*$\max(S, |m|) + \mathrm{poly}(\lambda)$*). In particular, we get adaptively secure strongly incompressible IBE with optimal ciphertext size, assuming the hardness of* SXDH *and* LWE *(or* DCR*).*

RNC-IBE and Incompressible IBE Constructions for polynomially bounded identity space: The second RNC-IBE construction supports polynomially many identities with compact master public key. It can be instantiated from a broader class of assumptions such as {DDH, LWE}. This construction introduces an additional feature where non-committing ciphertext can be generated together with the master public key, i.e., it does not require knowledge of the

target identity id^*. This is the first construction to offer such capability which may be of independent interest. Similar to the first construction, this scheme remains secure when the randomness used by the setup algorithm is provided to the adversary.

Theorem 3. *Assuming the hardness of $\mathcal{X}$ where $\mathcal{X} \in \{\mathsf{DDH}, \mathsf{LWE}\}$, there exists an adaptively secure NC-IBE that supports polynomially many identities.*

By combining these results with an incompressible secret key encryption scheme in a hybrid encryption framework, we obtain the first strongly incompressible IBE schemes from $\mathsf{DDH}/\mathsf{LWE}$. The incompressible IBE scheme supports polynomially many identities with compact master public key. Finally, it remains secure even if the adversary receives the randomness used during the setup.

Theorem 4. *Assuming the hardness of $\mathcal{X}$ where $\mathcal{X} \in \{\mathsf{DDH}, \mathsf{LWE}\}$, there exists an adaptively secure (super) strongly incompressible IBE schemes that supports polynomially many identities.*

Constructions from Indistinguishability Obfuscation: We additionally present an RNC-IBE scheme using indistinguishability obfuscation (iO) and one-way functions. While [53] also gave a construction of RNC-ABE using iO and one way functions, our approach differs substantially.

Theorem 5. *Assuming the existence of iO and one-way functions, there exists selective secure RNC-IBE.*

The construction and the proof for the above theorem is provided in the full version. In this construction, the ciphertext rate is poor because the size of the ciphertext depends on both the length of the identity and $|m| \cdot \mathrm{poly}(\lambda)$. However, by employing Dziembowzki's incompressible SKE scheme (with secret key size of $\mathrm{poly}(\lambda)$), we can obtain rate-$\frac{1}{2}$ strongly incompressible IBE schemes (see Theorem 10).

Theorem 6. *Assuming the existence of iO and one-way functions, there exists a rate-$\frac{1}{2}$ selectively secure strongly incompressible IBE schemes.*

1.2 Related Works

In the field of incompressible encryption, Dziembowski [32] introduced the first constructions for incompressible symmetric key encryption (SKE). He presented an information-theoretic scheme with a rate of $\frac{1}{3}$, as well as a construction achieving a rate of $\frac{1}{2}$ based on one-way functions. After a decade, Guan *et al.* [47] introduced two incompressible public key encryption (PKE) schemes – the first, although based on standard PKE, had poor compression efficiency, while the second employed indistinguishability obfuscators [36, 37, 63] to realize a rate-1 scheme. Branco *et al.* [16] followed up by designing a rate-1 incompressible PKE

scheme that offers chosen ciphertext attack (CCA) security, combining a rate-1 incompressible SKE with programmable hash proof systems. More recently, Guan *et al.* [48] advanced the notion by developing multi-user incompressible encryption. Here, the adversary is given multiple ciphertexts encrypted with different secret keys.

Goyal *et al.* [44] extended the concept to functional and attribute-based encryption (ABE), introducing a variety of incompressible security notions for functional encryption, attribute-based encryption, and identity-based encryption. Their work also provided constructions for incompressible functional encryption that achieves optimal trade-off between ciphertext-size and secret key size.

In another direction, Bhushan *et al.* [8] explored incompressible encryption in the context of leakage resilience. They presented a range of leakage-resilient incompressible encryption schemes tailored to various leakage functions. Their work also examined the challenges of constructing rate-1 schemes with short ciphertexts or schemes that can withstand significant leakage.

In addition to encryption, Guan *et al.* [47] also proposed incompressible signature schemes, which guarantee that an adversary cannot forge or reconstruct a signature from a compressed version. A related area is incompressible encodings [28, 34, 59], where it is computationally hard to reconstruct a codeword from a compressed version, even with access to the original message. Prior research [28, 34, 59] has shown positive results for incompressible encodings within the random oracle and common reference string (CRS) models.

The area of non-committing encryption (NCE) has also seen significant work focusing on building NCE schemes where the adversary gains access to the randomness used during the setup and encryption phases. These works developed schemes under various assumptions with the goal of achieving high ciphertext-rate. [6, 14, 17, 21, 26, 29, 51, 52, 71, 72].

Another direction explores optimizing parameters for weaker forms of NCE, where the adversary is restricted to gain access to the randomness used in either the setup (receiver) or encryption (sender). Jarecki and Lysyanskaya [56] introduced a scheme that is non-committing only for the receiver, whereas Canetti, Halevi and Katz [19] constructed a constant-rate NCE with erasures, where the adversary only receives the secret key and the ciphertext. Hiroka *et al.* [53] introduced non-committing attribute-based encryption (NC-ABE) using indistinguishable obfuscators, focusing on achieving ABE with certified deletion in quantum settings, where the adversary receives the master secret key along with the ciphertext.

2 Technical Overview

RNC-IBE from Bilinear Groups

Dual system encryption is a versatile framework employed in the construction of numerous IBE [25, 35, 54, 57, 67] and ABE [4, 5, 23, 24, 33, 41, 58, 61, 69] schemes. In our construction, we utilize a pairing group $e : \mathbb{G}_1 \times \mathbb{G}_2 \to \mathbb{G}_T$, where

$\mathbb{G}_1, \mathbb{G}_2, \mathbb{G}_T$ are all cyclic groups of prime order p, generated respectively by g_1, g_2, and $e(g_1, g_2)$, where e is a non-degenerate bilinear map, that is, for all $a, b \in \mathbb{Z}_p, e(g_1^a, g_2^b) = e(g_1, g_2)^{ab}$. We use bracket notations, where for all exponents $a \in \mathbb{Z}_p$ and all groups $s \in \{1, 2, T\}$, we denote by $[a]_s$ the group element g_s^a. This notation extends to vectors and matrices as well.

We describe our construction, which produces a session key in $\mathbb{G}_T$. The setup algorithm generates $\mathbf{a}, \mathbf{b} \leftarrow \mathbb{Z}_p^2$ and $\mathbf{W}_1, \mathbf{W}_2 \leftarrow \mathbb{Z}_p^{2\times 2}$ and sets the public parameters

$$\mathsf{pp} := ([\mathbf{a}]_1, [\mathbf{b}]_2, [\mathbf{W}_1\mathbf{a}]_1, [\mathbf{W}_2\mathbf{a}]_1, [\mathbf{W}_1^\top\mathbf{b}]_2, [\mathbf{W}_2^\top\mathbf{b}]_2)$$

.

It then generates a secret vector $\mathbf{k} \leftarrow \mathbb{Z}_p^2$, setting the master public key as $\mathsf{mpk} := [\mathbf{a}^\top\mathbf{k}]_T$ and the master secret key as $\mathsf{msk} := \mathbf{k}$.

To generate a secret key for a specific identity $\mathsf{id} \in \mathbb{Z}_p$, the algorithm sample a random element $s \leftarrow \mathbb{Z}_p$ and output $\mathsf{sk}_{\mathsf{id}} := ([s\mathbf{b}]_2, [\mathbf{k} + s(\mathbf{W}_1 + \mathsf{id} \cdot \mathbf{W}_2)^\top\mathbf{b}]_2)$. To generate a session key and its corresponding ciphertext for a target identity id^*, the algorithm generates a random element $r \leftarrow \mathbb{Z}_p$ and produces the ciphertext $\mathsf{ct} := ([r\mathbf{a}]_1, [r(\mathbf{W}_1 + \mathsf{id} \cdot \mathbf{W}_2)\mathbf{a}]_1)$ and the session key $\mathsf{seskey} := [r\mathbf{a}^\top\mathbf{k}]_T$. For decapsulation, given the secret key $\mathsf{sk}_{\mathsf{id}} = ([\mathbf{d}]_2, [\mathbf{d}']_2)$ and a ciphertext $\mathsf{ct} = ([\mathbf{c}]_1, [\mathbf{c}']_1)$, the algorithm outputs $e([\mathbf{c}]_1^\top, [\mathbf{d}']_2)/e([\mathbf{c}']_1^\top, [\mathbf{d}]_2)$.

We will demonstrate that the experiment can be indistinguishably changed into non-committing experiment where the session key seskey^* is set to $[x]_T$ for randomly chosen $x \leftarrow \mathbb{Z}_p$ and the master secret key is computed so that it in fact maps the challenge KEM ciphertext ct^* to $\mathsf{seskey}^* = [x]_T$. For simplicity, we assume that the adversary has queried a single identity, denoted by id, and that the adversary's target identity is denoted as id^*. We begin with the following:

$$\mathsf{sk}_{\mathsf{id}} := ([s\mathbf{b}]_2, [\mathbf{k} + s(\mathbf{W}_1 + \mathsf{id} \cdot \mathbf{W}_2)^\top\mathbf{b}]_2)$$

$$\mathsf{ct} := ([r^*\mathbf{a}]_1, [r^*(\mathbf{W}_1 + \mathsf{id}^* \cdot \mathbf{W}_2)\mathbf{a}]_1) \quad \text{and} \quad \mathsf{seskey} := [r^*\mathbf{a}^\top\mathbf{k}]_T$$

By applying the SXDH assumption, we can replace $r^*\mathbf{a}$ and $s\mathbf{b}$ with a truly random elements $\mathbf{u}, \mathbf{v} \leftarrow \mathbb{Z}_p$.

$$\mathsf{sk}_{\mathsf{id}} := ([\mathbf{v}]_2, [\mathbf{k} + (\mathbf{W}_1 + \mathsf{id} \cdot \mathbf{W}_2)^\top\mathbf{v}]_2)$$

$$\mathsf{ct}^* := ([\mathbf{u}]_1, [(\mathbf{W}_1 + \mathsf{id}^* \cdot \mathbf{W}_2)\mathbf{u}]_1) \quad \text{and} \quad \mathsf{seskey} := [\mathbf{u}]_T$$

Next, we change the sampling method by sampling $\mathbf{W}_1 := \hat{\mathbf{W}}_1 + w_1\mathbf{W}_0$ and $\mathbf{W}_2 := \hat{\mathbf{W}}_2 + w_2\mathbf{W}_0$ where $\hat{\mathbf{W}}_1, \hat{\mathbf{W}}_2 \leftarrow \mathbb{Z}_p^{2\times 2}$ and $\mathbf{W}_0 := \mathbf{b}^\perp(\mathbf{a}^\perp)^\top/(\mathbf{b}^\perp)^\top\mathbf{a}^\perp$. This maintains the same distribution, but now the secret key and ciphertext involve w_1, w_2 as follows.

$$\mathsf{sk}_{\mathsf{id}} := ([\mathbf{v}]_2, [\mathbf{k} + (\hat{\mathbf{W}}_1 + \mathsf{id} \cdot \hat{\mathbf{W}}_2)^\top\mathbf{v} + t(w_1 + \mathsf{id}w_2)\mathbf{a}^\perp]_2)$$

$$\mathsf{ct}^* := ([\mathbf{u}]_1, [(\hat{\mathbf{W}}_1 + \mathsf{id}^* \cdot \hat{\mathbf{W}}_2)\mathbf{u} + r(w_1 + \mathsf{id}^*w_2)]_1) \quad \text{and} \quad \mathsf{seskey} := [\mathbf{u}]_T$$

where $\mathbf{u} = r'\mathbf{a} + r\mathbf{b}^\perp$ and $\mathbf{v} = t'\mathbf{a} + t\mathbf{b}^\perp$ such that $r', r, t', t \in \mathbb{Z}_p$, i.e., we can express $\mathbf{u}, \mathbf{v}$ in terms of $\mathbf{a}, \mathbf{b}^\perp$ because they are linearly independent with high probability. Since, $\mathsf{id} \neq \mathsf{id}^*$, the following holds.

$$\{ w_1 + \mathsf{id}^* w_2, w_1 + \mathsf{id} w_2 \} \equiv \{ w_1 + \mathsf{id}^* w_2, w \}$$

where w is chosen uniformly at random. Therefore, we can change to

$$\mathsf{sk}_{\mathsf{id}} := ([\mathbf{v}]_2, [\mathbf{k} + (\hat{\mathbf{W}}_1 + \mathsf{id} \cdot \hat{\mathbf{W}}_2)^\top \mathbf{v} + w\mathbf{a}^\perp]_2)$$

$$\mathsf{ct}^* := ([\mathbf{u}]_1, [(\hat{\mathbf{W}}_1 + \mathsf{id}^* \cdot \hat{\mathbf{W}}_2)\mathbf{u} + r(w_1 + \mathsf{id}^* w_2)]_1) \quad \text{and} \quad \mathsf{seskey} := [\mathbf{u}]_T$$

In the actual proof, multiple secret keys are involved, and this modification cannot be made if all the secret keys contain information about w_1 and w_2. However, by introducing additional hybrid steps and ensuring that at any moment only one secret key retains information about w_1 and w_2, we will carefully modify the secret keys one by one.

Now, we revert back to the original $\mathbf{W}_1, \mathbf{W}_2$ and use DDH assumption to reach

$$\mathsf{sk}_{\mathsf{id}} := ([s\mathbf{b}]_2, [\mathbf{k} + s(\mathbf{W}_1 + \mathsf{id} \cdot \mathbf{W}_2)^\top \mathbf{b} + w\mathbf{a}^\perp]_2)$$

$$\mathsf{ct}^* := ([\mathbf{u}]_1, [(\mathbf{W}_1 + \mathsf{id}^* \cdot \mathbf{W}_2)\mathbf{u}]_1) \quad \text{and} \quad \mathsf{seskey} := [\mathbf{u}]_T$$

We now sample $k_1, k_2 \leftarrow \mathbb{Z}_p$ and set $\mathbf{k} := \frac{k_1}{|\mathbf{a}|^2} \cdot \mathbf{a} + \frac{k_2}{|\mathbf{a}^\perp|^2} \cdot \mathbf{a}^\perp$. This modification results in $\mathsf{mpk} = [k_1]_T$. Note that this is merely a conceptual change. We now modify the secret key as follows:

$$\mathsf{sk}_{\mathsf{id}} := ([s\mathbf{b}]_2, [\frac{k_1}{|\mathbf{a}|^2} \cdot \mathbf{a} + s(\mathbf{W}_1 + \mathsf{id} \cdot \mathbf{W}_2)^\top \mathbf{b} + w\mathbf{a}^\perp]_2)$$

$$\mathsf{ct}^* := ([\mathbf{u}]_1, [(\mathbf{W}_1 + \mathsf{id}^* \cdot \mathbf{W}_2)\mathbf{u}]_1) \quad \text{and} \quad \mathsf{seskey} := [\mathbf{u}]_T$$

This change is indistinguishable because w is chosen uniformly at random. We can express $\mathbf{u} = u_1\mathbf{a} + u_2\mathbf{a}^\perp$ where $u_1, u_2 \leftarrow \mathbb{Z}_p$ because $\mathbf{u}$ is chosen uniformly at random. Now, given a uniformly random $x \leftarrow \mathbb{Z}_p$, we can program the master secret key as $k_2 = \dfrac{x - u_1 k_1}{u_2}$. For verification, let us check that generating a secret key for id^* and decrypting ct^* would yield $[x]_T$. A secret key for id^* would be

$$\mathsf{sk}_{\mathsf{id}^*} := ([s^*\mathbf{b}]_2, [\mathbf{k} + s^*(\mathbf{W}_1 + \mathsf{id}^* \cdot \mathbf{W}_2)^\top \mathbf{b}]_2)$$

and decryption would result in

$$\begin{aligned}
\frac{e([\mathbf{u}^\top]_1, [\mathbf{k} + s^*(\mathbf{W}_1 + \mathsf{id}^* \cdot \mathbf{W}_2)^\top \mathbf{b}]_2)}{e([((\mathbf{W}_1 + \mathsf{id}^* \cdot \mathbf{W}_2)\mathbf{u})^\top]_1, [s^*\mathbf{b}]_2)} &= \frac{e([\mathbf{u}^\top]_1, [\mathbf{k}]_2) \cdot e([\mathbf{u}^\top]_1, [s^*(\mathbf{W}_1 + \mathsf{id}^* \cdot \mathbf{W}_2)^\top \mathbf{b}]_2)}{e([((\mathbf{W}_1 + \mathsf{id}^* \cdot \mathbf{W}_2)\mathbf{u})^\top]_1, [s^*\mathbf{b}]_2)} \\
&= e([\mathbf{u}^\top]_1, [\mathbf{k}]_2) = [\mathbf{u}^\top \mathbf{k}]_T \\
&= [(u_1\mathbf{a} + u_2\mathbf{a}^\perp)^\top (k_1\mathbf{a} + k_2\mathbf{a}^\perp)]_T \\
&= [u_1k_1 + u_2k_2]_T = [x]_T
\end{aligned}$$

RNC-IBE from Batch Encryption

Let us start by reviewing the concept of batch encryption. A batch encryption scheme is a form of public key encryption where the key generation process is a projection—this means that the secret key is used to produce a shorter public key. When the secret key has a length of n, the scheme allows the simultaneous encryption of $n \times 2$ messages. During decryption, only one message from each pair can be recovered, and this is determined by the corresponding bit in the secret key.

To elaborate, the setup algorithm $\mathsf{BE.Setup}$ takes as input a secret key $\mathsf{sk} \in \{0,1\}^n$ and outputs a public key pk. The encryption algorithm encrypts a matrix $M \in \mathcal{M}^{n \times 2}$ using the public key to generate a ciphertext ct. Here, $\mathcal{M}$ is an appropriate message space. The decryption algorithm takes as input the secret key sk and the ciphertext ct and outputs a vector $m \in \mathcal{M}^n$ such that $m_i = M_{i,\mathsf{sk}[i]}$, for all $i \in [n]$.

The RNC-IBE scheme uses the batch encryption and garbling scheme as follows. Let $d = \log(\mathrm{poly}(\lambda))$ be an integer and $T = 2^d$ be a polynomial in the security parameter that denotes the number of identities the scheme supports. The setup algorithm generates T pairs of NCE public and secret keys, denoted as $\{\mathsf{nce.pk}_j, \mathsf{nce.sk}_j\}_{j\in[T]}$, where each pair corresponds to a different identity. These keys together form the master secret key of the RNC-IBE scheme. The master public key is a public key of the batch encryption scheme generated by $\mathsf{be.pk} \leftarrow \mathsf{BE.Setup}(\{\mathsf{nce.pk}_j\}_{j\in[T]})$, i.e., $\{\mathsf{nce.pk}_j\}_{j\in[T]}$ is the secret key associated with $\mathsf{be.pk}$.

The secret key for the i^{th} identity consists of all the NCE public keys $\{\mathsf{nce.pk}_j\}_{j\in[T]}$ along with the i^{th} secret key $\mathsf{nce.sk}_i$. To encrypt a message m for the i^{th} identity, the encryption algorithm generates T garbled circuit labels as follows:

- For the i^{th} identity, it generates $(\tilde{\mathcal{C}}^{(i)}, \{\mathsf{lab}_{j,b}^{(i)}\}) \leftarrow \mathsf{GC.Grbl}(\mathsf{NCE.Enc}(\cdot, m))$.
- For the remaining identity, it generates $(\tilde{\mathcal{C}}^{(k)}, \{\mathsf{lab}_{j,b}^{(k)}\}) \leftarrow \mathsf{GC.Grbl}(\mathsf{NCE.Enc}(\cdot, m^{(k)}))$ where $m^{(k)}$ is randomly generated.

A matrix $M \in \{0,1\}^{nT\times 2}$ is then constructed such that $M[k \cdot n + j, b] = \mathsf{lab}_{j,b}^{(k)}$. The batch encryption scheme is then used to produce $\mathsf{be.ct} \leftarrow \mathsf{BE.Enc}(\mathsf{be.pk}, M)$. The final ciphertext is $\mathsf{ct} := (\{\tilde{\mathcal{C}}^{(k)}\}_k, \mathsf{be.ct})$.

The simulation works as follows. First, the simulator generates the NCE public keys and corresponding ciphertexts using the NCE simulators and constructs the master public key. The simulator then produces the non-committing ciphertext by simulating all garbled circuit labels, i.e., $(\tilde{\mathcal{C}}^{(k)}, \{\mathsf{lab}_j^{(k)}\}) \leftarrow \mathsf{GC.Sim}(\mathsf{nce.ct}^{(k)})$.

Upon receiving the target identity i and target message m, the simulator uses the NCE simulator to simulate $\mathsf{nce.sk}_i$ such that the ciphertext $\mathsf{nce.ct}_i$ will decrypt to m using $\mathsf{nce.sk}_i$. For the other identities, the simulator uses the NCE simulators to simulate $\mathsf{nce.sk}_k$ on random messages $m^{(k)}$. For more details, refer Sec. 6.

3 Preliminaries

Let PPT denote probabilistic polynomial time. We denote the set of all positive integers up to n as $[n] := \{1, \ldots, n\}$ and $[n]_0 := \{0, 1, \ldots, n\}$. In addition, we use $[i, j]$ to denote the set of all non-negative integers between i and j including i, j, i.e., $[i, j] := \{i, i+1, \ldots, j\}$. For any two binary string x, y, we use the notation $x \preceq y$ (or $x \in \mathsf{prefix(y)}$) to imply that x is a prefix of y and $x \| y$ to denote x concatenated with y. And $x[i, j]$ denotes the substring $x[i] \| x[i+1] \| \ldots \| x[j]$ when $i \leq j$ and $x[i, j] = \epsilon$ where $i > j$. Throughout this paper, unless specified, all polynomials we consider are positive polynomials. For any finite set S, $x \leftarrow S$ denotes a uniformly random element x from the set S. Suppose, S is an ordered set of n element, i.e., $S = (a_1, \ldots, a_n)$, then we use the notation $(b_1, \ldots, b_n) \leftarrow S$ to denote that b_i is assigned the value a_i, for all $i \in [n]$.

We use a pairing group $e : \mathbb{G}_1 \times \mathbb{G}_2 \to \mathbb{G}_T$, where $\mathbb{G}_1, \mathbb{G}_2, \mathbb{G}_T$ are all cyclic groups of prime order p, generated respectively by g_1, g_2, and $e(g_1, g_2)$, where e is a non-degenerate bilinear map, that is, for all $a, b \in \mathbb{Z}_p, e(g_1^a, g_2^b) = e(g_1, g_2)^{ab}$.

We use bracket notations, where for all exponents $a \in \mathbb{Z}_p$ and all groups $s \in \{1, 2, T\}$, we denote by $[a]_s$ the group element g_s^a. We generalize this notation for vectors and matrices as follows. For any $\mathbf{A} \in \mathbb{Z}_p^{n \times m}$, $[\mathbf{A}]_s := g_s^{\mathbf{A}}$ denotes an $n \times m$ matrix with elements from $\mathbb{G}_s$ such that (i, j)-th element is $[\mathbf{A}_{i,j}]_s$. Since, $[\cdot]_s$ is a linear functions, we can $[\mathbf{AB}]_s$ from $[\mathbf{A}]_s$ and $[\mathbf{B}]_s$ for any matrices $\mathbf{A}, \mathbf{B}$ over $\mathbb{Z}_p$. Also, given $[\mathbf{A}]_1$ and $[\mathbf{B}]_2$, we define $[\mathbf{AB}]_T = e([\mathbf{A}]_1, [\mathbf{B}]_2)$.

3.1 Hardness Assumptions

Decisional Diffie-Hellman: The Decisional Diffie-Hellman (DDH) assumption with respect to a group $\mathbb{G}$, is that for every PPT adversary $\mathcal{A}$ it holds that

$$\Big| \Pr_{\substack{(\mathbb{G},g,q) \leftarrow \mathcal{G}(1^\lambda) \\ a,b \leftarrow \mathbb{Z}_q}} [\mathcal{A}(1^\lambda, (\mathbb{G}, g, q), g^a, g^b, g^{a \cdot b}) = 1] - \Pr_{\substack{(\mathbb{G},g,q) \leftarrow \mathcal{G}(1^\lambda) \\ a,b,u \leftarrow \mathbb{Z}_q}} [\mathcal{A}(1^\lambda, (\mathbb{G}, g, q), g^a, g^b, g^u) = 1] \Big| = \mathsf{negl}(\lambda)$$

Computational Diffie-Hellman: The Computational Diffie-Hellman (CDH) assumption with respect to a group generator $\mathcal{G}$, is that for every PPT adversary $\mathcal{A}$ it holds that

$$\Pr_{\substack{(\mathbb{G},g,q) \leftarrow \mathcal{G}(1^\lambda) \\ a,b \leftarrow \mathbb{Z}_q}} [\mathcal{A}(1^\lambda, (\mathbb{G}, g, q), g^a, g^b) = g^{a \cdot b}] = \mathsf{negl}(\lambda)$$

Symmetric eXternal Diffie-Hellman: The Symmetric eXternal Diffie-Hellman (SXDH) assumption holds for a pairing group $(\mathbb{G}_1, \mathbb{G}_2, \mathbb{G}_T, p, g_1, g_2, g_T, e) \leftarrow \mathcal{G}(1^\lambda)$ if DDH holds for $\mathbb{G}_1$ and $\mathbb{G}_2$.

Learning with Error: The Learning with Error assumption holds if for all PPT adversary $\mathcal{A}$, there exists a negligible function $\mathsf{negl}(\cdot)$ such that for all $\lambda \in \mathbb{N}$, $k = k(\lambda), q = q(\lambda)$ and D_σ being an error distribution,

$$|\Pr[\mathcal{A}(\mathbf{A}, \mathbf{sA} + \mathbf{e}) = 1] - \Pr[\mathcal{A}(\mathbf{A}, \mathbf{u}) = 1]| = \mathsf{negl}(\lambda)$$

for all $n \in \mathbb{N}$ such that $\mathbf{A} \leftarrow \mathbb{Z}_q^{k \times n}, \mathbf{s} \leftarrow \mathbb{Z}_q^k, \mathbf{e} \leftarrow D_\sigma^n$ and $\mathbf{u} \leftarrow \mathbb{Z}_q^n$.

3.2 Batch Encryption

Let $B = B(\lambda, n)$ be a global parameter. A batch encryption (BE) scheme consists of the following four algorithms.

- $\mathsf{Params}(1^\lambda, 1^n)$: The algorithm takes as input the security parameter 1^λ and a parameter 1^n and outputs a public parameter pp.
- $\mathsf{Setup}(\mathsf{pp}, \mathsf{sk})$: The setup algorithm takes as input a public parameter pp and a secret key $\mathsf{sk} \in [B]^n$ and outputs a public key pk.
- $\mathsf{Enc}(\mathsf{pk}, M)$: The encryption algorithm takes as input a public key pk and a matrix $M \in \{0,1\}^{n \times B}$ and outputs a ciphertext ct.
- $\mathsf{Dec}(\mathsf{sk}, \mathsf{ct})$: The decryption algorithm takes as input a secret key sk and a ciphertext ct and outputs either outputs $\perp$ or a vector $m \in \{0,1\}^n$.

Correctness. For correctness, we require that for all $\lambda \in \mathbb{N}, n, B \in \mathbb{N}, \mathsf{sk} \in [B]^n, M \in \{0,1\}^{n \times B}$,

$$\Pr\left[\mathsf{Dec}(\mathsf{sk}, \mathsf{ct}) = m \mid \mathsf{ct} \leftarrow \mathsf{Enc}(\mathsf{pk}, M), \mathsf{pk} \leftarrow \mathsf{Setup}(\mathsf{pp}, \mathsf{sk}), \mathsf{pp} \leftarrow \mathsf{Params}(1^\lambda, 1^n)\right] = 1$$

where the probability is over the random bits used in the $\mathsf{Params}, \mathsf{Setup}, \mathsf{Enc}$ algorithm and $m[i] = M[i, \mathsf{sk}[i]], \forall i \in [n]$.

IND-based Security. Consider the following experiment with an adversary $\mathcal{A}$.

- **Initialization Phase:** The adversary takes 1^λ as input, and sends $1^n, x \in [B]^n$ to the challenger. The challenger runs $\mathsf{pp} \leftarrow \mathsf{Params}(1^\lambda, 1^n)$ and sends pk to $\mathcal{A}$.
- **Challenge Phase:** $\mathcal{A}$ outputs two message $M_0, M_1 \in \{0,1\}^{n \times B}$ to the challenger. The challenger computes $\mathsf{pk} \leftarrow \mathsf{Setup}(\mathsf{pp}, \mathsf{sk})$ and randomly chooses $b \in \{0,1\}$. It computes a ciphertext $\mathsf{ct}^* = \mathsf{Enc}(\mathsf{pk}, M_b)$ and sends $(\mathsf{pk}, \mathsf{ct}^*)$ to $\mathcal{A}$.
- **Response Phase:** $\mathcal{A}$ outputs b'. $\mathcal{A}$ wins the experiment if $b = b'$.

Definition 1. *An BE scheme satisfies indistinguishability-based security if for all PPT adversaries $\mathcal{A}$, there exists a negligible function $\mathsf{negl}(\cdot)$ such that for all $\lambda \in \mathbb{N}$,*

$$\Pr[\mathcal{A} \textit{ wins in the above experiment}] \leq \frac{1}{2} + \mathsf{negl}(\lambda)$$

In this work, we will require an *oblivious* batch encryption which has the following properties.

1. Params outputs pp without using any randomness other than pp itself.
2. Setup is a deterministic algorithm.

In other words, the randomness used in Params and Setup is pp only. We emphasis that the constructions given in [15] for blind batch encryptions are oblivious.

Theorem 7 ([15]). *Assuming the hardness of $\mathcal{C}$ where $\mathcal{C} \in \{\mathsf{CDH}, \mathsf{LWE}\}$, there exists a secure oblivious BE scheme such that the size of the public key is a polynomial in λ, i.e., $|\mathsf{pk}| = \mathrm{poly}(\lambda)$.*

3.3 Non-Committing Encryption

A non-committing encryption (NCE) scheme consists of the following algorithms.

$\mathsf{Setup}(1^\lambda; r_{\mathsf{Setup}})$: The setup algorithm takes as input the security parameter 1^λ . Using the random coins r_{Setup}, it outputs the public key pk and secret key sk.

$\mathsf{Enc}(\mathsf{pk}, m; r_{\mathsf{Enc}})$: The encryption algorithm takes as input a master public key pk, a message m and using random coins r_{Enc} outputs a ciphertext ct.

$\mathsf{Dec}(\mathsf{sk}, \mathsf{ct})$: The decryption algorithm takes as input a secret key sk and a ciphertext ct and outputs either a message m or $\perp$.

$\mathsf{Sim}_1(1^\lambda)$: The first simulator takes as input the security parameter $1^\lambda$and outputs a public key pk, a ciphertext ct^* and a state st_1.

$\mathsf{Sim}_2(\mathsf{st}_1, m)$: The second simulator takes as input a state st_1 and a message m and outputs $(r_{\mathsf{Enc}}, r_{\mathsf{Setup}})$.

Correctness. For correctness, we require that there exists a negligible function $\mathsf{negl}(\cdot)$ such that for all $\lambda \in \mathbb{N}, T \in \mathbb{N}$ and $(\mathsf{pk}, \mathsf{sk})$ output by $\mathsf{Setup}(1^\lambda)$, any message m,

$$\Pr_r[\mathsf{Dec}(\mathsf{sk}, \mathsf{ct}) = m \mid \mathsf{ct} = \mathsf{Enc}(\mathsf{pk}, m; r))] = 1 - \mathsf{negl}(\lambda)$$

where r is sampled uniformly at random.

Non-Committing Security. Consider the following two experiments with an adversary $\mathcal{A}$.

Real World:

- **Initialization Phase:** The challenger computes $(\mathsf{pk}, \mathsf{sk}) \leftarrow \mathsf{Setup}(1^\lambda; r_{\mathsf{Setup}})$ and sends pk to $\mathcal{A}$.
- **Challenge Phase:** The adversary $\mathcal{A}$ sends m^* to the challenger. The challenger computes $\mathsf{ct}^* \leftarrow \mathsf{Enc}(\mathsf{pk}, m^*; r_{\mathsf{Enc}})$ and returns $(\mathsf{ct}^*, r_{\mathsf{Enc}}, r_{\mathsf{Setup}})$ to $\mathcal{A}$.
- **Response Phase:** $\mathcal{A}$ outputs b.

Simulated World:

- **Initialization Phase:** The challenger computes $(\mathsf{pk}, \mathsf{ct}^*, \mathsf{st}_1) \leftarrow \mathsf{Sim}_1(1^\lambda)$ and sends mpk to $\mathcal{A}$.

- **Challenge Phase:** The adversary $\mathcal{A}$ sends m^* to the challenger. The challenger computes $(r_{\mathsf{Enc}}, r_{\mathsf{Setup}}) \leftarrow \mathsf{Sim}_2(\mathsf{st}_1, m^*)$ and returns $(\mathsf{ct}^*, r_{\mathsf{Enc}}, r_{\mathsf{Setup}})$ to $\mathcal{A}$.
- **Response Phase:** $\mathcal{A}$ outputs b.

Let p_{real} and p_{sim} be the probabilities with which $\mathcal{A}$ outputs 0 in the real world and simulated world, respectively.

Definition 2. *An NCE scheme is said to be secure if for all PPT adversaries $\mathcal{A}$, there exists a negligible function $\mathsf{negl}(\cdot)$ such that, for all $\lambda \in \mathbb{N}$,*

$$|p_{\text{real}} - p_{\text{sim}}| = \mathsf{negl}(\lambda)$$

Theorem 8 ([14, 72]). *Assuming the hardness of* LWE *and* DDH*, there exists a secure non-committing encryption scheme.*

3.4 Incompressible Secret Key Encryption

An incompressible secret key encryption scheme $\mathsf{IncSKE} = (\mathsf{Setup}, \mathsf{Enc}, \mathsf{Dec})$ with message space $\{\mathcal{M}_\lambda\}_\lambda$ consists of the following PPT algorithms.

- $\mathsf{Setup}(1^\lambda, 1^S)$: The setup algorithm is a randomized algorithm that takes as input the security parameter 1^λ, a parameter 1^S and outputs a secret key sk.
- $\mathsf{Enc}(\mathsf{sk}, m)$: The encryption algorithm is a randomized algorithm that takes as input a secret key sk and a message $m \in \mathcal{M}_\lambda$ and outputs a ciphertext ct.
- $\mathsf{Dec}(\mathsf{sk}, \mathsf{ct})$: The decryption algorithm takes as input a secret key sk and a ciphertext ct and outputs either a message $m \in \mathcal{M}_\lambda$ or $\perp$.

Correctness. For correctness, we require that for all $\lambda \in \mathbb{N}, S \in \mathbb{N}, m \in \mathcal{M}_\lambda$ and $\mathsf{sk} \leftarrow \mathsf{Setup}(1^\lambda, 1^S)$,

$$\Pr[\mathsf{Dec}(\mathsf{sk}, \mathsf{Enc}(\mathsf{sk}, m)) = m] = 1$$

where the probability is over the random bits used in the encryption algorithm.

Definition 3 (Incompressible SKE Security). *Consider the following experiment with an adversary $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2)$.*

- ***Initialization Phase:*** *$\mathcal{A}_1$ on input 1^λ, outputs an upper bound on the state size 1^S. The challenger runs $\mathsf{sk} \leftarrow \mathsf{Setup}(1^\lambda, 1^S)$.*
- ***Challenge Phase:*** *$\mathcal{A}_1$ outputs a message (m_0, m_1), along with an auxiliary information aux. The challenger randomly chooses $b \in \{0, 1\}$. It computes a ciphertext $\mathsf{ct}^* = \mathsf{Enc}(\mathsf{sk}, m_b)$ and sends it to $\mathcal{A}_1$.*
- ***First Response Phase:*** *$\mathcal{A}_1$ computes a state st such that $|\mathsf{st}| \leq S$.*
- ***Second Response Phase:*** *$\mathcal{A}_2$ receives $(\mathsf{sk}, \mathsf{aux}, \mathsf{st})$ and outputs b'. $\mathcal{A}$ wins the experiment if $b = b'$.*

An SKE scheme is said to be incompressible secure if for all PPT adversaries $\mathcal{A}$, there exists a negligible function $\mathsf{negl}(\cdot)$ *such that for all* $\lambda \in \mathbb{N}$,

$$\Pr[\mathcal{A} \textit{ wins in the above experiment}] \leq \frac{1}{2} + \mathsf{negl}(\lambda)$$

The rate of a scheme is defined as the ratio between the size of a message and the size of a ciphertext, i.e., $\frac{|m|}{|\mathsf{ct}|}$. We say a scheme has rate-1 if $\frac{|m|}{|\mathsf{ct}|} = |m| - o(|m|)$.

Theorem 9 ([16]). *Assume the hardness of* LWE *or* DCR*, there exists a rate-1 incompressible SKE whose secret key size is* $|\mathsf{sk}| = n(1 + o(1)) + \mathrm{poly}(\lambda)$ *where* n *is the size of the message.*

Theorem 10 ([32]). *There exists a rate-*$\frac{1}{2}$ *incompressible SKE from one-way functions whose secret key size is* $|\mathsf{sk}| = \mathrm{poly}(\lambda)$.

3.5 Incompressible IBE

In this section, we define the strong version of the incompressible security game for IBE scheme[6] where Setup takes an additional input 1^S and the second adversary obtains the *master secret key*. The game is played against two adversaries $\mathcal{A}_1, \mathcal{A}_2$. The first adversary $\mathcal{A}_1$ will be provided with the complete challenge ciphertext and produce a compressed version of it. The second adversary $\mathcal{A}_2$ is provided with the master public key, compressed challenge ciphertext which was created by $\mathcal{A}_1$ and certain secret keys.

Definition 4. *(Incompressible IBE Security).* Let $\mathsf{IBE} = (\mathsf{Setup}, \mathsf{KeyGen}, \mathsf{Enc}, \mathsf{Dec})$ be an IBE scheme in which the setup algorithm takes an additional parameter 1^S as input. Consider the following experiment with an adversary $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2)$.

Initialization Phase: $\mathcal{A}_1$ on input 1^λ, outputs an upper bound on the state size 1^S. The challenger runs $(\mathsf{msk}, \mathsf{mpk}) \leftarrow \mathsf{Setup}(1^\lambda, 1^S; r_{\mathsf{Setup}})$ and sends mpk to $\mathcal{A}_1$.

Pre-Challenge Query Phase: In this phase, $\mathcal{A}_1$ is allowed to make polynomially many key queries. For each query id sent to the challenger, the challenger computes $\mathsf{sk}_{\mathsf{id}} \leftarrow \mathsf{KeyGen}(\mathsf{msk}, \mathsf{id})$ and returns $\mathsf{sk}_{\mathsf{id}}$ to $\mathcal{A}_1$.

Challenge Phase: $\mathcal{A}_1$ outputs two messages m_0, m_1, an identity id^* along with an auxiliary information aux. If there exists a query for id^* made by $\mathcal{A}_1$, the challenger aborts the game. Else, it randomly chooses $b \in \{0, 1\}$ and computes a ciphertext $\mathsf{ct}^* = \mathsf{Enc}(\mathsf{mpk}, \mathsf{id}^*, m_b)$ and sends it to $\mathcal{A}_1$.

Post-Challenge Query Phase: This is similar to the pre-challenge query phase. The adversary $\mathcal{A}_1$ is allowed to send polynomially many key queries. For each query id, if $\mathsf{id}^* = \mathsf{id}$, the challenger sends $\perp$. Else, computes $\mathsf{sk}_{\mathsf{id}} \leftarrow \mathsf{KeyGen}(\mathsf{msk}, \mathsf{id})$ and returns $\mathsf{sk}_{\mathsf{id}}$ to $\mathcal{A}_1$.

[6] The only difference between a standard IBE scheme and an incompressible IBE scheme is that the setup algorithm takes an additional parameter 1^S that specifies the compression size.

First Response Phase: $\mathcal{A}_1$ computes a state st such that $|st| \leq S$.
Second Response Phase: $\mathcal{A}_2$ receives $(\mathsf{mpk}, \mathsf{msk}, aux, st)$. Finally, $\mathcal{A}_2$ outputs b'. $\mathcal{A}$ wins the experiment if $b = b'$.

An IBE scheme is said to be **strong** incompressible secure if for all PPT adversaries $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2)$, there exists a negligible function $\mathsf{negl}(\cdot)$ such that for all $\lambda \in \mathbb{N}$,

$$\Pr[\mathcal{A} \text{ wins in the above experiment}] \leq \frac{1}{2} + \mathsf{negl}(\lambda)$$

Definition 5. *An IBE scheme is said to be* ***super-strong*** *incompressible secure if for all PPT adversaries $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2)$, there exists a negligible function $\mathsf{negl}(\cdot)$ such that for all $\lambda \in \mathbb{N}$,*

$$\Pr[\mathcal{A} \textit{ wins in the above experiment}] \leq \frac{1}{2} + \mathsf{negl}(\lambda)$$

provided the second adversary $\mathcal{A}_2$ receives the random coins r_{Setup} used in the setup algorithm instead of $\mathsf{mpk}, \mathsf{msk}$.

4 Receiver Non-Committing Identity-Based Primitives: Definitions

In this section, we will present the definitions of a receiver non-committing identity-based encryption (RNC-IBE) scheme and receiver non-committing identity-based key-encapsulation mechanism (RNC-IB-KEM) where in the challenge phase, the challenger returns the master secret key msk instead of the random coins used in the Setup and Enc algorithms.

4.1 Receiver Non-Committing Identity-Based Encryption

A receiver non-committing identity-based encryption (RNC-IBE) scheme consists of the following algorithms.

$\mathsf{Setup}(1^\lambda, 1^n)$: The setup algorithm takes as input the security parameter 1^λ and the length of the identities 1^n (in some cases the number of identities 1^T). It outputs the master public key mpk and master secret key msk.
$\mathsf{Keygen}(\mathsf{msk}, \mathsf{id})$: The key generation algorithm takes as input a master secret key msk and an identity id and outputs a secret key $\mathsf{sk}_{\mathsf{id}}$.
$\mathsf{Enc}(\mathsf{mpk}, \mathsf{id}, m)$: The encryption algorithm takes as input a master public key mpk, an identity id and a message m and outputs a ciphertext ct.
$\mathsf{Dec}(\mathsf{sk}, \mathsf{ct})$: The decryption algorithm takes as input a secret key sk and a ciphertext ct and outputs either a message m or $\perp$.
$\mathsf{Sim}_1(1^\lambda, 1^n)$: The first simulator takes as input the security parameter 1^λ and the length of the identities 1^n (in some cases the number of identities 1^T) and outputs a master public key mpk and a state st_1.

$\mathsf{Sim}_2(\mathsf{st}_1, \mathsf{id})$: The second simulator is a stateful algorithm with an internal state st_2 that takes as input a state st_1 and an identity id and outputs a secret key $\mathsf{sk}_{\mathsf{id}}$ and updates st_2.

$\mathsf{Sim}_3(\mathsf{st}_2, \mathsf{id}^*)$: The third simulator takes as input a state st_2 and an identity id^* and outputs a ciphertext ct^* and a state st_3.

$\mathsf{Sim}_4(\mathsf{st}_3, m^*)$: The fourth simulator takes as input a state st_3 and a message m^* and outputs a master secret key msk.

Correctness. For correctness, we require that there exists a negligible function $\mathsf{negl}(\cdot)$ such that for all $\lambda \in \mathbb{N}, n \in \mathbb{N}$ and $(\mathsf{mpk}, \mathsf{msk}) \leftarrow \mathsf{Setup}(1^\lambda, 1^n)$, any identity id and message m,

$$\Pr_{r_1, r_2}\left[\mathsf{Dec}(\mathsf{sk}_{\mathsf{id}}, \mathsf{ct}) = m \mid \begin{array}{c} \mathsf{ct} = \mathsf{Enc}(\mathsf{mpk}, \mathsf{id}, m; r_1) \\ \mathsf{sk}_{\mathsf{id}} = \mathsf{Keygen}(\mathsf{msk}, \mathsf{id}; r_2) \end{array}\right] = 1 - \mathsf{negl}(\lambda)$$

Security. Consider the following two experiments with an adversary $\mathcal{A}$.

Real World:

- **Initialization Phase:** $\mathcal{A}$ on input 1^λ, outputs 1^T. The challenger computes $(\mathsf{mpk}, \mathsf{msk}) \leftarrow \mathsf{Setup}(1^\lambda, 1^T)$ and sends mpk to $\mathcal{A}$.
- **Pre-Challenge Query Phase:** In this phase, $\mathcal{A}$ is allowed to make multiple queries id. For each id, the challenger returns $\mathsf{sk}_{\mathsf{id}} \leftarrow \mathsf{Keygen}(\mathsf{msk}, \mathsf{id})$ to $\mathcal{A}$.
- **Challenge Phase:** The adversary $\mathcal{A}$ sends m^*, id^* to the challenger where id^* was never queried in the pre-challenge query phase. The challenger computes $\mathsf{ct}^* \leftarrow \mathsf{Enc}(\mathsf{mpk}, \mathsf{id}^*, m^*)$ and returns $(\mathsf{msk}, \mathsf{ct}^*)$ to $\mathcal{A}$.
- **Response Phase:** $\mathcal{A}$ outputs b.

Simulated World:

- **Initialization Phase:** $\mathcal{A}$ on input 1^λ, outputs 1^T. The challenger computes $(\mathsf{mpk}, \mathsf{st}_1) \leftarrow \mathsf{Sim}_1(1^\lambda, 1^T)$ and sends mpk to $\mathcal{A}$.
- **Pre-Challenge Query Phase:** In this phase, $\mathcal{A}$ is allowed to make multiple queries id. For each id, the challenger returns $\mathsf{sk}_{\mathsf{id}} \leftarrow \mathsf{Sim}_2(\mathsf{st}_1, \mathsf{id})$ to $\mathcal{A}$.
- **Challenge Phase:** The adversary $\mathcal{A}$ sends m^*, id^* to the challenger where id^* was never queried in the pre-challenge query phase. The challenger computes $(\mathsf{ct}^*, \mathsf{st}_3) \leftarrow \mathsf{Sim}_3(\mathsf{st}_2, \mathsf{id}^*)$ and $\mathsf{msk} \leftarrow \mathsf{Sim}_4(\mathsf{st}_3, m^*)$. It returns $(\mathsf{msk}, \mathsf{ct}^*)$ to $\mathcal{A}$.
- **Response Phase:** $\mathcal{A}$ outputs b.

It is important to note that in the challenge phase, the adversary obtains only the master secret key and the challenge ciphertext, and not the randomness used by the Setup algorithm or Enc algorithm to generate the challenge ciphertext. Let p_{real} and p_{sim} be the probabilities with which $\mathcal{A}$ outputs 0 in the real world and simulated world, respectively.

Definition 6. *An RNC-IBE scheme is said to be* ***adaptive*** *secure if for all PPT adversaries $\mathcal{A}$, there exists a negligible function $\mathsf{negl}(\cdot)$ such that, for all $\lambda \in \mathbb{N}$,*

$$|p_{\text{real}} - p_{\text{sim}}| = \mathsf{negl}(\lambda)$$

Definition 7. *An RNC-IBE scheme is said to be* ***selectively*** *secure if for all PPT adversaries $\mathcal{A}$, there exists a negligible function* $\mathsf{negl}(\cdot)$ *such that, for all* $\lambda \in \mathbb{N}$,

$$|p_{\text{real}} - p_{\text{sim}}| = \mathsf{negl}(\lambda)$$

provided the adversary commits to the challenge identity id^* *at the beginning of the game and* Sim_1 *addition takes* id^* *also as an additional input.*

4.2 Receiver Non-Committing Identity-Based Key-Encapsulation Mechanism

A receiver non-committing identity-based key-encapsulation mechanism (RNC-IB-KEM) consists of the following algorithms.

$\mathsf{Setup}(1^\lambda, 1^n)$: The setup algorithm takes as input the security parameter 1^λ and the length of the identities 1^n (in some cases the number of identities 1^T). It outputs the master public key mpk and master secret key msk.

$\mathsf{Keygen}(\mathsf{msk}, \mathsf{id})$: The key generation algorithm takes as input a master secret key msk and an identity id and outputs a secret key $\mathsf{sk}_{\mathsf{id}}$.

$\mathsf{Encap}(\mathsf{mpk}, \mathsf{id})$: The encryption algorithm takes as input a master public key mpk and an identity id and outputs a ciphertext ct and a session key seskey.

$\mathsf{Decap}(\mathsf{sk}, \mathsf{ct})$: The decryption algorithm takes as input a secret key sk and a ciphertext ct and outputs either a session key seskey or $\perp$.

$\mathsf{Sim}_1(1^\lambda, 1^n)$: The first simulator takes as input the security parameter 1^λ and the length of the identities 1^n (in some cases the number of identities 1^T) and outputs a master public key mpk and a state st_1.

$\mathsf{Sim}_2(\mathsf{st}_1, \mathsf{id})$: The second simulator is a stateful algorithm with an internal state st_2 that takes as input a state st_1 and an identity id and outputs a secret key $\mathsf{sk}_{\mathsf{id}}$ and updates st_2.

$\mathsf{Sim}_3(\mathsf{st}_2, \mathsf{id}^*)$: The third simulator takes as input a state st_2, an identity id^* and outputs a ciphertext ct^* and a state st_3.

$\mathsf{Sim}_4(\mathsf{st}_3, \mathsf{seskey})$: The third simulator takes as input a state st_3, a session key seskey and outputs a ciphertext msk.

Correctness. For correctness, we require that there exists a negligible function $\mathsf{negl}(\cdot)$ such that for all $\lambda \in \mathbb{N}, n \in \mathbb{N}$ and $(\mathsf{mpk}, \mathsf{msk}) \leftarrow \mathsf{Setup}(1^\lambda, 1^n)$, any identity id,

$$\Pr_{r_1, r_2}\left[\mathsf{Decap}(\mathsf{sk}_{\mathsf{id}}, \mathsf{ct}) = \mathsf{seskey} \ : \ \begin{array}{c}(\mathsf{ct}, \mathsf{seskey}) = \mathsf{Encap}(\mathsf{mpk}, \mathsf{id}; r_1) \\ \mathsf{sk}_{\mathsf{id}} \leftarrow \mathsf{Keygen}(\mathsf{msk}, \mathsf{id}; r_2)\end{array}\right] = 1 - \mathsf{negl}(\lambda)$$

Security. Consider the following two experiments with an adversary $\mathcal{A}$.

Real World:

– **Initialization Phase:** $\mathcal{A}$ on input 1^λ, outputs 1^T. The challenger computes $(\mathsf{mpk}, \mathsf{msk}) \leftarrow \mathsf{Setup}(1^\lambda, 1^T)$ and sends mpk to $\mathcal{A}$.

- **Pre-Challenge Query Phase:** In this phase, $\mathcal{A}$ is allowed to make multiple queries id. For each id, the challenger returns $\mathsf{sk}_{\mathsf{id}} \leftarrow \mathsf{Keygen}(\mathsf{msk}, \mathsf{id})$ to $\mathcal{A}$.
- **Challenge Phase:** The adversary $\mathcal{A}$ sends id^* to the challenger where id^* was never queried in the pre-challenge query phase. The challenger computes $(\mathsf{ct}^*, \mathsf{seskey}) \leftarrow \mathsf{Enc}(\mathsf{mpk}, \mathsf{id}^*)$ and returns $(\mathsf{msk}, \mathsf{ct}^*, \mathsf{seskey})$ to $\mathcal{A}$.
- **Response Phase:** $\mathcal{A}$ outputs b.

Simulated World:

- **Initialization Phase:** $\mathcal{A}$ on input 1^λ, outputs 1^T. The challenger computes $(\mathsf{mpk}, \mathsf{st}_1) \leftarrow \mathsf{Sim}_1(1^\lambda, 1^T)$ and sends mpk to $\mathcal{A}$.
- **Pre-Challenge Query Phase:** In this phase, $\mathcal{A}$ is allowed to make multiple queries id. For each id, the challenger returns $\mathsf{sk}_{\mathsf{id}} \leftarrow \mathsf{Sim}_2(\mathsf{st}_1, \mathsf{id})$ to $\mathcal{A}$.
- **Challenge Phase:** The adversary $\mathcal{A}$ sends id^* to the challenger where id^* was never queried in the pre-challenge query phase. The challenger computes $(\mathsf{ct}^*, \mathsf{st}_3) \leftarrow \mathsf{Sim}_3(\mathsf{st}_2, \mathsf{id}^*)$ and $\mathsf{msk} \leftarrow \mathsf{Sim}_4(\mathsf{st}_3, \mathsf{seskey})$ where seskey is randomly generated. It returns $(\mathsf{msk}, \mathsf{ct}^*, \mathsf{seskey})$ to $\mathcal{A}$.
- **Response Phase:** $\mathcal{A}$ outputs b.

It is important to note that in the challenge phase, the adversary obtains only the master secret key, the challenge ciphertext and the session key, and not the randomness used by the Setup algorithm or Encap algorithm to generate the challenge ciphertext and the session key. Let p_{real} and p_{sim} be the probabilities with which $\mathcal{A}$ outputs 0 in the real world and simulated world, respectively.

Definition 8. *An RNC-IB-KEM scheme is said to be secure if for all PPT adversaries $\mathcal{A}$, there exists a negligible function $\mathsf{negl}(\cdot)$ such that, for all $\lambda \in \mathbb{N}$,*

$$|p_{\text{real}} - p_{\text{sim}}| = \mathsf{negl}(\lambda)$$

By combining an RNC-IB-KEM with *one-time pad* encryption in a hybrid encryption approach, we can obtain an RNC-IBE scheme. This is possible because the RNC-IB-KEM (and RNC-IBE) requires the disclosure of the master secret key and not the randomness used by the Setup and/or Enc algorithms. However, the ciphertext-size will be the sum of the RNC-IB-KEM ciphertext-size and the size of the session-key.

Theorem 11. *Assuming the existence of secure RNC-IB-KEM, there exists secure RNC-IBE schemes.*

5 Receiver Non-Committing IB-KEM and IBE from Bilinear Groups

In this section, we present an adaptive secure receiver non-committing identity based key encapsulation mechanism (RNC-IB-KEM) using the concepts of dual system encryption [67].

5.1 Construction

Our construction is as follows. Let $\mathsf{HC} : \mathbb{G}_T \times \{0,1\}^{\log(p)} \to \{0,1\}$ denote a 1-bit randomness extractor over a group element and $\ell := \ell(\lambda)$ be a polynomial in λ.

$\mathsf{Gen}(1^\lambda)$**:**
- Generate a bilinear group $(\mathbb{G}_1, \mathbb{G}_2, \mathbb{G}_T, e, p, g_1, g_2)$.
- Generate $\mathbf{a}, \mathbf{b} \leftarrow \mathbb{Z}_p^2$ and $\mathbf{W}_1, \mathbf{W}_2 \leftarrow \mathbb{Z}_p^{2\times 2}$.
- Output $\mathsf{pp} := ([\mathbf{a}]_1, [\mathbf{b}]_2, [\mathbf{W}_1\mathbf{a}]_1, [\mathbf{W}_2\mathbf{a}]_1, [\mathbf{W}_1^\top\mathbf{b}]_2, [\mathbf{W}_2^\top\mathbf{b}]_2)$. We assume that $(\mathbb{G}_1, \mathbb{G}_2, \mathbb{G}_T, e, p, g_1, g_2)$ is included in pp and omit to write it.

$\mathsf{Setup}(\mathsf{pp})$**:**
- Parse $([\mathbf{a}]_1, [\mathbf{b}]_2, [\mathbf{W}_1\mathbf{a}]_1, [\mathbf{W}_2\mathbf{a}]_1, [\mathbf{W}_1^\top\mathbf{b}]_2, [\mathbf{W}_2^\top\mathbf{b}]_2) \leftarrow \mathsf{pp}$.
- Generate $h \leftarrow \{0,1\}^{\log(p)}$.
- Generate $\mathbf{k}_i \leftarrow \mathbb{Z}_p^2, \forall i \in [\ell]$.
- Output $\mathsf{mpk} := (\{\, [\mathbf{a}^\top\mathbf{k}_i]_T \,\}_{i\in[\ell]}, h)$ and $\mathsf{msk} := \{\, \mathbf{k}_i \,\}_{i\in[\ell]}$.

$\mathsf{Keygen}(\mathsf{pp}, \mathsf{msk}, \mathsf{id} \in \mathbb{Z}_p)$**:**
- Parse $([\mathbf{a}]_1, [\mathbf{b}]_2, [\mathbf{W}_1\mathbf{a}]_1, [\mathbf{W}_2\mathbf{a}]_1, [\mathbf{W}_1^\top\mathbf{b}]_2, [\mathbf{W}_2^\top\mathbf{b}]_2) \leftarrow \mathsf{pp}$ and $\{\, \mathbf{k}_i \,\} \leftarrow \mathsf{msk}$.
- Generate $s_i \leftarrow \mathbb{Z}_p$, $\forall i \in [\ell]$.
- Output $\mathsf{sk}_{\mathsf{id}} := \{\, ([s_i\mathbf{b}]_2, [\mathbf{k}_i + s_i(\mathbf{W}_1 + \mathsf{id}\cdot\mathbf{W}_2)^\top\mathbf{b}]_2) \,\}_{i\in[\ell]}$.

$\mathsf{Encap}(\mathsf{pp}, \mathsf{mpk}, \mathsf{id})$**:**
- Parse $([\mathbf{a}]_1, [\mathbf{b}]_2, [\mathbf{W}_1\mathbf{a}]_1, [\mathbf{W}_2\mathbf{a}]_1, [\mathbf{W}_1^\top\mathbf{b}]_2, [\mathbf{W}_2^\top\mathbf{b}]_2) \leftarrow \mathsf{pp}$ and $(\{\, [\mathbf{a}^\top\mathbf{k}_i]_T \,\}_{i\in[\ell]}, h) \leftarrow \mathsf{mpk}$.
- Generate $r \leftarrow \mathbb{Z}_p$.
- Output $\mathsf{ct} := ([r\mathbf{a}]_1, [r(\mathbf{W}_1 + \mathsf{id}\cdot\mathbf{W}_2)\mathbf{a}]_1)$ and $\mathsf{seskey} := \{\, \mathsf{HC}([r\mathbf{a}^\top\mathbf{k}_i]_T, h) \,\}_{i\in[\ell]}$.

$\mathsf{Decap}(\mathsf{pp}, \mathsf{sk}_{\mathsf{id}}, \mathsf{ct})$**:**
- Parse $\{\, ([\mathbf{d}_i]_2, [\mathbf{d}_i']_2) \,\}_{i\in[\ell]} \leftarrow \mathsf{sk}_{\mathsf{id}}$ and $([\mathbf{c}]_1, [\mathbf{c}']_1) \leftarrow \mathsf{ct}$.
- Output $\mathsf{seskey} := \{\, \mathsf{HC}(e([\mathbf{c}]_1^\top, [\mathbf{d}_i']_2)/e([\mathbf{c}']_1^\top, [\mathbf{d}_i]_2), h) \,\}_{i\in[\ell]}$.

$\mathsf{Sim}_1(1^\lambda, 1^p)$:
- Generate a bilinear group $(\mathbb{G}_1, \mathbb{G}_2, \mathbb{G}_T, e, p, g_1, g_2)$.
- Generate $\mathbf{a}, \mathbf{b} \leftarrow \mathbb{Z}_q$ and $\mathbf{W}_1, \mathbf{W}_2 \leftarrow \mathbb{Z}_p^{2\times 2}$.
- Generate $k_{i,1} \leftarrow \mathbb{Z}_p$, $\forall i \in [\ell]$ and $h \leftarrow \{0,1\}^{\log(p)}$.
- Output $\mathsf{pp} := ([\mathbf{a}]_1, [\mathbf{b}]_2, [\mathbf{W}_1\mathbf{a}]_1, [\mathbf{W}_2\mathbf{a}]_1, [\mathbf{W}_1^\top\mathbf{b}]_2, [\mathbf{W}_2^\top\mathbf{b}]_2)$ and $\mathsf{mpk} := (\{\, [k_{i,1}]_T \,\}_{i\in[\ell]}, h)$ and $\mathsf{st}_1 = (\{\, k_{i,1} \,\}_{i\in[\ell]}, \mathbf{a})$.

$\mathsf{Sim}_2(\mathsf{st}_1, \mathsf{id})$:
- Generate $s_i, w_i \leftarrow \mathbb{Z}_p$, $\forall i \in [\ell]$.
- Output $\mathsf{sk}_{\mathsf{id}} := \{\, ([s_i\mathbf{b}]_2, [\frac{k_1}{|\mathbf{a}|^2}\cdot\mathbf{a} + s_i(\mathbf{W}_1 + \mathsf{id}\cdot\mathbf{W}_2)^\top\mathbf{b} + w_i\mathbf{a}^\perp]_2) \,\}_{i\in[\ell]}$ and $\mathsf{st}_2 := \mathsf{st}_1$.

$\mathsf{Sim}_3(\mathsf{st}_2, \mathsf{id}^*)$:
- Generate $u_1, u_2 \leftarrow \mathbb{Z}_p$ and set $\mathbf{u} = u_1\mathbf{a} + u_2\mathbf{a}^\perp$.
- Outputs $\mathsf{ct}^* := ([\mathbf{u}]_1, [(\mathbf{W}_1 + \mathsf{id}^*\cdot\mathbf{W}_2)\mathbf{u}]_1)$ and $\mathsf{st}_3 := (\mathsf{st}_2, u_1, u_2)$.

$\mathsf{Sim}_4(\mathsf{st}_3, \mathsf{seskey} \in \{0,1\}^\ell)$:
- Generate $x_i \in \mathbb{Z}_p$ such that $\mathsf{seskey}_i = \mathsf{HC}([x_i]_T, h)$ via rejection sampling.
- Set $k_{i,2} := \frac{x_i - u_1 k_{i,1}}{u_2}$.

– Output $\mathsf{msk} := \{ \frac{k_{i,1}}{|\mathbf{a}|^2} \cdot \mathbf{a} + \frac{k_{i,2}}{|\mathbf{a}^\perp|^2} \cdot \mathbf{a}^\perp \}_{i \in [\ell]}$.

Remark 1. The randomness used in the setup algorithm includes h, which is part of the master public key and the set $\{ k_i \}$, which constitutes the entire master secret key. Since, the challenger outputs h in the initialization phase and provides the master secret key in the challenge, the adversary effectively gains access to all the randomness used in the setup algorithm during the challenge phase.

Parameters. The size of a ciphertext is poly(λ), i.e., it is independent of ℓ. Whereas, the size of the master public key, master secret key and secret keys depend on ℓ.

Correctness. For correctly generated $\mathsf{sk}_{\mathsf{id}} := (\{ [s_i \mathbf{b}]_2, [\mathbf{k}_i + s_i(\mathbf{W}_1 + \mathsf{id} \cdot \mathbf{W}_2)^\top \mathbf{b}]_2) \}_{i \in [\ell]}$ and $\mathsf{ct} := ([r\mathbf{a}]_1, [r(\mathbf{W}_1 + \mathsf{id} \cdot \mathbf{W}_2)\mathbf{a}]_1)$, we have

$$\begin{aligned} \frac{e([r\mathbf{a}]_1^\top, [\mathbf{k}_i + s_i(\mathbf{W}_1 + \mathsf{id} \cdot \mathbf{W}_2)^\top \mathbf{b}]_2)}{e([r(\mathbf{W}_1 + \mathsf{id} \cdot \mathbf{W}_2)\mathbf{a}]_1^\top, [s_i \mathbf{b}]_2)} &= \frac{[r\mathbf{a}^\top \mathbf{k}_i + r s_i \mathbf{a}^\top (\mathbf{W}_1 + \mathsf{id} \cdot \mathbf{W}_2)^\top \mathbf{b}]_T}{[r s_i \mathbf{a}^\top (\mathbf{W}_1 + \mathsf{id} \cdot \mathbf{W}_2)^\top \mathbf{b}]_T} \\ &= [r\mathbf{a}^\top \mathbf{k}_i]_T \end{aligned}$$

for all $i \in [\ell]$. Since, $\mathsf{HC}(\cdot, \cdot)$ is a deterministic function, the correctness follows immediately.

Theorem 12. *Assuming the hardness of* SXDH*, the above scheme is a secure receiver non-committing identity-based key encapsulation mechanism.*

Proof. Consider the following experiment for an adversary $\mathcal{A}$ that makes q queries to $O_{\mathsf{UserKeyGen}}$.

Hyb_0**:** This corresponds to the real experiment of non-committing security.

1. The challenger generates $\mathsf{pp}, \mathsf{mpk}, \mathsf{msk}$ as follows.
   – Generate a bilinear group $(\mathbb{G}_1, \mathbb{G}_2, \mathbb{G}_T, e, p, g_1, g_2)$.
   – Generate $\mathbf{a}, \mathbf{b} \leftarrow \mathbb{Z}_p^2$ and $\mathbf{W}_1, \mathbf{W}_2 \leftarrow \mathbb{Z}_p^{2 \times 2}$.
   – Generate $\mathbf{k}_i \leftarrow \mathbb{Z}_p^2$, $\forall i \in [\ell]$ and $h \leftarrow \{0,1\}^{\log(p)}$.
   – Set $\mathsf{pp} := ([\mathbf{a}]_1, [\mathbf{b}]_2, [\mathbf{W}_1\mathbf{a}]_1, [\mathbf{W}_2\mathbf{a}]_1, [\mathbf{W}_1^\top \mathbf{b}]_2, [\mathbf{W}_2^\top \mathbf{b}]_2)$, $\mathsf{mpk} := (\{ [\mathbf{a}^\top \mathbf{k}_i]_T \}_i, h)$, and $\mathsf{msk} := \{ \mathbf{k}_i \}_i$.

   The challenger sends pp and mpk to $\mathcal{A}$.
2. $\mathcal{A}$ can get access to the following oracle that provides access to $\mathsf{Keygen}(\mathsf{pp}, \mathsf{msk}, \cdot)$.
   $O_{\mathsf{UserKeyGen}}(\mathsf{id}^j)$**:** Given the j-query $\mathsf{id}^j \in \mathbb{Z}_p$ as an input, it returns $\mathsf{sk}_{\mathsf{id}^j}$ generated as follows.
   – Generate $s_i^j \leftarrow \mathbb{Z}_p$, $\forall i \in [\ell]$.
   – Set $\mathsf{sk}_{\mathsf{id}^j} := \{ ([s_i^j \mathbf{b}]_2, [\mathbf{k}_i + s_i^j (\mathbf{W}_1 + \mathsf{id}^j \cdot \mathbf{W}_2)^\top \mathbf{b}]_2) \}_i$.
3. $\mathcal{A}$ outputs $\mathsf{id}^* \in \mathbb{Z}_p$. The challenger generates $(\mathsf{ct}^*, \mathsf{seskey}^*)$ as follows.
   – Generate $r \leftarrow \mathbb{Z}_p$.
   – Set $\mathsf{ct}^* := ([r\mathbf{a}]_1, [r(\mathbf{W}_1 + \mathsf{id}^* \cdot \mathbf{W}_2)\mathbf{a}]_1)$ and $\mathsf{seskey}^* := \{ \mathsf{HC}([r\mathbf{a}^\top \mathbf{k}_i]_T, h) \}_i$.

   The challenger sends $(\mathsf{msk}, \mathsf{ct}^*, \mathsf{seskey}^*)$ to $\mathcal{A}$.
4. $\mathcal{A}$ outputs $\mathsf{coin}' \in \{0,1\}$.

Using a sequence of hybrid experiments, we will prove that the experiment can be indistinguishably changed into non-committing experiment. We will denote the probability that $\mathcal{A}$ outputs 0 in the hybrid Hyb_i using $p_{\mathcal{A},H_i}$.

Below, we assume that $\mathbf{a}$ and $\mathbf{b}^\perp$ are linearly independent and $\mathbf{a}^\perp$ and $\mathbf{b}$ are also linearly independent, which hold with overwhelming probability over the choice of $\mathbf{a}$ and $\mathbf{b}$.

Changing the challenge ciphertext into semi-functional mode.

Hybrid Hyb_1: This is the same as Hyb_0 except $(\mathsf{ct}^*, \mathsf{seskey}^*)$ is generated as $\mathsf{ct}^* := ([\mathbf{u}]_1, [(\mathbf{W}_1 + \mathsf{id}^* \cdot \mathbf{W}_2)\mathbf{u}]_1)$ and $\mathsf{seskey}^* := \{\, \mathsf{HC}([\mathbf{u}^\top \mathbf{k}_i]_T, h) \,\}_i$, where $\mathbf{u} \leftarrow \mathbb{Z}_p^2$.

We have $|p_{\mathcal{A},H_1} - p_{\mathcal{A},H_0}| = \mathsf{negl}(\lambda)$ from the DDH assumption on $\mathbb{G}_1$.

Lemma 1. *For all PPT adversaries $\mathcal{A}$, there exists a PPT adversary $\mathcal{B}$ such that, $|p_{\mathcal{A},H_1} - p_{\mathcal{A},H_0}| \leq p_{\mathcal{B},\mathsf{DDH}}$.*

We define $\mathsf{Hyb}_{1,0,5}$ as Hyb_1.

Changing the user secret keys into semi-functional mode. We change the user secret keys into semi-functional mode using $\mathsf{Hyb}_{1,i,1}, \cdots, \mathsf{Hyb}_{1,i,5}$ for $i \in [q]$. Below, we define $\mathbf{W}_0 := \mathbf{b}^\perp (\mathbf{a}^\perp)^\top / (\mathbf{b}^\perp)^\top \mathbf{a}^\perp$.

Hybrid $\mathsf{Hyb}_{1,i,1}$: This is the same as $\mathsf{Hyb}_{1,i-1,5}$ except that $O_{\mathsf{UserKeyGen}}$ behaves as follows.

$O_{\mathsf{UserKeyGen}}(\mathsf{id})$**:** Given the j-th query $\mathsf{id}^j \in \mathbb{Z}_p$ as an input, it behaves as follows.

- If $j < i$, return $\mathsf{sk}_{\mathsf{id}^j}$ generated as follows.
  - Generate $s_d^j, w_d^j \leftarrow \mathbb{Z}_p$, $\forall d \in [\ell]$.
  - Set $\mathsf{sk}_{\mathsf{id}^j} := \{\, ([s_d^j \mathbf{b}]_2, [\mathbf{k}_d + s_d^j (\mathbf{W}_1 + \mathsf{id}^j \cdot \mathbf{W}_2)^\top \mathbf{b} + w_d^j \mathbf{a}^\perp]_2) \,\}_d$.
- If $j = i$, return $\mathsf{sk}_{\mathsf{id}^i}$ generated as follows.
  - Generate $\mathbf{v}_d^i \leftarrow \mathbb{Z}_p^2$, $\forall d \in [\ell]$.
  - Set $\mathsf{sk}_{\mathsf{id}^i} := \{\, ([\mathbf{v}_d^i]_2, [\mathbf{k}_d + (\mathbf{W}_1 + \mathsf{id}^i \cdot \mathbf{W}_2)^\top \mathbf{v}_d^i]_2) \,\}_d$.
- If $j > i$, return $\mathsf{sk}_{\mathsf{id}^j}$ generated as follows.
  - Generate $s_d^j \leftarrow \mathbb{Z}_p$, $\forall d \in [\ell]$.
  - Set $\mathsf{sk}_{\mathsf{id}^j} := \{\, ([s_d^j \mathbf{b}]_2, [\mathbf{k}_d + s_d^j (\mathbf{W}_1 + \mathsf{id}^j \cdot \mathbf{W}_2)^\top \mathbf{b}]_2) \,\}_d$.

We have $|p_{\mathcal{A},H_{1,1,1}} - p_{\mathcal{A},H_1}| = \mathsf{negl}(\lambda)$ from the DDH assumption on $\mathbb{G}_2$.

Lemma 2. *For all PPT adversaries $\mathcal{A}$, there exists a PPT adversary $\mathcal{B}$ such that, $|p_{\mathcal{A},H_{1,1,1}} - p_{\mathcal{A},H_1}| \leq \ell \cdot p_{\mathcal{B},\mathsf{DDH}}$.*

Hybrid $\mathsf{Hyb}_{1,i,2}$: This is the same as $\mathsf{Hyb}_{1,i,1}$ except that we generate $\hat{\mathbf{W}}_1, \hat{\mathbf{W}}_2 \leftarrow \mathbb{Z}_p^{2\times 2}$ and $w_1, w_2 \leftarrow \mathbb{Z}_p$, and set $\mathbf{W}_1 := \hat{\mathbf{W}}_1 + w_1 \mathbf{W}_0$ and $\mathbf{W}_2 := \hat{\mathbf{W}}_2 + w_2 \mathbf{W}_0$.

We have $|p_{\mathcal{A},H_{1,i,2}} - p_{\mathcal{A},H_{1,i,1}}| = 0$ since the distribution of $\mathbf{W}_1, \mathbf{W}_2$ do not change.

Lemma 3. *For all PPT adversaries $\mathcal{A}$ and all $\lambda \in \mathbb{N}$, $|p_{\mathcal{A},H_{1,i,2}} - p_{\mathcal{A},H_{1,i,1}}| = 0$.*

We prove that w_1 and w_2 appears only in ct^* in the form of $w_1 + \mathsf{id}^* \cdot w_2$ and in $\mathsf{sk}_{\mathsf{id}^i}$ in the form of $w_1 + \mathsf{id}^i \cdot w_2$. First, we have $[\mathbf{W}_\alpha \mathbf{a}]_1 = [\hat{\mathbf{W}}_\alpha \mathbf{a}]_1$ and $[\mathbf{W}_\alpha^\top \mathbf{b}]_2 = [\hat{\mathbf{W}}_\alpha^\top \mathbf{b}]_2$ for $\alpha \in \{1,2\}$. For ct^*, we can write $\mathbf{u} = r\mathbf{a} + r'\mathbf{b}^\perp$, and thus we have

$$\begin{aligned}\mathsf{ct}^* &= ([\mathbf{u}]_1, [(\mathbf{W}_1 + \mathsf{id}^* \cdot \mathbf{W}_2)\mathbf{u}]_1) \\ &= ([\mathbf{u}]_1, [(\hat{\mathbf{W}}_1 + \mathsf{id}^* \cdot \hat{\mathbf{W}}_2)\mathbf{u} + r'(w_1 + \mathsf{id}^* \cdot w_2)\mathbf{b}^\perp]_1).\end{aligned}$$

Also, for $\mathsf{sk}_{\mathsf{id}^i}$, we can write $\mathbf{v}_d^i = s_d^i \mathbf{b} + t_d^i \mathbf{a}^\perp$ and thus we have

$$\begin{aligned}\mathsf{sk}_{\mathsf{id}^i} &= \{\, ([\mathbf{v}_d^i]_2, [\mathbf{k}_d + (\mathbf{W}_1 + \mathsf{id}^i \cdot \mathbf{W}_2)^\top \mathbf{v}_d^i]_2) \,\}_d \\ &= \{\, ([\mathbf{v}_d^i]_2, [\mathbf{k}_d + (\hat{\mathbf{W}}_2 + \mathsf{id}^i \cdot \hat{\mathbf{W}}_2)^\top \mathbf{v}_d^i + t_d^i(w_1 + \mathsf{id}^i \cdot w_2)\mathbf{a}^\perp]_2) \,\}_d .\end{aligned}$$

Moreover, for $j \neq i$, we can write $\mathsf{sk}_{\mathsf{id}^j}$ as

$$\mathsf{sk}_{\mathsf{id}^j} = \begin{cases} \{\, ([s_d^j \mathbf{b}]_2, [\mathbf{k}_d + s_d^j(\hat{\mathbf{W}}_1 + \mathsf{id}^j \cdot \hat{\mathbf{W}}_2)^\top \mathbf{b} + w_d^j \mathbf{a}^\perp]_2) \,\}_d & \text{for } \ j < i \\ \{\, ([s_d^j \mathbf{b}]_2, [\mathbf{k}_d + s_d^j(\hat{\mathbf{W}}_1 + \mathsf{id}^j \cdot \hat{\mathbf{W}}_2)^\top \mathbf{b}]_2) \,\}_d & \text{for } \ j > i \end{cases}$$

Hybrid $\mathsf{Hyb}_{1,i,3}$: This is the same as $\mathsf{Hyb}_{1,i,2}$ except that for the i-th query id^i, $O_{\mathsf{UserKeyGen}}$ returns $\mathsf{sk}_{\mathsf{id}^i} := \{\, ([\mathbf{v}_d^i]_2, [\mathbf{k}_d + (\mathbf{W}_1 + \mathsf{id}^i \cdot \mathbf{W}_2)^\top \mathbf{v}_d^i + w_d^i \mathbf{a}^\perp]_2) \,\}_d$, where $w_d^i \leftarrow \mathbb{Z}_p$.

It is important to note that only the secret key $\mathsf{sk}_{\mathsf{id}^i}$ contains information about w_1 and w_2, while the remaining secret keys do not. Therefore, we have $|p_{\mathcal{A},H_{1,i,3}} - p_{\mathcal{A},H_{1,i,2}}| = \mathsf{negl}(\lambda)$, since $\mathsf{id}^* \neq \mathsf{id}^i$, which implies

$$\{\, w_1 + \mathsf{id}^* \cdot w_2, w_1 + \mathsf{id}^i \cdot w_2 \,\}_{w_1,w_2} \equiv \{\, w_1 + \mathsf{id}^* \cdot w_2, u \,\}_{w_1,w_2,u}$$

and with probability $\frac{p-1}{p}$, a randomly chosen t_d^i is invertible.

Lemma 4. *For all PPT adversaries $\mathcal{A}$ and $\lambda \in \mathbb{N}$, $|p_{\mathcal{A},H_{1,i,3}} - p_{\mathcal{A},H_{1,i,2}}| = \mathsf{negl}(\lambda)$.*

Hybrid $\mathsf{Hyb}_{1,i,4}$: This is the same as $\mathsf{Hyb}_{1,i,3}$ except that we undo the change between $\mathsf{Hyb}_{1,i,1}$ and $\mathsf{Hyb}_{1,i,2}$. Namely, we generate $\mathbf{W}_1, \mathbf{W}_2 \leftarrow \mathbb{Z}_p^{2\times 2}$.

Lemma 5. *For all PPT adversaries $\mathcal{A}$, $|p_{\mathcal{A},H_{1,i,4}} - p_{\mathcal{A},H_{1,i,3}}| = 0$.*

Hybrid $\mathsf{Hyb}_{1,i,5}$: This is the same as $\mathsf{Hyb}_{1,i,4}$ except that for the i-th query id^i, $O_{\mathsf{UserKeyGen}}$ returns $\mathsf{sk}_{\mathsf{id}^i} := \{\, ([s_d^i \mathbf{b}]_2, [\mathbf{k}_d + s_d^i(\mathbf{W}_1 + \mathsf{id}^i \cdot \mathbf{W}_2)^\top \mathbf{b} + w_d^i \mathbf{a}^\perp]_2) \,\}_d$, where $s_d^i, w_d^i \leftarrow \mathbb{Z}_p$.

We have $|p_{\mathcal{A},H_{1,i,5}} - p_{\mathcal{A},H_{1,i,4}}| = \mathsf{negl}(\lambda)$ from the DDH assumption on $\mathbb{G}_2$.

Lemma 6. *For all PPT adversaries $\mathcal{A}$, there exists a PPT adversary $\mathcal{B}$ such that, $|p_{\mathcal{A},H_{1,i,5}} - p_{\mathcal{A},H_{1,i,4}}| \leq \ell \cdot p_{\mathcal{B},\mathsf{DDH}}$.*

We also have $|p_{\mathcal{A},H_{1,i+1,1}} - p_{\mathcal{A},H_{1,i,5}}| = \mathsf{negl}(\lambda)$ from DDH assumption on $\mathbb{G}_2$.

Lemma 7. *For all PPT adversaries $\mathcal{A}$, there exists a PPT adversary $\mathcal{B}$ such that, $|p_{\mathcal{A},H_{1,i+1,1}} - p_{\mathcal{A},H_{1,i,5}}| \leq \ell \cdot p_{\mathcal{B},\mathsf{DDH}}$.*

Final steps towards non-committing mode.

Hybrid Hyb_2: We define Hyb_2 as the same game as $\mathsf{Hyb}_{1,q,5}$. The detailed description is as follows.

1. The challenger generates $\mathsf{pp}, \mathsf{mpk}, \mathsf{msk}$ as follows.
   - Generate a bilinear group $(\mathbb{G}_1, \mathbb{G}_2, \mathbb{G}_T, e, p, g_1, g_2)$.
   - Generate $\mathbf{a}, \mathbf{b} \leftarrow \mathbb{Z}_q$ and $\mathbf{W}_1, \mathbf{W}_2 \leftarrow \mathbb{Z}_p^{2\times 2}$.
   - Generate $\mathbf{k}_i \leftarrow \mathbb{Z}_p^2$, $\forall i \in [\ell]$ and $h \leftarrow \{0,1\}^{\log(p)}$.
   - Set $\mathsf{pp} := ([\mathbf{a}]_1, [\mathbf{b}]_2, [\mathbf{W}_1\mathbf{a}]_1, [\mathbf{W}_2\mathbf{a}]_1, [\mathbf{W}_1^\top \mathbf{b}]_2, [\mathbf{W}_2^\top \mathbf{b}]_2)$ and $\mathsf{mpk} := (\{\, [\mathbf{a}^\top \mathbf{k}_i]_T \,\}_i, h)$ and $\mathsf{msk} := \{\, \mathbf{k}_i \,\}_i$.

   The challenger sends pp and mpk to $\mathcal{A}$.
2. $\mathcal{A}$ can get access to the following oracle.

   $O_{\mathsf{UserKeyGen}}(\mathsf{id}^j)$**:** Given the j-query $\mathsf{id}^j \in \mathbb{Z}_p$ as an input, it returns $\mathsf{sk}_{\mathsf{id}^j}$ generated as follows.
   - Generate $s_d^j, w_d^j \leftarrow \mathbb{Z}_p$, $\forall d \in [\ell]$.
   - Set $\mathsf{sk}_{\mathsf{id}^j} := \{\, ([s_d^j \mathbf{b}]_2, [\mathbf{k}_d + s_d^j(\mathbf{W}_1 + \mathsf{id}^j \cdot \mathbf{W}_2)^\top \mathbf{b} + w_d^j \mathbf{a}^\perp]_2) \,\}_d$.
3. $\mathcal{A}$ outputs $\mathsf{id}^* \in \mathbb{Z}_p$. The challenger generates $(\mathsf{ct}^*, \mathsf{seskey}^*)$ as follows.
   - Generate $\mathbf{u} \leftarrow \mathbb{Z}^2$.
   - Set $\mathsf{ct}^* := ([\mathbf{u}]_1, [(\mathbf{W}_1 + \mathsf{id}^* \cdot \mathbf{W}_2)\mathbf{u}]_1)$ and $\mathsf{seskey}^* := \{\, \mathsf{HC}([\mathbf{u}^\top \mathbf{k}_i]_T, h) \,\}_i$.

   The challenger sends $(\mathsf{msk}, \mathsf{ct}^*, \mathsf{seskey}^*)$ to $\mathcal{A}$.
4. $\mathcal{A}$ outputs $\mathsf{coin}' \in \{0,1\}$.

Hybrid Hyb_3: This is the same as Hyb_2 except that we generate $k_{i,1}, k_{i,2} \leftarrow \mathbb{Z}_p$ and set $\mathbf{k}_i \leftarrow \frac{k_{i,1}}{|\mathbf{a}|^2} \cdot \mathbf{a} + \frac{k_{i,2}}{|\mathbf{a}^\perp|^2} \cdot \mathbf{a}^\perp$. By this change, we have $\mathsf{mpk} = \{\, [k_{i,1}]_T \,\}_i$.

This is just conceptual change and we have $|p_{\mathcal{A},H_3} - p_{\mathcal{A},H_2}| = 0$.

Lemma 8. *For all PPT adversaries $\mathcal{A}$ and all $\lambda \in \mathbb{N}$, $|p_{\mathcal{A},H_3} - p_{\mathcal{A},H_2}| = 0$.*

Hybrid Hyb_4: This is the same as Hyb_3 except that for the j-th query id^j, $O_{\mathsf{UserKeyGen}}$ returns $\mathsf{sk}_{\mathsf{id}^j} := \{\, ([s_d^j \mathbf{b}]_2, [\frac{k_1}{|\mathbf{a}|^2} \cdot \mathbf{a} + s_d^j(\mathbf{W}_1 + \mathsf{id}^j \cdot \mathbf{W}_2)^\top \mathbf{b} + w_d^j \mathbf{a}^\perp]_2) \,\}_d$, where $s_d^j, w_d^j \leftarrow \mathbb{Z}_p$.

We have $|p_{\mathcal{A},H_4} - p_{\mathcal{A},3}| = 0$ since $\frac{k_2}{|\mathbf{a}^\perp|^2} + w_d^j$ and w_d^j identically distributes for every $j \in [q], d \in [\ell]$ when w_d^j is chosen uniformly at random.

Lemma 9. *For all PPT adversaries $\mathcal{A}$ and for all $\lambda \in \mathbb{N}$, $|p_{\mathcal{A},H_4} - p_{\mathcal{A},3}| = 0$.*

Hybrid Hyb_5: This is the same as Hyb_4 except that we generate $u_1, u_2 \leftarrow \mathbb{Z}_p$ and set $\mathbf{u} \leftarrow u_1 \mathbf{a} + u_2 \mathbf{a}^\perp$.

This is just conceptual change and we have $|p_{\mathcal{A},H_5} - p_{\mathcal{A},4}| = 0$.

Lemma 10. *For all PPT adversaries $\mathcal{A}$ and for all $\lambda \in \mathbb{N}$, $|p_{\mathcal{A},H_5} - p_{\mathcal{A},4}| = 0$.*

Hybrid Hyb_6: This is the same as Hyb_5 except that we generate $x_i \leftarrow \mathbb{Z}_p$ and we set $k_{i,2} = \frac{x_i - u_1 k_{i,1}}{u_2}$. By this change, we have $\mathsf{seskey}^* = \{\, \mathsf{HC}([x_i]_T, h) \,\}_i$.

We have $|p_{\mathcal{A},H_6} - p_{\mathcal{A},5}| \leq \mathsf{negl}(\lambda)$ since k_2 still distributes uniformly at random when u_2 is invertible which occurs with high probability.

Lemma 11. *For all PPT adversaries $\mathcal{A}$ and $\lambda \in \mathbb{N}$, $|p_{\mathcal{A},H_6} - p_{\mathcal{A},5}| \leq \mathsf{negl}(\lambda)$.*

Hybrid Hyb_7: This is the same as Hyb_6 except that we defer the generation of k_2 until the challenge phase. The detailed description is as follows.

1. The challenger generates $\mathsf{pp}, \mathsf{mpk}, \mathsf{msk}$ as follows.
   - Generate a bilinear group $(\mathbb{G}_1, \mathbb{G}_2, \mathbb{G}_T, e, p, g_1, g_2)$.
   - Generate $\mathbf{a}, \mathbf{b} \leftarrow \mathbb{Z}_q$ and $\mathbf{W}_1, \mathbf{W}_2 \leftarrow \mathbb{Z}_p^{2\times 2}$.
   - Generate $k_{i,1} \leftarrow \mathbb{Z}_p, \forall i \in [\ell]$ and $h \leftarrow \{0,1\}^{\log(p)}$.
   - Set $\mathsf{pp} := ([\mathbf{a}]_1, [\mathbf{b}]_2, [\mathbf{W}_1\mathbf{a}]_1, [\mathbf{W}_2\mathbf{a}]_1, [\mathbf{W}_1^\top \mathbf{b}]_2, [\mathbf{W}_2^\top \mathbf{b}]_2)$ and $\mathsf{mpk} := (\{\, [k_{i,1}]_T \,\}_i, h)$.

   The challenger sends pp and mpk to $\mathcal{A}$.
2. $\mathcal{A}$ can get access to the following oracle.

   $O_{\mathsf{UserKeyGen}}(\mathsf{id}^j)$**:** Given the j-query $\mathsf{id}^j \in \mathbb{Z}_p$ as an input, it returns $\mathsf{sk}_{\mathsf{id}^j}$ generated as follows.
   - Generate $s_d^j, w_d^j \leftarrow \mathbb{Z}_p, \forall d \in [\ell]$.
   - Set $\mathsf{sk}_{\mathsf{id}^j} := \{\, ([s_d^j \mathbf{b}]_2, [\frac{k_1}{|\mathbf{a}|^2} \cdot \mathbf{a} + s_d^j (\mathbf{W}_1 + \mathsf{id}^j \cdot \mathbf{W}_2)^\top \mathbf{b} + w_d^j \mathbf{a}^\perp]_2) \,\}_d$.
3. $\mathcal{A}$ outputs $\mathsf{id}^* \in \mathbb{Z}_p$. The challenger generates $(\mathsf{ct}^*, \mathsf{seskey}^*)$ as follows.
   - Generate $u_1, u_2 \leftarrow \mathbb{Z}_p$ and set $\mathbf{u} = u_1\mathbf{a} + u_2\mathbf{a}^\perp$.
   - Generate $x_i \leftarrow \mathbb{Z}_p, \forall i \in [\ell]$
   - Set $\mathsf{ct}^* := ([\mathbf{u}]_1, [(\mathbf{W}_1 + \mathsf{id}^* \cdot \mathbf{W}_2)\mathbf{u}]_1)$ and $\mathsf{seskey}^* := \{\, \mathsf{HC}([x_i]_T, h) \,\}_i$.
   - Set $k_{i,2} = \frac{x_i - u_1 k_{i,1}}{u_2}$ and $\mathsf{msk} := \{\, \frac{k_{i,1}}{|\mathbf{a}|^2} \cdot \mathbf{a} + \frac{k_{i,2}}{|\mathbf{a}^\perp|^2} \cdot \mathbf{a}^\perp \,\}_i$.

   The challenger sends $(\mathsf{msk}, \mathsf{ct}^*, \mathsf{seskey}^*)$ to $\mathcal{A}$.
4. $\mathcal{A}$ outputs $\mathsf{coin}' \in \{0,1\}$.

It is easy to see that Hyb_7 can be simulated using the RNC-IB-KEM simulators. Using the above lemma along with triangular inequality, the theorem follows. □

Combining Theorem 11 and Theorem 12, we obtain the following.

Theorem 13. *Assuming hardness of* SXDH*, there exists secure RNC-IBE schemes.*

6 Receiver Non-Committing IBE from Batch Encryption

In this section, we construct an adaptive secure RNC-IBE scheme which supports polynomially many identities. Let

- $\mathsf{BE} = (\mathsf{BE.Params}, \mathsf{BE.Setup}, \mathsf{BE.Enc}, \mathsf{BE.Dec})$ be oblivious batch encryption.
- $\mathsf{GC} = (\mathsf{GC.Grbl}, \mathsf{GC.Eval}, \mathsf{GC.Sim})$ be a garbling scheme.
- $\mathsf{NCE} = (\mathsf{NCE.Setup}, \mathsf{NCE.Enc}, \mathsf{NCE.Dec})$ be a NCE scheme.

Let d be the length of the identities such that $T = 2^d = \mathrm{poly}(\lambda)$ and n be the length of the public keys of NCE.

$\mathsf{Setup}(1^\lambda, 1^T)$:

- Generate $\mathsf{be.pp} \leftarrow \mathsf{BE.Params}(1^\lambda, 1^{nT})$.
- Generate $(\mathsf{nce.pk}_i, \mathsf{nce.sk}_i) \leftarrow \mathsf{NCE.Setup}(1^\lambda; r^{(i)}_{\mathsf{NCE.Setup}})$ for $i \in \{0,1\}^d$.
- Generate $\mathsf{be.pk} = \mathsf{BE.Setup}(\mathsf{be.pp}, \{\mathsf{nce.pk}_i\}_{i\in\{0,1\}^d})$.
- Output $\mathsf{msk} := \{\mathsf{nce.pk}_i, \mathsf{nce.sk}_i\}_{i\in\{0,1\}^d}$ and $\mathsf{mpk} := (\mathsf{be.pp}, \mathsf{be.pk})$.

$\mathsf{Keygen}(\mathsf{msk}, \mathsf{id} \in \{0,1\}^d)$:

- Output $\mathsf{sk_{id}} := (\{\mathsf{nce.pk}_i\}_{i\in\{0,1\}^d}, \mathsf{nce.sk_{id}})$.

$\mathsf{Enc}(\mathsf{mpk}, \mathsf{id}, m)$:

- Generate r uniformly at random.
- Generate $(\tilde{\mathcal{C}}^{(\mathsf{id})}, \{\mathsf{lab}^{(\mathsf{id})}_{i,b}\}) \leftarrow \mathsf{GC.Grbl}(1^\lambda, \mathsf{NCE.Enc}(\cdot, m; r))$.
- Generate $(\tilde{\mathcal{C}}^{(k)}, \{\mathsf{lab}^{(k)}_{i,b}\}) \leftarrow \mathsf{GC.Grbl}(1^\lambda, \mathsf{NCE.Enc}(\cdot, m^k; r^k))$ where m^k, r^k are randomly generated for all $k \in \{0,1\}^d \setminus \{\mathsf{id}\}$.
- Generate a matrix $M \in [\{0,1\}^\lambda]^{nT\times 2}$ such that $M[int(k)\cdot n + j, b] = \mathsf{lab}^{(k)}_{j,b}$ for all $k \in \{0,1\}^d, b \in \{0,1\}, j \in [n]$ where $int : \{0,1\}^d \to [T]_0$ is a lexicographical mapping from the set of binary string to integers.
- Compute $\mathsf{be.ct} \leftarrow \mathsf{BE.Enc}(\mathsf{mpk}, M)$.
- Output $\mathsf{ct} := (\{\tilde{\mathcal{C}}^{(k)}\}_{k\in\{0,1\}^d}, \mathsf{be.ct})$.

$\mathsf{Dec}(\mathsf{sk_{id}}, \mathsf{ct})$:

- Compute $d \leftarrow \mathsf{BE.Dec}(\mathsf{be.pp}, \{\mathsf{nce.pk}_i\}_{i\in\{0,1\}^d}, \mathsf{be.ct})$.
- Set $\mathsf{lab}_i := d[int(\mathsf{id})\cdot n + i]$ for all $i \in [n]$.
- Compute $\mathsf{nce.ct_{id}} \leftarrow \mathsf{GC.Eval}(\tilde{\mathcal{C}}^{(\mathsf{id})}, \{\mathsf{lab}_i\}_{i\in[n]})$.
- Output $m \leftarrow \mathsf{NCE.Dec}(\mathsf{nce.sk_{id}}, \mathsf{nce.ct_{id}})$.

$\mathsf{Sim}_1(1^\lambda, 1^T)$:

- Compute $\mathsf{be.pp} \leftarrow \mathsf{BE.Params}(1^\lambda, 1^{nT})$
- Generate $(\mathsf{nce.pk}_i, \mathsf{nce.ct}_i, \mathsf{nce.st}_i) \leftarrow \mathsf{NCE.Sim}_1(1^\lambda)$ for $i \in \{0,1\}^d$.
- Compute $\mathsf{be.pk} := \mathsf{BE.Setup}(\mathsf{be.pp}, \{\mathsf{nce.pk}_i\}_{i\in\{0,1\}^d})$.
- Set $\mathsf{mpk} := (\mathsf{be.pp}, \mathsf{be.pk})$.
- Generate $(\tilde{\mathcal{C}}^{(k)}, \{\mathsf{lab}^{(k)}_i\}) \leftarrow \mathsf{GC.Sim}(1^\lambda, \mathsf{nce.ct}_k)$ for all $k \in \{0,1\}^d$.
- Generate a matrix $M \in [\{0,1\}^\lambda]^{nT\times 2}$ and sets $M[int(k)\cdot n + j, b] = \mathsf{lab}^{(k)}_j$ for all $k \in \{0,1\}^d, j \in [n]$.
- Compute $\mathsf{be.ct} \leftarrow \mathsf{BE.Enc}(\mathsf{mpk}, M)$.
- Set $\mathsf{ct} := (\{\tilde{\mathcal{C}}^{(k)}\}_{k\in\{0,1\}^d}, \mathsf{be.ct})$ and $\mathsf{st}^{(1)} := \{\mathsf{be.pp}, \{\mathsf{nce.ct}_i, \mathsf{nce.pk}_i, \mathsf{nce.st}_i\}_i\}$.
- Output $(\mathsf{mpk}, \mathsf{ct}, \mathsf{st}^{(1)})$.

$\mathsf{Sim}_2(\mathsf{st}^{(1)}, \mathsf{id})$:

- If $(\mathsf{id}, \cdot) \notin \mathsf{st}^{(2)}$, it computes $(\cdot, r^{\mathsf{id}}_{\mathsf{NCE.Setup}}) \leftarrow \mathsf{NCE.Sim}_2(\mathsf{nce.st_{id}}, m)$ where m is randomly generated and updates $\mathsf{st}^{(2)} = \mathsf{st}^{(2)} \cup \{\mathsf{id}, r^{\mathsf{id}}_{\mathsf{NCE.Setup}}\}$.
- Compute $(\cdot, \mathsf{nce.sk_{id}}) \leftarrow \mathsf{NCE.Setup}(1^\lambda; r^{(\mathsf{id})}_{\mathsf{NCE.Setup}})$.
- Output $\mathsf{sk_{id}} := (\{\mathsf{nce.pk}_i\}_{i\in\{0,1\}^d}, \mathsf{nce.sk_{id}})$.

$\mathsf{Sim}_3(\mathsf{st}^{(1)}, \mathsf{st}^{(2)}, \mathsf{id}, m)$:

- Compute $(\cdot, r^{\mathsf{id}}_{\mathsf{NCE.Setup}}) \leftarrow \mathsf{NCE.Sim}_2(\mathsf{nce.st_{id}}, m)$.
- For all $\mathsf{id}' \notin \{\mathsf{st}^{(2)}[0] \cup \{\mathsf{id}\}\}$, it computes $(\cdot, r^{\mathsf{id}'}_{\mathsf{NCE.Setup}}) \leftarrow \mathsf{NCE.Sim}_2(\mathsf{nce.st_{id'}}, m_{\mathsf{id}'})$ where $m_{\mathsf{id}'}$'s are randomly generated.
- Output $r_{\mathsf{Setup}} := \{\mathsf{be.pp}, \{r^i_{\mathsf{NCE.Setup}}\}_{i\in\{0,1\}^d}\}$.

Remark 2. From the oblivious property of BE, the randomness used in the setup algorithm is $r_{\mathsf{Setup}} = \{\, \mathsf{be.pp}, \{\, r^{(i)}_{\mathsf{NCE.Setup}} \,\}_{i\in\{\,0,1\,\}^d} \,\}$

Parameters. The size of a ciphertext is $\mathrm{poly}(2^d, |m|, \lambda)$, whereas the size of the master public key, master secret key and secret keys are of the form $\mathrm{poly}(2^d, \lambda)$.
Correctness. For a correctly generated secret key $\mathsf{sk_{id}} := (\{\, \mathsf{nce.pk}_i \,\}_{i\in\{\,0,1\,\}^d}, \mathsf{nce.sk_{id}})$ and a ciphertext $\mathsf{ct} := (\{\, \tilde{\mathcal{C}}^{(k)} \,\}_{k\in\{\,0,1\,\}^d}, \mathsf{be.ct})$, after performing the BE decryption $d \leftarrow \mathsf{BE.Dec}(\mathsf{be.pp}, \{\, \mathsf{nce.pk}_i \,\}_{i\in\{\,0,1\,\}^d}, \mathsf{be.ct})$, we have $d[int(\mathsf{id})\cdot n + i] = \mathsf{lab}_{i,\mathsf{nce.pk_{id}}[i]}$. This is due to the correctness of the BE scheme. From the correctness of GC scheme, we have $\mathsf{nce.ct_{id}} \leftarrow \mathsf{GC.Eval}(\tilde{\mathcal{C}}^{(\mathsf{id})}, \{\, \mathsf{lab}_i \,\}_{i\in[n]})$ where $\mathsf{nce.ct_{id}} = \mathsf{NCE.Enc}(\mathsf{nce.pk_{id}}, m^*)$. Finally, by the decryption correctness of the NCE, we have $m \leftarrow \mathsf{NCE.Dec}(\mathsf{nce.sk_{id}}, \mathsf{nce.ct_{id}})$.

Theorem 14. *Assuming* BE *is a secure oblivious batch encryption,* GC *is a secure garbling scheme and* NCE *is a secure non-committing encryption scheme, the above scheme is an adaptively secure RNC-IBE that support polynomially many identities.*

Proof. We will show that the above scheme is an adaptive secure RNC-IBE using a sequence of hybrid arguments.

Hybrid Hyb_0 : This is the original adaptive NC-IBE game.

1. The challenger generates $\mathsf{mpk}, \mathsf{msk}$ as follows.
   - Generate $\mathsf{be.pp} \leftarrow \mathsf{BE.Params}(1^\lambda, 1^{nT})$.
   - Generate $(\mathsf{nce.pk}_i, \mathsf{nce.sk}_i) \leftarrow \mathsf{NCE.Setup}(1^\lambda; r^{(i)}_{\mathsf{NCE.Setup}})$ for $i \in \{\, 0,1 \,\}^d$.
   - Generate $\mathsf{be.pk} = \mathsf{BE.Setup}(\mathsf{be.pp}, \{\mathsf{nce.pk}_i\}_{i\in\{\,0,1\,\}^d})$.
   - Set $\mathsf{msk} := \{\, \mathsf{nce.pk}_i, \mathsf{nce.sk}_i \,\}_{i\in\{\,0,1\,\}^d}$ and $\mathsf{mpk} := (\mathsf{be.pp}, \mathsf{be.pk})$.

   The challenger sends mpk to $\mathcal{A}$.
2. $\mathcal{A}$ can get access to the following oracle.
   $O_{\mathsf{UserKeyGen}}(\mathsf{id}^j)$**:** Given the j-query id^j as an input, it returns $\mathsf{sk}_{\mathsf{id}^j} = (\{\, \mathsf{nce.pk}_i \,\}_{i\in\{\,0,1\,\}^d}, \mathsf{nce.sk}_{\mathsf{id}^j})$.
3. $\mathcal{A}$ outputs id^*, m^*. The challenger generates ct^* as follows.
   - Generate r uniformly at random.
   - Generate $(\tilde{\mathcal{C}}^{(\mathsf{id})}, \{\, \mathsf{lab}^{(\mathsf{id})}_{i,b} \,\}) \leftarrow \mathsf{GC.Grbl}(1^\lambda, \mathsf{NCE.Enc}(\cdot, m; r))$.
   - Generate $(\tilde{\mathcal{C}}^{(k)}, \{\, \mathsf{lab}^{(k)}_{i,b} \,\}) \leftarrow \mathsf{GC.Grbl}(1^\lambda, \mathsf{NCE.Enc}(\cdot, m^k; r^k))$ where m^k, r^k are randomly generated for all $k \in \{\, 0,1 \,\}^d \setminus \{\, \mathsf{id} \,\}$.
   - Generate a matrix $M \in [\{\, 0,1 \,\}^\lambda]^{nT\times 2}$ such that $M[int(k)\cdot n + j, b] = \mathsf{lab}^{(k)}_{j,b}$ for all $k \in \{\, 0,1 \,\}^d, b \in \{\, 0,1 \,\}, j \in [n]$.
   - Set $\mathsf{ct} := (\{\, \tilde{\mathcal{C}}^{(k)} \,\}_{k\in\{\,0,1\,\}^d}, \mathsf{be.ct})$.

   The challenger sends $(\mathsf{ct}^*, \{\, r^{(i)}_{\mathsf{NCE.Setup}} \,\})$ to $\mathcal{A}$.
4. $\mathcal{A}$ outputs $\mathsf{coin}' \in \{0,1\}$.

Hybrid Hyb_1 : In this game, the matrix M during the challenge phase is computed as $M[int(\mathsf{id})\cdot n + j, b] = \mathsf{lab}^{(\mathsf{id})}_{j,\mathsf{nce.pk_{id}}[j]}$ for all $j \in [n-1], b \in \{\, 0,1 \,\}, \mathsf{id} \in \{\, 0,1 \,\}^d$.

Lemma 12. *Assume that* BE *is a secure batch encryption scheme, then for all* PPT *adversaries* $\mathcal{A}$ *there exists a negligible function* $\mathsf{negl}(\cdot)$ *such that for all* $\lambda \in \mathbb{N},$, $|p_{\mathcal{A},\mathsf{Hyb}_1} - p_{\mathcal{A},\mathsf{Hyb}_0}| \leq \mathsf{negl}(\lambda)$.

Proof. This follows from the security of batch encryption. The reduction can generate $\{ \mathsf{nce.pk}_i \}$ and sends it to the challenger who replies with $\mathsf{be.pp}$ (which is also the randomness used during $\mathsf{BE.Params}$). It can use $\mathsf{be.pp}$ to generates the master public key mpk. The reduction can generate the two matrices used in Hyb_0 and Hyb_1 and sends it to the challenger. Note that these matrices differ only at indices $(int(k) \cdot n + j, 1 - \mathsf{nce.pk}_k[j])$. It will receive $\mathsf{be.ct}$ from the challenger and then simulate the entire game using these values. Observe that the reduction generates all the keys $(\mathsf{nce.pk}_{\mathsf{id}}, \mathsf{nce.sk}_{\mathsf{id}})$, so it has $r^{\mathsf{id}}_{\mathsf{NCE.Setup}}$ and $\mathsf{be.pp}$. □

Hybrid Hyb_2 : In this game, the garbled circuit $\tilde{\mathcal{C}}^{(k)}$ and labels are simulated, i.e., $(\tilde{\mathcal{C}}^{(k)}, \{ \mathsf{lab}_i^{(k)} \}) \leftarrow \mathsf{GC.Sim}(1^\lambda, \mathsf{nce.ct}^{(k)})$ where $\mathsf{nce.ct}^{(k)} \leftarrow \mathsf{NCE.Enc}_1(\mathsf{nce.pk}_k, m^{(k)}; r^{(k)})$ for all $k \in \{ 0,1 \}^d \backslash \{ \mathsf{id}^* \}$. And, $\mathsf{nce.ct}^{(\mathsf{id}^*)} \leftarrow \mathsf{NCE.Enc}(\mathsf{nce.pk}_{\mathsf{id}^*}, m^*)$.

Lemma 13. *Assume that* GC *is a secure garbling scheme, then for all* PPT *adversaries* $\mathcal{A}$ *there exists a negligible function* $\mathsf{negl}(\cdot)$ *such that for all* $\lambda \in \mathbb{N},$, $|p_{\mathcal{A},\mathsf{Hyb}_2} - p_{\mathcal{A},\mathsf{Hyb}_1}| \leq \mathsf{negl}(\lambda)$.

Proof. This directly follows from the security of GC because the matrix M requires only $\mathsf{lab}^{(\mathsf{id})}_{i,\mathsf{nce.pk}_{\mathsf{id}}[i]}$ labels for all $\mathsf{id} \in \{ 0,1 \}$. □

Hybrid Hyb_3 : In this game, $\mathsf{nce.ct}^{(\mathsf{id})}$ and $\mathsf{nce.sk}_{\mathsf{id}}$ are all simulated using the NCE simulators.

Lemma 14. *Assume that* NCE *is a secure NCE scheme, then for all* PPT *adversaries* $\mathcal{A}$ *there exists a negligible function* $\mathsf{negl}(\cdot)$ *such that for all* $\lambda \in \mathbb{N},$, $|p_{\mathcal{A},\mathsf{Hyb}_3} - p_{\mathcal{A},\mathsf{Hyb}_2}| \leq \mathsf{negl}(\lambda)$.

Proof. We will show that an adversary $\mathcal{A}$ that can distinguish Hyb_3 from Hyb_2 can be transformed into an adversary $\mathcal{B}$ that breaks the NCE security. We achieve this by considering $T+1$ many intermediate hybrids $\mathsf{Hyb}_{2,i}$ where $\mathsf{Hyb}_{2,0} = \mathsf{Hyb}_2$ and $\mathsf{Hyb}_{2,T+1} = \mathsf{Hyb}_3$. The description of $\mathsf{Hyb}_{2,i}$ is as follows.

- The challenger generates the first i public keys using the first NCE simulator. Recall that the simulator will also generate fake ciphertexts. The remaining public-secret keys are generated honestly.
- To respond to a key query, if $\mathsf{id} > i$, it returns $\mathsf{sk}_{\mathsf{id}} = (\{ \mathsf{nce.pk}_i \}_{i \in \{ 0,1 \}^d}, \mathsf{nce.sk}_{\mathsf{id}})$. Whereas, if $\mathsf{id} \leq i$, it first checks whether $(\mathsf{id}, r^{\mathsf{id}}_{\mathsf{NCE.Setup}}) \in \mathsf{st}^{(2)}$. If it exists, then it computes $(\cdot, \mathsf{nce.sk}_{\mathsf{id}}) \leftarrow \mathsf{NCE.Setup}(1^\lambda; r^{\mathsf{id}}_{\mathsf{Setup}})$ and then returns $\mathsf{sk}_{\mathsf{id}}$ accordingly. If not, it calls the second simulator of the NCE on a random message m which returns $(\cdot, r^{(\mathsf{id})}_{\mathsf{Setup}})$. It computes $(\cdot, \mathsf{nce.sk}_{\mathsf{id}}) \leftarrow \mathsf{NCE.Setup}(1^\lambda; r^{\mathsf{id}}_{\mathsf{Setup}})$ and then returns $\mathsf{sk}_{\mathsf{id}}$ accordingly. It also updates $\mathsf{st}^{(2)} = \mathsf{st}^{(2)} \cup (\mathsf{id}, r^{(\mathsf{id})}_{\mathsf{Setup}})$.

– In the challenge phase, when it receives the challenger identity id^* and message m^*, it will generate $\mathsf{nce.ct}^{(k)}$ for all $k > i$ honestly, i.e., using NCE.Enc algorithm. The remaining $\mathsf{nce.ct}^{(k)}$ for $k \leq i$ will be the ciphertext simulated by the first simulator in the initialization phase. Then, it produces to generate $(\tilde{\mathcal{C}}^{(k)}, \{\, \mathsf{lab}_i^{(k)} \,\})$ for all k and sets up M appropriate to generate be.ct. Finally,
  - if $\mathsf{id}^* \leq i$, it will use the NCE simulator on m^* to obtain $r^{\mathsf{id}^*}_{\mathsf{Setup}}$.
  - For all $\mathsf{id} \leq i$ that was not queried and not equal to id^*, it will use the NCE simulator on a random m to obtain $r^{\mathsf{id}}_{\mathsf{Setup}}$.

  Finally, it will output $(\tilde{\mathcal{C}}^{(k)}, \mathsf{be.ct})$ and $\{\, r^{\mathsf{id}}_{\mathsf{Setup}} \,\}_{\mathsf{id}}$.

It is easy to show that an $\mathcal{A}$ that can distinguish $\mathsf{Hyb}_{2,i}$ from $\mathsf{Hyb}_{2,i+1}$ can be transformed into an adversary $\mathcal{B}$ that breaks the NCE security. □

Using the above lemmas and triangular inequality, for all PPT adversaries $\mathcal{A}$, there exists a negligible function $\mathsf{negl}(\cdot)$ such that for all $\lambda \in \mathbb{N}$, $|p_{\mathcal{A},\mathsf{Hyb}_0} - p_{\mathcal{A},\mathsf{Hyb}_3}| \leq \mathsf{negl}(\lambda)$. □

Using Theorem 8 and Theorem 7, we have Theorem 3.

7 Rate-1 Strongly Incompressible IBE from RNC-IB-KEM

In this section, we show a construction for strongly incompressible IBE using an RNC-IB-KEM and an incompressible SKE scheme. Let $\mathsf{IBKEM} = (\mathsf{IBKEM.Setup}, \mathsf{IBKEM.KeyGen}, \mathsf{IBKEM.Encap}, \mathsf{IBKEM.Decap})$ be an RNC-IB-KEM and $\mathsf{IncSKE} = (\mathsf{IncSKE.Setup}, \mathsf{IncSKE.Enc}, \mathsf{IncSKE.Dec})$ be an incompressible SKE scheme such that IncSKE.Setup outputs a truly random string.

$\mathsf{Setup}(1^\lambda, 1^S)$: The setup algorithm takes as input the security parameter λ and the upper bound for the state bound S. It computes $(\mathsf{ibkem.mpk}, \mathsf{ibkem.msk}) \leftarrow \mathsf{IBKEM.Setup}(1^\lambda, 1^{|\mathsf{inc.sk}|})$ and outputs $\mathsf{mpk} := \mathsf{ibkem.mpk}$ and $\mathsf{msk} := \mathsf{ibkem.msk}$.

$\mathsf{KeyGen}(\mathsf{msk}, \mathsf{id})$: The key generation algorithm takes as input a master secret key $\mathsf{msk} = \mathsf{ibkem.msk}$ and an identity id and computes $\mathsf{ibkem.sk}_{\mathsf{id}} \leftarrow \mathsf{IBKEM.KeyGen}(\mathsf{ibkem.msk}, \mathsf{id})$. It outputs $\mathsf{ibkem.sk}_{\mathsf{id}}$.

$\mathsf{Enc}(\mathsf{mpk}, \mathsf{id}, m)$: The encryption algorithm takes as input a master public key $\mathsf{mpk} = \mathsf{ibkem.mpk}$, an identity id and message m. It generates $(\mathsf{ibkem.ct}, \mathsf{seskey}) \leftarrow \mathsf{IBKEM.Encap}(\mathsf{ibkem.mpk}, \mathsf{id})$ and sets $\mathsf{inc.sk} := \mathsf{seskey}$. It then computes $\mathsf{inc.ct} \leftarrow \mathsf{IncSKE.Enc}(\mathsf{inc.sk}, m)$ and returns $\mathsf{ct} := (\mathsf{ibkem.ct}, \mathsf{inc.ct})$.

$\mathsf{Dec}(\mathsf{sk}, \mathsf{ct})$: The decryption algorithm takes as input a secret key $\mathsf{sk} = \mathsf{ibkem.sk}$ and a ciphertext $\mathsf{ct} = (\mathsf{ibkem.ct}, \mathsf{inc.ct})$. It first computes $\mathsf{inc.sk} \leftarrow \mathsf{IBKEM.Decap}(\mathsf{ibkem.sk}, \mathsf{ibkem.ct})$. Then, it computes $m \leftarrow \mathsf{IncSKE.Dec}(\mathsf{inc.sk}, \mathsf{inc.ct})$. It returns m.

Correctness. The correctness is straight-forward from the correctness of NC-IBE scheme and incompressible SKE scheme.

Parameters. The size of a ciphertext for a message m is $|\mathsf{ibkem.ct}| + |\mathsf{inc.ct}|$. If the incompressible SKE scheme has rate-1 (see Theorem 9), then $|\mathsf{inc.ct}| = |m|(1 + o(1)) + \mathrm{poly}(\lambda)$. If the NC-IBE scheme generates ciphertext whose size is independent of the session key, we have $|\mathsf{ibkem.ct}| = \mathrm{poly}(\lambda)$. Therefore, the rate of the incompressible IBE scheme is $1 - o(1)$.

Theorem 15. *Assuming that* IBKEM *is a secure RNC-IB-KEM scheme and* IncSKE *is a incompressible SKE scheme, the above construction is a secure strongly incompressible IBE scheme.*

Proof sketch. We will show that the construction is secure using a sequence of hybrid arguments.

Hybrid H_0: This is the original strongly incompressible IBE security game.

Hybrid H_1: In this game, the IBKEM scheme is changed to simulation mode. To be precise, the first simulator is used generate the master public key ibkem.mpk. The second simulator will be used to handle the key queries. The third simulator will be invoked on id^* to produce the simulated ciphertext ibkem.ct. Finally, for a randomly generated inc.sk, the fourth simulator on input inc.sk will produce the master secret key ibkem.msk. The indistinguishability of H_0 from H_1 follows directly from the security of the RNC-IBE.

We now argue that there is no PPT adversary that can win in H_1 with non-negligible probability. This follows from the security of the incompressible SKE scheme. This is because an adversary $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2)$ that wins in H_1 can be used to build an adversary $\mathcal{B} = (\mathcal{B}_1, \mathcal{B}_2)$ that breaks the security of the underlying incompressible SKE as follows. $\mathcal{B}$ can simulate H_1 using the first two IBKEM simulators upto the challenge phase. On receiving (m_0, m_1) from $\mathcal{A}_1$, it will forward it to the SKE challenger and receives $\mathsf{inc.ct}^*$. It will relay $(\mathsf{ibkem.ct}, \mathsf{inc.ct}^*)$ where ibkem.ct is generated by the third IBKEM simulator. In the second phase, it will receive inc.sk from the SKE challenger and will use the fourth IBKEM simulator will generate ibkem.msk. □

Combining Theorem 15, Theorem 9, Theorem 12, we obtain Theorem 2.

Remark 3. We highlight that a similar approach can be employed to construct incompressible IBE from RNC-IBE and incompressible SKE. In this approach, during the encryption process with inputs m and identity id, a secret key inc.sk of the incompressible SKE is freshly generated and encrypted, producing $\mathsf{ibe.ct} \leftarrow \mathsf{IBE.Enc}(\mathsf{mpk}, \mathsf{id}, \mathsf{inc.sk})$. Then, the message m is encrypted as $\mathsf{inc.ct} \leftarrow \mathsf{IncSKE.Enc}(\mathsf{inc.sk}, m)$, and the final ciphertext is $\mathsf{ct} := (\mathsf{ibe.ct}, \mathsf{inc.ct})$. Note that the size of the ciphertext ct depends on both $|\mathsf{inc.sk}|$ and $|\mathsf{inc.ct}|$ because ibe.ct is an encryption of inc.sk. Therefore, combining Theorem 3 and Theorem 10, we get Theorem 4.

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