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# A New Leakage-Resilient IBE Scheme in the Relative Leakage Model

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**Abstract.** We propose the first leakage-resilient Identity-Based Encryption (IBE) scheme with full domain hash structure. Our scheme is leakage-resilient in the relative leakage model and the random oracle model under the decisional bilinear Diffie-Hellman (DBDH) assumption.

**Key words:** identity based encryption, leakage-resilient, relative leakage, bilinear Diffie-Hellman assumption

## 1 Introduction

Cryptographic schemes are used to be analyzed in an attack model in which the internal secret states are completely hidden from the adversary/attacker. However several works [12, 13] indicated that the attack model fails to capture many attacks in the real world, since the attacker may obtain some partial information about the secret states via various *key leakage attacks*. Therefore it is urgent to design leakage-resilient cryptographic schemes which remain provably secure in the strengthened attack model which takes *key leakage attacks* into account.

Recently, the research community pay a lot of attention to construct IBE schemes with leakage-resilience. Alwen et al. [1] presented three leakage-resilient IBE schemes from the Gentry IBE [10], the Boneh-Gentry-Hamburg IBE [4], and Gentry-Peikert-Vaikuntanathan IBE [11], respectively. Among them, the first scheme is secure in the standard model, while the other two schemes are secure in the random oracle model. Chow et al. [6] gave three new leakage-resilient IBE schemes from the Boneh-Boyen IBE [2], the Waters IBE [16], and the Lewko-Waters IBE [14], respectively. All of them are secure in the standard model.

**Our Contributions.** According to [5], IBE schemes from pairings can be classified into three broad families, the full-domain hash family (e.g.

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Boneh-Franklin IBE [3]), the exponent inversion family (e.g. Gentry-IBE [10]), and the commutative blinding family (e.g. Boneh-Boyen IBE [2]). The existing work [1, 6] have shown that IBE schemes from the exponent inversion family and commutative blinding family can be tailored to be leakage-resilient ones. It is natural to ask if we can strengthen the IBE schemes from the full domain hash family to be leakage-resilient.

We give an affirmative answer to the above question by presenting an IBE scheme with the full domain hash structure based on a variant of Boneh-Franklin IBE [7]. Its leakage-resilient chosen plaintext security can be tightly reduced to the DBDH assumption in the relative leakage model and the random oracle model.

## 2 Preliminaries

**Notations.**  $x \xleftarrow{R} S$  denotes that  $x$  is picked uniformly at random from the set  $S$ . We write PPT for probabilistic polynomial time. By  $\text{negl}(n)$  we denote a negligible function of  $n$ . We denote the bit-wise XOR operation by  $\oplus$ . We denote by  $\mathcal{I}$  the identity space and by  $\mathcal{SK}$  the private key space.

### 2.1 Bilinear Diffie-Hellman Assumption

The decisional BDH (DBDH) assumption [2, 3] is defined via the following game: the challenger runs the bilinear group generator  $\text{GroupGen}(1^\kappa)$  to generate  $(p, \mathbb{G}, \mathbb{G}_T, e)$ , picks four random exponents  $x, y, z, w$  from  $\mathbb{Z}_p$ , then computes  $g^x, g^y, g^z, T_0 = e(g, g)^{xyz}$  and  $T_1 = e(g, g)^{xyw}$ . We denote by  $D$  the tuple  $(p, \mathbb{G}, \mathbb{G}_T, e, g, g^x, g^y, g^z)$ . The challenger picks a random bit  $c$  and gives to the adversary  $\mathcal{B}$  the challenge instance  $(D, T_c)$ . We say  $\mathcal{B}$  succeeds in solving the DBDH problem if it outputs the right guess  $c'$  for  $c$  at the end of the game, whose advantage is defined as:

$$|\Pr[c = c'] - 1/2| = |\Pr[\mathcal{B}(D, e(g, g)^{xyz}) = 0] - \Pr[\mathcal{B}(D, e(g, g)^{xyw}) = 0]|$$

**Definition 2.1** *The  $(t, \epsilon)$ -DBDH assumption holds if no  $t$ -time adversary has at least  $\epsilon$  in solving the DBDH problem in  $\mathbb{G}$ .*

### 2.2 Randomness Extractors

The following notions and primitives will be used in our construction. We refer the readers to [1, 15] for a complement knowledge.

For a random variable  $X$ , we define  $\mathbf{H}_\infty(X) = -\log(\max_x \Pr[X = x])$  as its min-entropy. We use the notion of *average min-entropy* [8] which

captures the remaining unpredictability of a random variable  $X$  conditioned on another random variable  $Y$ , formally defined as

$$\tilde{\mathbf{H}}_{\infty}(X|Y) = -\log(E_{y \leftarrow Y}[\max_x \Pr[X = x|Y = y]])$$

where  $E_{y \leftarrow Y}$  denotes the expected value over all values of  $Y$ .

The average min-entropy measures exactly the optimal probability of guessing  $X$  given knowledge of  $Y$ . The following lemma was proved in [9] regarding average min-entropy:

**Lemma 1.** *For any random variables  $X, Y, Z$ , if  $Y$  has  $2^{\ell}$  possible values, then  $\tilde{\mathbf{H}}_{\infty}(X|(Y, Z)) \geq \tilde{\mathbf{H}}_{\infty}(X|Z) - \ell$ .*

The statistical distance between two random variables  $X, Y$  over a finite domain  $\Omega$  is defined as

$$\mathbf{SD}(X, Y) = \frac{1}{2} \sum_{\omega \in \Omega} |\Pr[X = \omega] - \Pr[Y = \omega]|$$

Same as [1, 6, 15], a main tool used in our construction is the strong randomness extractor, which is formally defined as follows to the setting of the average min-entropy.

**Definition 2.2** *A polynomial-time function  $\text{ext} : \mathbb{G} \times \{0, 1\}^{\mu} \rightarrow \{0, 1\}^m$  is an average case  $(k, \epsilon)$ -strong extractor if for all pairs of random variables  $(X, Y)$  such that  $X \in \mathbb{G}$  and  $\tilde{\mathbf{H}}_{\infty}(X|Y) \geq k$ , we have that*

$$\mathbf{SD}((\text{ext}(X, U_{\mu}), U_{\mu}, Y), (U_m, U_{\mu}, Y)) \leq \epsilon$$

where  $\mathbb{G}$  is a non-empty set, and  $U_{\mu}, U_m$  are two uniformly distributed random variables over  $\{0, 1\}^{\mu}, \{0, 1\}^m$  respectively.

Dodis et al. [8] proved that any strong extractor is in fact an average-case strong extractor, for a proper setting of the parameters:

**Lemma 2.** *For any  $\delta > 0$ , if  $\text{ext}$  is a worst case  $(m - \log(1/\delta), \epsilon)$ -strong extractor, then  $\text{ext}$  is also an average-case  $(m, \epsilon + \delta)$ -strong extractor.*

As a specific example, they proved the following lemma which essentially gives an explicit construction of an average-case strong extractor:

**Lemma 3.** *Let  $X, Y$  be two random variables such that  $X \in \mathbb{G}$  and  $\tilde{\mathbf{H}}_{\infty}(X|Y) \geq k$ . Let  $\mathcal{H} = \{H : \mathbb{G} \rightarrow \{0, 1\}^m\}$  be a family of universal hash functions. If  $m \leq k - 2\log(1/\epsilon)$  then we have*

$$\mathbf{SD}((H(X), U_s, Y), (U_m, U_s, Y)) \leq \epsilon$$

### 2.3 Leakage Model for IBE Setting

In this paper we use the relative leakage model suitable for the IBE setting. The leakage-resilient chosen plaintext security is defined by the following LeakCPA game, which is refined from the CpaLeak game introduced in [6].

**Setup.** The challenger generates the public parameters  $mpk$  and the master secret key  $msk$ . It gives  $mpk$  to the adversary and keeps  $msk$  to itself.

**Phase 1.** The adversary can make one of the following two types of queries to the challenger:

1. Leak( $I, h_i$ ) query, where  $h_i : \mathcal{SK} \rightarrow \{0, 1\}^{\ell_i}$ . The challenger checks if the overall amount leakage will exceed  $\ell$ . If not, it responds with  $h_i(sk)$ . Otherwise it responds with a reject symbol  $\perp$ .
2. Reveal( $I$ ) query, where  $I$  is the identity. The challenger responds with the associated private key  $sk$ .

**Challenge.** The adversary submits two messages  $M_0, M_1$  of equal size and a challenge identity  $I^*$ , with the restriction that  $I^*$  has not been revealed. The challenger picks a random bit  $\beta$  and encrypts  $M_\beta$  under  $I^*$ . It sends the resulting ciphertext  $C^*$  to the adversary.

**Phase 2.** The same as Phase 1 with the restriction that no leakage queries or reveal queries related to  $I^*$  are allowed.

**Guess.** The adversary outputs a bit  $\beta'$ . We say it succeeds if  $\beta = \beta'$ .

The advantage of an adversary  $\mathcal{A}$  on breaking an IBE scheme  $\mathcal{E}$  with security parameter  $\kappa$  and leakage bound  $\ell$  is defined as  $\text{Adv}_{\mathcal{A}, \mathcal{E}}^{\text{CPALeak}}(\kappa, \ell) = |\Pr[\beta = \beta'] - \frac{1}{2}|$ .

**Definition 2.3** *An IBE scheme  $\mathcal{E}$  is  $\ell$ -leakage fully secure if for all PPT adversaries  $\mathcal{A}$  it holds that  $\text{Adv}_{\mathcal{A}, \mathcal{E}}^{\text{CPALeak}}(\kappa, \ell) \leq \text{negl}(\kappa)$ .*

## 3 Our Scheme

Our scheme consists of the following four algorithms:

**Setup.** Run  $\text{GroupGen}(1^\kappa) \rightarrow (p, \mathbb{G}, \mathbb{G}_T, e)$ , pick  $x \xleftarrow{R} \mathbb{Z}_p$ ,  $g_2 \xleftarrow{R} \mathbb{G}^*$ , and a cryptographic hash function  $H : \{0, 1\}^* \rightarrow \mathbb{G}$ . Let  $g_1 = g^x$ ,  $\ell = \ell(\kappa)$  be an upper bound on the amount of leakage. Then set an average-case  $(\log |\mathbb{G}_T| - \ell, \epsilon_{\text{ext}})$ -strong extractor function  $\text{ext} : \mathbb{G}_T \times \{0, 1\}^\mu \rightarrow \{0, 1\}^n$ . The message space is  $\mathcal{M} \in \{0, 1\}^n$ , while  $mpk = (g, g_1, g_2)$  and  $msk = x$ .

**KeyGen.** For a given identity  $I$ , pick  $t \xleftarrow{R} \mathbb{Z}_p$ , compute  $u = H(I)$ , and then generate the private key for  $I$  as  $sk = (d_1, d_2) = (t, (ug_2^{-t})^x)$ .

**Encrypt.** To encrypt a message  $M$  under identity  $I$ , pick an exponent  $r \xleftarrow{R} \mathbb{Z}_p$  and a seed  $s \xleftarrow{R} \{0, 1\}^\mu$  for the extractor function, generate the ciphertext as  $C = (c_1, c_2, c_3, c_4) = (g^r, s, e(g_1, g_2)^r, M \oplus \text{ext}(e(u, g_1)^r, s))$ .

**Decrypt.** To decrypt a ciphertext  $C = (c_1, c_2, c_3, c_4)$  encrypted under  $I$  using the associated private key  $sk = (d_1, d_2)$  to compute  $M = c_4 \oplus \text{ext}(e(c_1, d_2)c_3^{d_1}, c_2)$ . It is easy to verify that if the private key matches, we get the right decryption.

### 3.1 Security Analysis

**Theorem 3.1** *If the DBDH assumption holds and the extractor's second parameter  $\epsilon_{\text{ext}}$  is negligible in  $\kappa$ , then the proposed scheme is  $\ell$ -leakage secure, where  $\ell = \log |\mathbb{G}_T| - k$  and  $k$  is the extractor's first parameter.*

To prove the theorem, we organize the proof as a sequence of games, which are defined as follows:

**Game<sub>Real</sub>**: The real CPALeak game.

**Game<sub>Final</sub>**: The real CPALeak game except in the challenge phase the challenger generates the ciphertext as follows:

$$\begin{aligned} z, w &\xleftarrow{R} \mathbb{Z}_p, \beta \xleftarrow{R} \{0, 1\} & W &= e(u^*, g_1)^z e(g_1, g_2)^{t^*(w-z)} \\ c_1^* &= g^z & c_2^* &\xleftarrow{R} \{0, 1\}^\mu \\ c_3^* &= e(g_1, g_2)^w & c_4^* &= M_\beta \oplus \text{ext}(W, c_2^*) \end{aligned}$$

where  $t^*$  is the tag of private key  $sk^*$  of the challenge identity  $I^*$ ,  $z$  and  $w$  are randomly picked from  $\mathbb{Z}_p$ . The challenge ciphertext is  $C^* = (c_1^*, c_2^*, c_3^*, c_4^*)$ . Note that if  $w \neq z$ , then  $C^*$  is not a valid ciphertext since it is only decrypted correctly when using the private key with tag  $t^*$ .

**Lemma 3.2** *If there exists a PPT algorithm  $\mathcal{A}$  such that  $\text{Adv}_{\mathcal{A}, \mathcal{E}}^{\text{Game}_{\text{Real}}} - \text{Adv}_{\mathcal{A}, \mathcal{E}}^{\text{Game}_{\text{Final}}} = \epsilon$ , then we can build a PPT algorithm  $\mathcal{B}$  with advantage  $\epsilon$  in breaking the DBDH problem.*

*Proof.* Suppose  $\mathcal{B}$  is given a DBDH challenge  $(p, \mathbb{G}, \mathbb{G}_T, e, g, g^x, g^y, g^z, T)$ . We now describe how it interacts with  $\mathcal{A}$  in the following game:

**Setup.**  $\mathcal{B}$  sets  $g_1 = g^x$  (implicitly sets  $msk = x$ ),  $g_2 = g^y$ , picks a suitable extractor function  $\text{ext}$ , then gives  $\mathcal{A}$  the public parameters  $mpk = (p, \mathbb{G}, \mathbb{G}_T, e, g, g_1, g_2, \text{ext})$ .

**Hash queries.** For a fresh hash query on  $I$ ,  $\mathcal{B}$  picks  $a, t \xleftarrow{R} \mathbb{Z}_p$  and responds with  $u = g^a g_2^t$ .

**KeyGen queries.** For an arbitrary identity  $I$ ,  $\mathcal{B}$  computes a private key for it as follows: (1) compute  $u = H(I)$ ; (2) set  $d_1 = t$ ,  $d_2 = g_1^a = (ug_2^{-t})^x = (g^a g_2^t g_2^{-t})^x$ ; (3) return  $sk = (d_1, d_2)$ .

We note that the keygen queries are always implicitly called by  $\mathcal{B}$  when it answers the associated leak queries and reveal queries.

**Phase 1.** To answer the leak queries and reveal queries issued by  $\mathcal{A}$ ,  $\mathcal{B}$  creates two lists  $L$  and  $K$ , which are initially empty.  $L$  is a list of triples of identities, private keys, and a leakage counter, while  $K$  is a list of tuples of identities, private keys.

– **Leak**( $I, h_i$ ) query:  $\mathcal{B}$  checks if there is a tuple  $\langle I, sk \rangle$  in the existing  $K$  list. If it is not  $\mathcal{B}$  runs  $sk \leftarrow \text{KeyGen}(msk, I)$ , inserts the tuple  $(I, sk)$  to the  $K$  list and the triple  $\langle I, sk, 0 \rangle$  to the  $L$  list. After this step there must exist a triple  $\langle I, sk, num \rangle$  in the  $L$  list,  $\mathcal{B}$  checks if  $num + \ell_i \leq \ell$ . If this is true, it responds with  $h_i(sk)$  and sets  $num \leftarrow num + \ell_i$  in  $\langle I, sk, num \rangle$ . Otherwise  $\mathcal{B}$  responds with a reject symbol  $\perp$ .

– **Reveal**( $I$ ) query:  $\mathcal{B}$  checks if there is a tuple  $\langle I, sk \rangle$  in the  $K$  list. If it is  $\mathcal{B}$  responds with  $sk$ . If it is not  $\mathcal{B}$  runs  $sk \leftarrow \text{KeyGen}(msk, I)$ , inserts the tuple  $\langle I, sk \rangle$  to the  $K$  list and the triple  $\langle I, sk, 0 \rangle$  to the  $L$  list, and responds the leak query with  $sk$ .

Notice that  $\mathcal{B}$  can calculate a valid private key for any identity. Therefore,  $\mathcal{B}$  is able to answer all the leakage queries **Leak**( $I, h_i$ ) and reveal queries **Reveal**( $I$ ), with the corresponding private key  $sk = (d_1, d_2)$ .

**Challenge.**  $\mathcal{A}$  submits two messages  $M_0, M_1$  and an identity  $I^*$  on which it want to be challenged to  $\mathcal{B}$ .  $\mathcal{B}$  computes  $sk^* = (d_1^*, d_2^*) = (t^*, g_1^{a^*})$ , then generates the challenge ciphertext as follows:

$$\begin{aligned} \beta &\stackrel{R}{\leftarrow} \{0, 1\} & c_1^* &= g^z \\ c_2^* &\stackrel{R}{\leftarrow} \{0, 1\}^\mu & c_3^* &= T \\ W &= e(c_1^*, d_2^*)(c_3^*)^{d_1^*} = e(g^z, g_1^{a^*})T^{t^*} & c_4^* &= M_\beta \oplus \text{ext}(W, c_2^*) \end{aligned}$$

**Phase 2.** The same as Phase 1.

**Guess.**  $\mathcal{A}$  outputs a guess  $\beta'$ .  $\mathcal{B}$  returns 0 if  $\beta = \beta'$  or 1 if  $\beta \neq \beta'$ .

We will prove that the advantage of  $\mathcal{B}$  in breaking the DBDH problem is  $\epsilon$ . To see this, notice that if  $T = e(g, g)^{xyz}$  the challenge ciphertext is a correct ciphertext according to the original encryption algorithm and thus  $\mathcal{A}$  plays the **Game<sub>Real</sub>**. This is because  $W = e(g^z, g_1^{a^*})T^{t^*} = e(g^{a^*}, g_1^z)e(g_2^{t^*}, g_1^z) = e(g^{a^*}g_2^{t^*}, g_1^z) = e(u^*, g_1)^z$  as one can easily verify. Thus the probability that  $\mathcal{A}$  succeeds in the game is exactly  $\frac{1}{2} + \text{Adv}_{\mathcal{A}, \mathcal{E}}^{\text{Game}_{\text{Real}}}$ . Since  $\mathcal{B}$  outputs 0 when  $\mathcal{A}$  succeeds we get that

$$\Pr[\mathcal{B}(D, e(g, g)^{xyz}) = 0] = \frac{1}{2} + \text{Adv}_{\mathcal{A}, \mathcal{E}}^{\text{Game}_{\text{Real}}}$$

On the other hand if  $T = e(g, g)^{xyw} = c_3^*$  then  $\mathcal{A}$  essentially plays the  $\mathbf{Game}_{\mathbf{Final}}$ , because  $W = e(g^z, g_1^{a^*})T^{t^*} = e(g^{a^*}, g_1^z)e(g_2^{t^*}, g_1^{(w-z)+z}) = e(u^*, g_1)^z e(g_1, g_2)^{t^*(w-z)}$  as one can easily verify. Therefore we have that

$$\Pr[\mathcal{B}(D, e(g, g)^{xyw}) = 0] = \frac{1}{2} + \text{Adv}_{\mathcal{A}, \mathcal{E}}^{\mathbf{Game}_{\mathbf{Final}}}$$

Combining the above equations we get that the advantage of  $\mathcal{B}$  in DBDH is  $|\Pr[\mathcal{B}(D, e(g, g)^{xyz}) = 0] - \Pr[\mathcal{B}(D, e(g, g)^{xyw}) = 0]| = \text{Adv}_{\mathcal{A}, \mathcal{E}}^{\mathbf{Game}_{\mathbf{Real}}} - \text{Adv}_{\mathcal{A}, \mathcal{E}}^{\mathbf{Game}_{\mathbf{Final}}} = \epsilon$ . Therefore we prove the lemma.  $\square$

**Lemma 3.3** *For any PPT adversary  $\mathcal{A}$  we have  $\text{Adv}_{\mathcal{A}, \mathcal{E}}^{\mathbf{Game}_{\mathbf{Final}}} \leq 2\epsilon_{\text{ext}}$ .*

*Proof.* In the  $\mathbf{Game}_{\mathbf{Final}}$ , it is true that  $W = e(u^*, g_1)^z e(g_1, g_2)^{t^*(w-z)}$ , where  $t^*$  is the tag of the private key for  $I^*$ . If we assume that the exact private key with tag  $t^*$  is perfect hidden from the adversary, then  $W$  distributes uniformly at random in  $\mathbb{G}_T$ , and therefore the challenge ciphertext  $C^*$  is totally independent of  $M_\beta$  in an PPT adversary  $\mathcal{A}$ 's view. This is because  $w = z \bmod p$  with negligible probability in  $\kappa$  and  $t^*$  is chosen randomly for  $I^*$ .

Suppose we denote by  $R$  the set of all terms (public parameters, private keys, challenge ciphertext) given to the adversary  $\mathcal{A}$  except the leakage, the random seed  $c_2^*$ , and the part of the challenge ciphertext  $c_4^*$ , then according to the above argument  $\tilde{\mathbf{H}}_\infty(C|R) = \log |\mathbb{G}_T|$ . But the attacker has access to at most  $\ell$  bits of leakage from the private key, i.e. to a random variable  $Y$  with  $2^\ell$  values, thus by lemma 1 we know that

$$\tilde{\mathbf{H}}_\infty(C|(Y, R)) \geq \tilde{\mathbf{H}}_\infty(C|R) - \ell = \log |\mathbb{G}_T| - \ell$$

According to the definition of  $(\log |\mathbb{G}_T| - \ell, \epsilon_{\text{ext}})$ -strong extractor we have that  $\mathbf{SD}(\text{ext}(W, S), S, Y, R), (U_m, S, Y, R)) \leq \epsilon_{\text{ext}}$ , where  $S$  is the random variable for the seed  $c_2^* \in \{0, 1\}^\mu$  distributed uniformly at random,  $Y, R$  are the values of all the random variables known to the adversary: leakage and the rest, respectively. Thus the statistical distance of  $c_4^* = M_\beta \oplus \text{ext}(W, c_2^*)$  from the uniform distribution is at most  $\epsilon_{\text{ext}}$  for each  $\beta$ . The statistical distance between the two possible ciphertexts is at most  $2\epsilon_{\text{ext}}$  and no adversary (even an unbounded one) can distinguish them with advantage more than this.  $\square$

Suppose  $\epsilon_{DBDH}$  is the maximum advantage of all PPT adversaries in the DBDH game. Then according to the above lemma, for any PPT adversary  $\mathcal{A}$  we have  $\text{Adv}_{\mathcal{A}, \mathcal{E}}^{\mathbf{Game}_{\mathbf{Real}}} - \text{Adv}_{\mathcal{A}, \mathcal{E}}^{\mathbf{Game}_{\mathbf{Final}}} \leq \epsilon_{DBDH}$ . Therefore

$$\text{Adv}_{\mathcal{A}, \mathcal{E}}^{\mathbf{Game}_{\mathbf{Real}}} \leq \text{Adv}_{\mathcal{A}, \mathcal{E}}^{\mathbf{Game}_{\mathbf{Final}}} + \epsilon_{DBDH}(\kappa) \leq 2\epsilon_{\text{ext}}(\kappa) + \epsilon_{DBDH}(\kappa)$$

The proposed scheme is leakage-resilient CPA secure if both  $\epsilon_{DBDH}(\kappa)$  and  $\epsilon_{\text{ext}}(\kappa)$  are negligible functions of  $\kappa$ . This proves the theorem.  $\square$

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