Non-Malleable Codes Against Affine Errors

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Abstract—Non-malleable code is a relaxed version of error-correction codes and the decoding of modified codewords results in the original message or a completely unrelated value. Thus, if an adversary corrupts a codeword then he cannot get any information from the codeword. This means that non-malleable codes are useful to provide a security guarantee in such situations that the adversary can overwrite the encoded message. In 2010, Dziembowski et al. showed a construction for non-malleable codes against the adversary who can falsify codewords bitwise independently. In this paper, we consider an extended adversarial model (affine error model) where the adversary can falsify codewords bitwise independently or replace some bit with the value obtained by applying an affine map over a limited number of bits. We prove that the non-malleable codes (for the bitwise error model) provided by Dziembowski et al. are still non-malleable against the adversary in the affine error model.

I. INTRODUCTION

As we know, error-correction codes can recover the original message from a corrupted codeword (within admissible errors) and error-detection codes can detect if the codeword is corrupted while the error correction may not be possible. The notion of non-malleable codes, invented by Dziembowski, Pietrzak, and Wichs [11], is a relaxed notion of error detection codes or error correction codes. (The term “non-malleability” comes from non-malleable cryptography [9].) For non-malleable codes, we suppose that errors would be caused by some adversary’s malicious behaviors. If the adversary tampers a codeword of a non-malleable code, its decoding results in either the original message or an independent message of the original one. Thus, non-malleable codes are applicable to situations where error-detection and error-correction are impossible. For example, they provides a security guarantee against adversaries who can overwrite encoded messages.

We suppose that the adversary tampers a codeword C by applying a function f to C. We consider the situation where a message s ∈ {0, 1}^k is randomly encoded and the encoded message is tampered by f. We denote the resulting corrupted codeword by a random variable Tamper^s_f. For the non-malleability, it is desirable that, for any s, s’ ∈ {0, 1}^k, the random variables Tamper^s_f and Tamper^{s’}_f are almost identical to each other. But, it may happen that the decoding result ˆs for a tampered codeword coincides with the original message s. In this case, it is clear that Tamper^s_f is dependent on s. Thus, we consider a probability distribution D_f whose support includes ˆs and a special symbol same*. By using the above probability distribution, the notion of non-malleability codes can be defined. A code is non-malleable if there exists a probability distribution D_f such that, for any s ∈ {0, 1}^k, the following two probability distributions are statistically indistinguishable: (1) the induced probability distribution from Tamper^s_f and (2) the probability distribution which is the identical to D_f but if same* appears then we replace it with s.

In general, there is no non-malleable code for any tampering functions. In [11], Dziembowski et al. consider a class of bitwise independent tampering functions and give a construction of non-malleable codes with respect to the class of bitwise independent tampering functions. Faust et al. [12] provide efficient non-malleable codes with respect to tampering functions which can be computed by poly-size circuits. Chandran et al. [3] consider block-wise tampering and show the impossibility of non-malleable codes with respect to block-wise tampering in the information theoretic setting. They also give a construction of non-malleable codes with respect to block-wise tampering from the viewpoint of the computational complexity theory. Aggarwal et al. [1] consider more possibility of computational non-malleable codes. In the literature (e.g., [7], [5], [10], [4]), several tampering models are proposed and connections to other research areas such as randomness extractors and locally decodable codes are discussed.

In this paper, we extend bitwise independent tampering to “affine” tampering, where the adversary can falsify codewords bitwise independently or replace some bit with the value obtained by applying an affine map over a limited number of bits. We prove that the non-malleable codes with respect to bitwise independent tampering, provided by Dziembowski et al. [11], are still non-malleable with respect to the affine tampering in the information theoretic setting.

II. NOTATIONS

Let g be a randomized function and g(x; r) be the functional value on input x which can be computed with supplementary randomness r. If we do not have to specify the randomness r, we denote it by g(x). If D is a probability distribution, d ← D means that a value d is chosen according to the probability distribution D. For a finite set B, |B| denote the number of elements in B. For an n-bit string x ∈ {0, 1}^n, w_H(x) denotes the Hamming weight of x. For two strings x and x’ of equal length, d_H(x, x’) = w_H(x, x’) denotes the Hamming distance between x and x’.

SD(X_0, X_1) def =
In this section, we review the previous results by Dziembowski et al. in [11].

Definition 1: (Coding Scheme) A coding scheme is a pair of two functions \((Enc, Dec)\), where \(Enc : \{0, 1\}^k \to \{0, 1\}^n\) is a (randomized) encoding function and \(Dec : \{0, 1\}^n \to \{0, 1\}^k \cup \{\perp\}\) is a deterministic decoding function satisfying that \(\Pr[Dec(Enc(s)) = s] = 1\) for every \(s \in \{0, 1\}^k\).

The desired property for non-malleable codes is discussed in Section I. We give a formal definition of non-malleable codes below.

Definition 2: (Non-malleability) Let \(F\) be a class of tampering functions and \((Enc, Dec)\) be a coding scheme. For each \(f \in F\) and \(s \in \{0, 1\}^k\), define a random variable as follows:

\[
\text{Tamper}_s^f \overset{def}{=} \begin{cases} \emptyset & \text{if } f \text{ is } \text{bitwise independent tampering}, \\ \{c \leftarrow Enc(s); \tilde{c} \leftarrow f(c); \tilde{s} \leftarrow Dec(\tilde{c}); \text{Output } \tilde{s}\} & \text{otherwise}. \end{cases}
\]

The randomness of \(\text{Tamper}_s^f\) comes from the randomness to compute the encoding function \(Enc\). If, for each \(f \in F\) and for each \(s\), there exists a universal probability distribution \(D_f\) over \(\{0, 1\}^k \cup \{\perp, \text{same}^+\}\) such that

\[
\text{Tamper}_s^f \approx \begin{cases} \{\tilde{s} \leftarrow D_f; & \text{if } \tilde{s} = \text{same}^+ \text{ then output } s; \\ \text{otherwise, output } \tilde{s}\} & \text{otherwise}. \end{cases}
\]

then we say that \((Enc, Dec)\) is non-malleable with respect to \(F\). If the statistical distance in the above is bounded by \(\varepsilon\), we say the non-malleable code \((Enc, Dec)\) is \(\varepsilon\)-secure.

Dziembowski et al. [11] showed a non-malleable code against the adversary who can tamper codewords bitwise independently. Their construction is just a combination of algebraic manipulation detection (AMD) codes by Cramer et al. [8] and a linear error-correction secret sharing scheme [11].

Definition 3: (AMD codes [8]) Let \((A, V)\) be a coding scheme, where \(A : \{0, 1\}^k \to \{0, 1\}^n\) is an encoding function and \(V\) is a decoding function. If, for some \(\rho\), for every \(s \in \{0, 1\}^k\) and for \(\Delta \in \{0, 1\}^n \setminus \{0^n\}\), \(\Pr[V(A(m) + \Delta) \neq \perp] \leq \rho\), then we say that \((A, V)\) is an algebraic manipulation detection (AMD) coding scheme of \(\rho\)-security.

Definition 4: (LECSS scheme [11]) Let \((E, D)\) be a coding scheme. Suppose that \((E, D)\) satisfies the following three properties:

Linearity:
For every \(c \in \{0, 1\}^n\) such that \(D(c) \neq \perp\) and for every \(\Delta \in \{0, 1\}^n\), we have the following:

\[
D(c + \Delta) = \begin{cases} \perp & \text{if } D(\Delta) = \perp, \\ D(c) + D(\Delta) & \text{otherwise}. \end{cases}
\]

Distance \(d\):
For every \(c \in \{0, 1\}^n\) whose Hamming weight is less than \(d\), we have \(D(\bar{c}) = \perp\).

Secrecy \(t\):
For any \(s\), let \(C = (C_1, \ldots, C_n) = Enc(s)\) be a random variable, where \(C_i\) is the \(i\)-th bit of \(C\). Then \(\{C_i\}_{1 \leq i \leq n}\) are \(t\)-wise independent. Each (marginal) \(C_i\) is the uniform distribution over \(\{0, 1\}\). Then we say that \((E, D)\) is a \((t, d)\)-linear error-correction secret-sharing (LECSS) scheme.

Bitwise independent tampering can be described as

\[
f(c_1, \ldots, c_n) = (f_1(c_1), \ldots, f_n(c_n)),
\]

where each \(f_i\) is

- the bit-flipping function (i.e., \(f_i(b) = 1 \oplus b\)),
- the identity function (i.e., \(f_i(b) = b\)),
- the 0-constant function (i.e., \(f_i(b) = 0\)), or
- the 1-constant function (i.e., \(f_i(1) = 1\)).

We denote the class of bitwise independent tampering functions by \(F_{BIT}\). That is,

\[
F_{BIT} = \left\{ f = (f_1, \ldots, f_n) : f_i \in \{\text{bit-flipping, identity, 0-constant, 1-constant}\} \right\}.
\]

Theorem 5: ([11]) Suppose that \((E, D)\) is a \((d, t)\)-LECSS scheme where \(d > n/4\) and \((A, V)\) is a \(\rho\)-secure AMD coding scheme. By using these schemes, we define a coding scheme \((End, Dec)\) as follows:

\[
\begin{align*}
\text{Enc}(s) &= E(A(s)); \\
\text{Dec}(c) &= \begin{cases} \perp & \text{if } D(c) = \perp, \\ V(D(c)) & \text{otherwise}. \end{cases}
\end{align*}
\]

Then, \((Enc, Dec)\) is \(\varepsilon\)-secure non-malleable with respect to \(F_{BIT}\), where \(\varepsilon \leq \max(\rho, 2^{-O(1)})\).

IV. MAIN RESULTS

In this paper, we show that Dziembowski’s non-malleable code with respect to \(F_{BIT}\) is also non-malleable with respect to a class of affine tampering functions, which is a generalization of \(F_{BIT}\). Informally speaking, the class of affine tampering functions includes all the bitwise independent tampering functions and also includes functions \(f\) such that \(\tilde{c}_i = f(c_1, c_2) = c_1 \oplus c_2 \oplus 1\), where bits at some positions are altered into a sum of several bits and some constant. Here, we define a new function: \(f_i\), is said to be \(\ell\)-affine if \(f_i(b_1, \ldots, b_n) = \bigoplus_{j \in B} b_j \perp b\) for some bit \(b \in \{0, 1\}\) and some set \(B \subseteq \{1, \ldots, n\}\) such that \(|B| \leq \ell\). We define a class
of affine tampering functions as follows:

$$F_{f,\text{AFFINE}} = \left\{ f = (f_1, \ldots, f_n) : f_i \in \left\{ \text{bit-flipping, identity, 0-constant, 1-constant, } \ell\text{-affine} \right\}, \text{and all } \ell\text{-affine functions are } \ell\text{-wise independent} \right\}$$

where functions $g_1, \ldots, g_k$ are said to be $\ell$-wise independent if their functional values on the uniform random inputs are $\ell$-wise independent.

**Remark:** For each $\ell$-affine function $f_i(b_1, \ldots, b_n) = \left( \bigoplus_{j \in E} b_j \right) \oplus b_i$, there is the corresponding vector $\beta_i = (a_1, \ldots, a_n)$, where $a_j = 1$ if $j \in B$ and $a_j = 0$ otherwise. Note that $w_H(\beta_i) \leq \ell$. To choose $\ell$-wise independent functions, we first choose vectors $\beta_1, \ldots, \beta_k$ such that $\text{rank}(\beta_1, \ldots, \beta_k) \geq \min\{k, \ell\}$. From such vectors $\beta_1, \ldots, \beta_k$, we can construct $k$ $\ell$-affine functions which are $\ell$-wise independent.

**Theorem 6:** Suppose that $(E, D)$ is a $(d, t)$-LECSS scheme where $d > 3n/8$ and $(A, V)$ is a $\rho$-secure AMD coding scheme and define a coding scheme $(\text{Enc}, \text{Dec})$ as follows:

$$\begin{align*}
\text{Enc}(s) &= E(A(s)); \\
\text{Dec}(c) &= \begin{cases} \\
\perp & \text{if } D(c) = \perp, \\
V(D(c)) & \text{otherwise}.
\end{cases}
\end{align*}$$

Then $(\text{Enc}, \text{Dec})$ is $\varepsilon$-secure non-malleable with respect to $F_{f,\text{AFFINE}}$, where $\varepsilon \leq \max(\rho, 2^{-\Omega(t)})$.

In the proof in [11] that $(\text{Enc}, \text{Dec})$ stated in Theorem 5 is non-malleable with respect to $F_{\text{BIT}}, \{1, \ldots, n\}$ is partitioned into two subsets $B_1$ and $B_2$, where $B_1 = \{i : f_i \text{ is either 0-constant or 1-constant}\}$ and $B_2 = \{i : f_i \text{ is bit-flipping or identity}\}$. They considered several cases with respect to $|B_1|$ and $|B_2|$ and analyzed the security for each case. We partition $\{1, \ldots, n\}$ into three subsets (say, $B_1, B_2$ and $B_3$) and consider several cases with respect to $|B_1|, |B_2|$ and $|B_3|$.

**Proof:** We show that $(\text{Enc}, \text{Dec})$ is non-malleable with respect to $F_{f,\text{AFFINE}}$ and its security $\varepsilon$ satisfies

$$\varepsilon \leq \max \left( \rho, \frac{1}{2^t} + \left( \frac{t}{n(d/n - 3/8)} \right)^{1/2} \right)$$

for any even $t > 6$. We let $f = (f_1, \ldots, f_n)$ be a tampering function in $F_{f,\text{AFFINE}}$ and define a universal distribution $D_f$ for showing that $(\text{Enc}, \text{Dec})$ is non-malleable with respect to $F_{f,\text{AFFINE}}$.

For any message $s \in \{0, 1\}^k$, we consider several probability distributions and use the following notations: $C^s := E\text{nc}(s)$, $\hat{C}^s := f(C^*)$, $\Delta^* := C^s - C^*$, $\hat{S}^s := \text{Dec}(\hat{C}^s)$. $C_i^s$, $\hat{C}_i^s$ and $\Delta_i^s$ for each $i \in \{1, \ldots, n\}$ denote the $i$-th bit of $C^s$, $\hat{C}^s$ and $\Delta^*$, respectively. We partition $i \in \{1, \ldots, n\}$ into three subsets $B_1, B_2$ and $B_3$ as follows: $B_1 = \{i : f_i \text{ is 0-constant or 1-constant}\}$, $B_2 = \{i : f_i \text{ is bit-flipping or identity}\}$ and $B_3 = \{i : f_i \text{ is } \ell\text{-affine}\}$. We let $p = |B_1|$, $q = |B_2|$ and $r = |B_3|$, which satisfy $p + q + r = n$. We define a probability distribution $\hat{S}^s$ as follows: First, sample $\hat{s}$ as $\hat{s} \leftarrow D_f$. If $\hat{s}$ is $\text{same}^*$ then output $s$ instead of $\text{same}^*$. Otherwise, output $\hat{s}$ as it is. We will construct $D_f$ such that, for any $s$, $SD(\hat{S}^s, \text{Patch}(D_f, s)) \leq \varepsilon$. Before discussing each case, we need some useful property:

**Fact:** If $i \in B_3$ then $\hat{C}_i^s$ is the uniform distribution over $\{0, 1\}$ and the joint distribution $\{C_i^s\}_{i \in B_3}$ is $t$-wise independent because of $t$-secrecy of the LECCS scheme and $t$-wise independence of affine functions in $F_{f,\text{AFFINE}}$ for any $s$.

**Case 1:** $p \leq t - r$

We show that $\Delta^*$ for each $s$ is identical to $\Delta^*$ for any other $s$.

- If $f_i$ is the identity function, then we have $\Delta_i^* = 0$. If $f_i$ is bit-flipping, then we have $\Delta_i^* = 1$.
- If $i \in B_1 \cup B_3$ then $\Delta_i^*$ is the uniform distribution over $\{0, 1\}$, since $|B_1 \cup B_3| = |B_1| + |B_3| = p + r \leq t$ imply that $\{C_i^s\}_{i \in B_1 \cup B_3}$ is $t$-wise independent. Thus, we have $\Delta_i^* = \hat{C}_i^s - C_i^s$ is the uniform distribution regardless of $s$.

Therefore, there exists a universal probability distribution $\Delta$ such that $\Delta = \Delta^*$ for any $s$ and we have

$$\begin{align*}
\hat{S}^s &= \text{Dec}(\hat{C}^s) \\
&= V(D(C^* + \Delta^*)) \\
&= V(D(C^* + D(\Delta^*))) \\
&= V(A(s) + D(\Delta^*)) \\
&= V(A(s) + D(\Delta)),
\end{align*}$$

where (1) is by the linearity of the LECCS scheme.

1) If $D(\Delta) \neq 0$ then the security of AMD codes imply that $\Pr[S^s = \perp] \geq 1 - \rho$.  

2) If $D(\Delta) = 0$ then we have $\Pr[S^s = s] = 1$.

From 1) and 2), we define $D_f$ as follows: First, sample $\delta$ as $\delta \leftarrow D_f$. If $D(\delta) = 0$ then output $\text{same}^*$. Otherwise, output $\perp$. Then, we have $SD(\hat{S}^s, \text{Patch}(D_f, s)) \leq \rho$ for any $s$. This completes the proof in Case 1.

**Case 2:** $p \geq n - t$

In this case, we show that $\hat{C}^s$ for each $s$ is identical to $\hat{C}^{s'}$ for any other $s$.

- If $f_i$ is 0-constant, then $\hat{C}_i^s = 0$. If $f_i$ is 1-constant, then $\hat{C}_i^s = 1$.
- For any $i \in B_2 \cup B_3$, $\hat{C}_i^s$ is the uniform distribution over $\{0, 1\}$, since $p \geq n - t$ implies that $|B_2 \cup B_3| = |B_2| + |B_3| = q + r \leq t$. Thus, we can say that $\{\hat{C}_i^s\}_{i \in B_2 \cup B_3}$ is the uniform distribution and $\{\hat{C}_i^s = f(C_i^s)\}_{i \in B_2 \cup B_3}$ are independent uniform distributions for any $s$.

Furthermore, there exists a universal distribution $\hat{C}$ such that $\tilde{C} = \hat{C}^s$ for any $s$ and we have $\hat{S}^s = \text{Dec}(\hat{C}^s) = \text{Dec}(\hat{C})$. We define the distribution $D_f$ which samples $\hat{C}$ as above and computes $\text{Dec}(\hat{C})$. This implies that $SD(\hat{S}^s, \text{Patch}(D_f, s)) = SD(\hat{S}^s, D_f) = 0$ for any $s$. This completes the proof in Case 2.
Case 3: $t - r < p \leq (n - r)/2$

In this case, we show that a probability distribution that always outputs $\perp$ is a universal distribution $D_f$. Since, for any $s$,

$$\Pr[\hat{S}^* \neq \perp] = \Pr[\text{Dec}(\hat{C}^*) \neq \perp] = \Pr[D(\Delta^*) \neq \perp],$$

it suffices to show that $\Pr[D(\Delta^*) \neq \perp]$ is small. Let $\{\Delta_i^*\}_{i \in B_2}$ be any value which is consistent with the fixed bits of $\Delta$ so that $\{\Delta_i^* = \delta_i^*\}_{i \in B_2}$ and for which $D(\delta^*) \neq \perp$. If no such value exists then we are done since $D(\Delta^*) = \perp$ with probability 1. So let us assume that some such value exists. Since $t < p + r \leq (n + r)/2$, $\{\Delta_i^*\}_{i \in B_1 \cup B_3}$ are t-wise independent uniform distributions and we have $\Pr[\Delta^* = \delta^*] \leq 1/2^t$. On the other hand, we show that $d_H(\Delta^*, \delta^*)$ is not so large. The expected value of the Hamming distance between $\Delta^*$ and $\delta^*$ satisfies the following.

$$\mathbb{E}[d_H(\Delta^*, \delta^*)] = \mathbb{E}\left[ \sum_{i=1}^{n} d_H(\Delta_i^*, \delta_i^*) \right]$$

$$= \mathbb{E}\left[ \sum_{i \in B_1 \cup B_3} d_H(\Delta_i^*, \delta_i^*) \right]$$

$$= \sum_{i \in B_1 \cup B_3} \mathbb{E}[d_H(\Delta_i^*, \delta_i^*)]$$

$$= \frac{p + r}{2}. \quad (4)$$

In the above, (2) holds since $\Delta^* = \delta^*$ for $i \in B_2$ and thus $\{d_H(\Delta_i^*, \delta_i^*)\}_{i \in B_2} = 0$. (3) is by the linearity of the expectation. For (4), since $\Delta_i^*$ are independent for $i \in B_1 \cup B_3$, we consider the probability that $d_H(\Delta^*, \delta^*) = \sum_{i \in B_1 \cup B_3} d_H(\Delta_i^*, \delta_i^*)$ is larger than $d$. Since $\{d_H(\Delta_i^*, \delta_i^*)\}_{i \in B_1 \cup B_3}$ are t-wise independent, we can apply a Chernoff-Hoeffding tail bound as in [2], [13]. Thus, we have

$$\Pr[d_H(\Delta^*, \delta^*) \geq d] \leq \Pr\left[ d_H(\Delta_i^*, \delta_i^*) - \frac{p + r}{2} \geq d - \frac{p + r}{2} \right]$$

$$\leq \left( \frac{nt}{(d - \frac{p + r}{2})^2} \right)^{t/2} \leq \left( \frac{nt}{(d - \frac{n - t}{2})^2} \right)^{t/2} \leq \left( \frac{t}{n(\frac{4 - t}{2})^2} \right)^{t/2}. \quad (6)$$

In the above, (5) follows from Lemma 2.2 in [2] by Bellare and Rompel. For (6), we use $r < n/2$ since $r \leq t$. Hence, we have

$$\Pr[D(\Delta^*) \neq \perp] \leq \Pr[\Delta^* = \delta^* \lor d_H(\Delta^*, \delta^*) \geq d]$$

$$\leq \frac{1}{2^t} + \left( \frac{t}{n(\frac{4 - t}{2})^2} \right)^{t/2}$$

and this completes the proof in Case 3.

Case 4: $(n - r)/2 < p \leq n - t$

In this case, we show that a probability distribution that always outputs $\perp$ suffices for a universal distribution $D_f$. To this end, we show that the probability that $\Pr[\hat{S}^* \neq \perp] = \Pr[D(\Delta^*) \neq \perp]$ is small for any $s$. Since $(n - r)/2 < p \leq n - t$, we $t < q + r < (n + r)/2$ and thus $\{\Delta_i^*\}_{i \in B_2}$ are fixed by $f$. Let $\tilde{c} \in \{0, 1\}^n$ be any value which is consistent with the fixed portion of $\hat{C}^*$ so that $\{\tilde{C}_{i}^* = \tilde{c}_{i}^*\}_{i \in B_1}$. If no such value exist then we are done. Otherwise, we can use the similar discussion as in Case 3 and we have

$$\Pr[d_H(\hat{C}^*, \tilde{c}^*) \geq d] \leq \Pr[\hat{C}^* = \tilde{c}^* \lor d_H(\hat{C}^*, \tilde{c}^*) \geq d]$$

$$\leq \frac{1}{2^t} + \left( \frac{t}{n(\frac{4 - t}{2})^2} \right)^{t/2}.$$ 

This complete the proof in Case 4.

For any cases of $p, q, r$, we have completed the proof. Thus, we can say that Theorem 6 holds.

**Remark:** In Theorem 6, we use a $(d, t)$-LECSS code where $d > 3n/8$. This requires that an LECSS code for non-malleability with respect to $F_{I - AFFINE}$ must be better than ones with respect to $F_{BIT}$. Chen et al. [6] have shown the existence such LECSS codes.

V. **CONCLUDING REMARKS**

We have extended the bitwise independent tampering to the affine tampering and shown that the non-malleable codes in [11] with respect to the bitwise independent tampering is also non-malleable with respect to the affine tampering. Our tampering model for the affine tampering may be a bit artificial because of some technical reason. As mentioned, the property of being “affine” is useful to construct t-wise independent functions. But, this does not rule out the possibility to construct t-wise independent functions from non-affine tampering functions. Thus, in future, we may find a wider class of tampering functions for which there exists a non-malleablebde coding scheme.

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