





Fig. 1. The Gaussian two-hop link. Phase 1 is indicated by solid line, and phase 2 by dashed line. R/E: Relay/Eavesdropper. For the adder model, simply remove all noise and restrict signals to integers in the range  $[0, B - 1]$ .

## II. SYSTEM MODEL

We consider a two-hop network with two-phase communications as shown in Figure 1. Node 1 wants to send a confidential message  $W$  to node 2 via a relay node that is untrusted. In [7], we showed that by allowing node 2 to transmit, a two-phase protocol can be developed to provide secrecy. In this protocol, during the first phase, both nodes 1 and 2 transmit. We use  $X_i$  to denote the signal transmitted by node  $i$ . In the Gaussian model,  $X_i$  is continuous and the signal received by the relay is denoted by  $Y_r = X_1 + X_2 + Z_r$ . In phase two, the relay node transmits  $X_r$ . Let  $Y_2$  denote the signal received by node 2 during phase two,  $Y_2 = X_2 + Z_R$ .  $Z_r$  and  $Z_R$  are independent Gaussian random variables with zero mean and unit variance. For simplicity, we assume that each node has an average power constraint  $P$  and the channel gain of each link is unity.

The discrete adder model is very similar to the Gaussian model. The differences are that now  $X_i, i = 1, 2$  and  $X_r$  are integers in the range  $[0, B - 1]$ , where  $B$  is an integer and  $B \geq 2$ ; the noise terms are removed. Hence, we have  $Y_r = X_1 + X_2$  and  $Y_2 = X_r$ .

## III. REVIEW OF KNOWN RESULTS

### A. AMD Code [10]

Let the input to an AMD encoder be denoted by  $s$ . It is assumed that  $s$  is a  $1 \times d$  vector. Each component of it is taken from a finite field  $\mathbf{GF}(q^m)$ . The output of the encoder is given by the tuple  $\{s, x, h\}$ , where  $x$  is uniformly distributed over  $\mathbf{GF}(q^m)$  and is independent from  $s$ , and  $h$  is computed according to the hash rule:  $h = x^{d+2} + \sum_{i=1}^d s_i x^i$ , where  $s_i$  is the  $i$ th component of  $s$  and the addition and multiplication is defined over  $\mathbf{GF}(q^m)$ .  $h$  is usually called the ‘‘hash tag’’. Suppose the node 2 receives  $s', x', h'$ , where  $s' \neq s$ . Let  $\Delta_x = x' - x$ .  $\Delta_h = h' - h$ . Then [10] has the following result:

*Theorem 1:* [10, Theorem 2] If the distribution of  $x$  conditioned on  $\{\Delta_x, \Delta_h, s', s\}$  is uniform over the field  $\mathbf{GF}(q^m)$ ,  $q$  being a prime, and  $d + 2$  is not divisible by  $q$ , then the probability that the hash rule holds is bounded by  $\frac{d+1}{q^m}$ .

### B. Limits of AMD Code

To motivate the next section, we briefly describe the difficulty in applying AMD code in our setting. Consider the discrete adder model. It is immediate that a direct application

to this model is impossible due to the fact  $X_r^n$  is computed from  $Y_r^n = X_1^n + X_2^n$ , where the component-wise ‘‘+’’ means real addition rather than modulus addition. Hence  $Y_r^n$  is not independent from  $X_1^n$  and the condition in Theorem 1 is not fulfilled.

To solve this problem, we consider using a serial concatenation of the wire-tap code and the AMD code. Let  $f$  be the stochastic encoding function of the wire-tap code. Then  $Y_r^n$  contains little information from the input to the wire-tap code encoder, which in this case is an AMD codeword  $\{s, x, h\}$ , if the wire-tap code is designed according to [9].

Let the input to the decoder of the wire-tap code be  $\hat{X}_1^n$ . Then in the case of the discrete adder model, it is given by

$$\hat{X}_1^n = X_r^n - X_2^n \bmod B \quad (1)$$

If the relay is honest, then  $\hat{X}_1^n$  should equal  $X_1^n$ , which is mapped to  $f^{-1}(X_1^n)$ . Otherwise, the decoder will receive  $f^{-1}(X_r^n - X_2^n \bmod B)$ . The difference is hence given by

$$f^{-1}(X_r^n - X_2^n \bmod B) - f^{-1}(X_1^n) \quad (2)$$

Note that (2) corresponds to  $\Delta_x$ , and  $f^{-1}(X_1^n)$  corresponds to  $x$  in Theorem 1, and even though the adversary has little information about  $f^{-1}(X_1^n)$ , in general (2) is *not* independent from  $f^{-1}(X_1^n)$ .

This problem can be alleviated if  $f^{-1}$  is linear. In this case, (2) is given by

$$f^{-1}((X_r^n - (X_1^n + X_2^n) \bmod B) \bmod B) \quad (3)$$

Note that  $X_r^n$  is computed from  $Y_r^n = (X_1^n + X_2^n) \bmod B$  can also be computed from  $Y_r^n$ . If  $Y_r^n$  is almost independent from the confidential message then (3) is almost independent from  $f^{-1}(X_1^n)$ . One case where  $f^{-1}$  is linear is the Type II wire-tap channel.

### C. Type II Wire-tap Channel

In the Type II wire-tap channel model [13], the main channel is a noiseless binary channel and the eavesdropper channel is a binary erasure channel with erasure probability  $\epsilon$ . It was shown in [11] that the dual code of a ‘‘good’’ linear code can be used to construct codes for the Type II wire-tap channel as follows.

Let  $H$  denote the parity check matrix of a linear block code  $\mathcal{C}$ . Let the input to the encoder be  $\mathbf{s}$ , and the output of the encoder be  $\mathbf{x}$ . Here each component of  $H$ ,  $\mathbf{x}$ ,  $\mathbf{s}$  is taken from  $\mathbf{GF}(2)$ . The encoder is defined by  $\mathbf{x} \leftarrow \{\mathbf{x} : H\mathbf{x} = \mathbf{s}\}$ , which means  $\mathbf{x}$  is randomly chosen from the set  $\{\mathbf{x} : H\mathbf{x} = \mathbf{s}\}$  under a uniform distribution. The decoder is simply:

$$f^{-1}(\mathbf{x}) = H\mathbf{x} \quad (4)$$

To find a good  $H$ , [11] looks at the generation matrix  $G$  of  $\mathcal{C}$ . Let  $\mathbf{y}$  be the output of the erasure channel. Let  $G_e$  be formed by the columns of  $G$  that corresponds to the erased positions. Define  $F_0$  as a binary random variable such that  $F_0 = 1$  denote the event that  $G_e$  has full column rank. Then it was shown in [13] [11, Theorem 2] the following result holds:

*Theorem 2:*  $I(\mathbf{s}; \mathbf{y} | F_0 = 1) = 0$

*Remark 1:* Note that this theorem holds even if the encoder only uses a subset of all possible input sequences.  $\square$

We next evaluate the probability  $\Pr(F_0 = 0)$ . To make  $\Pr(F_0 = 0)$  small, [11] proposes to use the parity check matrix  $H_L$  of a good LDPC code  $\mathcal{C}_L$  for binary erasure channel as  $G$ . It can be shown that  $\Pr(F_0 = 0)$  can be bounded via the decoding error probability  $P_e$  of  $\mathcal{C}_L$  as  $\Pr(F_0 = 0) \leq 2P_e$ . In [14, Theorem 6], a lower bound is provided on the error exponent  $-\frac{1}{L} \log_2 P_e$ . By evaluating [14, (43)], we find that the expurgated regular (3, 4) LDPC code ensemble, whose rate is 1/4, has a positive error exponent on average when the channel erasure probability is 1/2. Hence there must exist a (3, 4) LDPC code that the decoding error probability decreases exponentially fast with code length. The corresponding wire-tap code has rate 1/4. Let  $L$  be the length of the wire-tap code. Then  $\Pr(F_0 = 0)$  can be bounded as:

$$\Pr(F_0 = 0) < \exp(-L\alpha) \quad (5)$$

for some positive  $\alpha > 0$ .

#### D. Simulating a Binary Erasure Channel

In order to leverage the linear wire-tap code from Section III-C, we have to simulate a binary erasure channel in both the adder model and the Gaussian model. In the adder model, this could be done simply by restricting the transmitted signal to be binary. Doing so leads to a significant rate loss. However, since only  $\{x, h\}$  needs to be transmitted in this fashion, as we will see later, the overall rate loss is negligible.

In the Gaussian model, we use every  $N$  channel use to simulate 1 channel use in a binary erasure channel via a simple repetition code. Note that this leads to the loss of coding gain and to an arbitrarily small rate. However, since we only use this scheme to transmit  $\{x, h\}$ , which also has an arbitrarily small rate, the overall rate loss will be shown to be small as well in the sequel.

The relaying and the signaling scheme under the repetition code is described as following: Let  $\mathbf{1}$  denote an  $n$ -bit vector of ones. Let  $\mathbf{0}$  denote an  $n$ -bit vector of zeros. Node 1 sends  $\sqrt{P}\mathbf{1}$  as 1 and  $-\sqrt{P}\mathbf{1}$  as 0. Node 2 sends  $\sqrt{P}\mathbf{1}$  or  $-\sqrt{P}\mathbf{1}$  with equal probability. Let  $u_i^N$  be the signal transmitted by node  $i$ . Define  $\oplus$  over  $\{-\sqrt{P}\mathbf{1}, \sqrt{P}\mathbf{1}\}$  so that this set forms a modulus 2 group  $G(2)$ . Define  $\mathbf{I}_1$  as the isomorphism from  $\{\sqrt{P}\mathbf{1}, -\sqrt{P}\mathbf{1}\}$  to  $G(2)$ .

The relay node receives  $u_1^N + u_2^N + Z_r^N$ , which is a vector from the set  $\{-2\sqrt{P}\mathbf{1}, \mathbf{0}, 2\sqrt{P}\mathbf{1}\}$  corrupted by Gaussian noise. Note that this is a degraded version of  $u_1^N + u_2^N$ . Hence if a wire-tap code supports the notion of secrecy as in Theorem 2, the same result holds here as well due to data processing inequality. This means that, we can design the wire-tap code assuming the adversary receives  $u_1^N + u_2^N$  instead. Note that if  $u_1^N + u_2^N = \mathbf{0}$ , then the conditional probability distribution of the signal transmitted by node 1 is the same as this distribution without the conditioning. Hence an erasure has occurred.

The relay scheme is composed of the following two steps:

- 1) It uses a maximum likelihood detector to determine  $\{-2\sqrt{P}\mathbf{1}, \mathbf{0}, 2\sqrt{P}\mathbf{1}\}$ .
- 2) It transmits  $-\sqrt{P}\mathbf{1}$  if it detects  $-2\sqrt{P}\mathbf{1}$  or  $2\sqrt{P}\mathbf{1}$ . It transmits  $\sqrt{P}\mathbf{1}$  if it detects  $\mathbf{0}$ .

Node 2 uses a maximum likelihood detector to determine whether the relay transmits  $-\sqrt{P}\mathbf{1}$  or  $\sqrt{P}\mathbf{1}$ . Let the detection result be  $u_r^N$ . It then recovers the binary bit sent by node 1 from  $\mathbf{I}_1(u_r^N) \oplus \mathbf{I}_1(u_2^N)$ . If the relay is honest and all decoding operation is successful, it should equal  $\mathbf{I}_1(u_1^N)$ . Their difference is given by

$$\mathbf{I}_1(u_r^N) \oplus \mathbf{I}_1(u_1^N \oplus u_2^N) \quad (6)$$

Note that  $u_r^N$  is computed from the received signal by the relay node  $Y_r^N$  and the noise  $Z_r^N$ .  $\mathbf{I}_1(u_1^N \oplus u_2^N)$  can be computed from  $Y_r^N$  as well if the detection operation at the relay is successful. With the repetition code, the detection error probability decreases exponentially fast with  $N$ . As we have shown in Section III-C, with high probability  $Y_r^N$  is independent from the confidential message  $\{x, h\}$  conveyed by the linear wire-tap code. Hence, with high probability (6) is independent from  $\{x, h\}$ .

## IV. MAIN RESULTS

The overall achievability scheme can then be summarized as follows: A string of confidential messages  $s$  is first encoded into an AMD code tuple  $\{s, x, h\}$ , which is transmitted in two stages:

- 1) During the first stage,  $\{x, h\}$  is transmitted using the linear wire-tap code in Section III-C via the simulated Type II wire-tap channel described in Section III-D. The transmission is done via the two phase protocol described in Section II.
- 2) During the second stage,  $s$  is transmitted using a randomly generated wire-tap code and the two phase protocol described in Section II. This wire-tap code from [9] is nonlinear but offers a higher rate.

Let the overall number of channel uses be  $N_T$ . Let  $R_e$  denote the achievable *strong* secrecy rate for these two models when the relay is honest. For the discrete adder model, it can be shown that  $R_e = H(X_1) - I(X_1; X_1 + X_2)$  which equals 0.5 when  $B = 2$  and is lower bounded by  $\log_2 B - 0.8$  in general. For the Gaussian model, it can be shown that combining lattice code [12] and the wiretap code from [9],  $R_e = \left[\frac{1}{2} \log_2 \left(\frac{1}{2} + P\right) - 1\right]^+$ .<sup>1</sup>

*Remark 2:* The reason for using lattice codes in the Gaussian model is in effect to avoid the continuous channel outputs in the Gaussian case. This is needed in order to apply the result of [9]. Also note that we can not use the approach in [15] to achieve strong secrecy as [15] requires transmitting a negligible amount of random bits via an *authenticated* public channel for privacy amplification. In our case, all bits must be transmitted via the relay, which is *not* authenticated.  $\square$

<sup>1</sup>Due to the half duplex constraints, rates should be multiplied by the time sharing factor 1/2.

*Theorem 3:* For the discrete adder channel in Section II, for a rate smaller but arbitrarily close to  $R_e$ , there exists  $\alpha_i > 0, i = 1, 2$  such that

- 1) When the relay is honest, data can be reliably transmitted at this rate with the mutual information between the adversary's knowledge and the confidential message decreases at the rate of  $O(\exp(-\alpha_1 N_T))$ .
- 2) The probability that a counterfeit message is accepted by node 2 decreases at the rate of  $O(\exp(-\alpha_2 N_T))$

*Theorem 4:* For the Gaussian model in Section II, for a rate smaller but arbitrarily close to  $R_e$ : Then there exists  $\alpha_i > 0, i = 3, 4, 5$  such that

- 1) When the relay is honest, data can be transmitted at the rate such that the mutual information between the adversary's knowledge and the confidential message decreases at the rate of  $O(\exp(-\alpha_3 N_T^{1/2}))$ . The error probability decreases at the speed of  $O(\exp(-\alpha_4 N_T^{1/2}))$ .
- 2) When the relay is not honest, the probability that a counterfeit message is accepted by node 2 decreases at the rate of  $O(\exp(-\alpha_5 N_T^{1/2}))$ .

## V. PROOFS OF THE MAIN RESULTS

The following notation is used:  $X_i(j), i = 1, 2, X_r(j)$  denote the signals transmitted by node 1, 2 and the relay during the  $j$ th stage,  $j = 1, 2$ . Similarly,  $Y_i(j), i = 1, 2, Y_r(j)$  denote the signals received during the  $j$ th stage.

Note that in order to use AMD code, as shown in Theorem 1, we need to prove  $x$  is almost independent from  $\{\Delta_h, \Delta_x, s, s'\}$ . Define  $F$  as a binary random variable such that  $F = 1$  indicates that (1) the relay node can decode the binary modulus sum of the signals transmitted by node 1 and 2 during stage one and (2) the binary linear wire-tap code does not leak information. Then for the discrete adder model,  $\Pr(F = 0)$  is bounded by (5), since the channel is noiseless. For the Gaussian model, let  $\gamma$  denote the error exponent of the detection operation. Then  $\Pr(F = 0)$  can be bounded as:

$$\Pr(F = 0) \leq L \exp(-N\gamma) + \exp(-L\alpha) \quad (7)$$

where  $L$  is the length of the binary wiretap code described in Section III-C. The first term in (7) is the union bound on the probability that any detection operation at the relay or node 2 is not successful. The second term comes from the imperfectness of the linear wire-tap code. For both models, we have the following lemma:

*Lemma 1:*  $I(x; \Delta_h, \Delta_x, s, s') \leq m \Pr(F = 0)$

*Proof Outline:*

$$I(x; \Delta_h, \Delta_x, s, s') \leq I(x; \Delta_h, \Delta_x, s, s', F) \quad (8)$$

$$= I(x; \Delta_h, \Delta_x, s, s' | F) \quad (9)$$

$$= \Pr(F = 1) I(x; \Delta_h, \Delta_x, s, s' | F = 1) + \Pr(F = 0) I(x; \Delta_h, \Delta_x, s, s' | F = 0) \quad (10)$$

(9) is due to the fact that  $F$  is solely determined by  $X_2(1)$  and the structure of the code. Hence  $I(x, F) = 0$ . The second

term in (10) is upper bounded by  $m \Pr(F = 0)$ . For the first term, we have:

$$I(x; \Delta_h, \Delta_x, s, s' | F = 1) \quad (11)$$

$$= I(x; \Delta_h, \Delta_x, s' | F = 1, s) + I(x; s | F = 1) \quad (12)$$

$$= I(x; \Delta_h, \Delta_x, s' | F = 1, s) \quad (13)$$

$$\leq I(x; Y_r(1), \Delta_h, \Delta_x, s' | F = 1, s) \quad (14)$$

$$= I(x; Y_r(1), s' | F = 1, s) \quad (15)$$

$$\leq I(x; Y_r(1), Y_2(2), X_2(2), s' | F = 1, s) \quad (16)$$

$$= I(x; Y_r(1), Y_2(2), X_2(2) | F = 1, s) \quad (17)$$

(15) is because based on (3) and (6), we observe that, conditioning on  $F = 1, \{\Delta_x, \Delta_h\}$  can be computed from  $Y_r(1)$  by the relay. (17) is due to the fact  $s'$  can be computed from  $Y_2(2), X_2(2)$  by node 2. It can then be shown that (17) equals 0. For details, the reader is referred to [16]. ■

In the following, we use ‘‘HRH’’ for ‘‘hash rule holds’’ when for  $s \neq s', x^{d+2} + \sum_{i=1}^d s_i x^i = x'^{d+2} + \sum_{i=1}^d s'_i x'^i + \Delta_h$ . This means the message  $s', x', h'$  will be accepted by node 2. Hence the probability that the adversary wins is given by:

$$\begin{aligned} & \Pr(A \text{ wins}) \\ &= \sum_{\substack{x, \Delta_x \\ \Delta_h, s, s'}} \Pr(\text{HRH} | x, \Delta_h, \Delta_x, s, s') \Pr(\Delta_h, \Delta_x, s, s') \end{aligned} \quad (18)$$

Define  $Q$  as the term (18) with  $\Pr(x | \Delta_h, \Delta_x, s, s')$  replaced by  $\Pr(x)$ . Then we have

*Lemma 2:*

$$|\Pr(A \text{ wins}) - Q(A \text{ wins})| \leq \sqrt{(2 \ln 2) m \Pr(F = 0)} \quad (19)$$

*Proof:* The right hand side of (19) is bounded by:

$$\sum_{\substack{x, \Delta_x \\ \Delta_h, s, s'}} \left( |\Pr(x | \Delta_h, \Delta_x, s, s') \Pr(x) - \Pr(x) \Pr(\Delta_h, \Delta_x, s, s')| \right) \quad (20)$$

Then we use Pinsker's inequality:

$$I(A; B) \geq \frac{1}{2 \ln 2} D^2(p(A, B), p(A)p(B)) \quad (21)$$

where  $D(p(x), q(x)) = \sum_x |p(x) - q(x)|$ . Let  $A$  be  $x$ . Let  $B$  be  $\Delta_h, \Delta_x, s, s'$ . Then, applying Lemma 1 leads to Lemma 2. ■

From Theorem 1,  $Q(A \text{ wins})$  is bounded by  $\frac{d+1}{2^m}$ . Hence, using Lemma 2, we have

$$\Pr(A \text{ wins}) \leq \frac{d+1}{2^m} + \sqrt{(2 \ln 2) m \Pr(F = 0)} \quad (22)$$

To prove Theorem 3 and Theorem 4, it remains to apply (22) to the two models we considered, as will be done next.

### A. The Discrete Adder Model (Theorem 3)

For the adder model, (22) is given by

$$\Pr(A \text{ wins}) \leq \frac{d+1}{2^m} + \sqrt{m} \exp(-L\alpha') \quad (23)$$

where  $\alpha'$  is chosen to be within the range  $0 < \alpha' < \alpha/2$ . Now choose the length of information bits in the linear wire-tap code as the length of  $x$  and  $h$ . This means  $L/4 = 2m$ . The number of channel uses needed to transmit  $s$  is slightly larger than  $\frac{dm}{R_e}$ , where  $R_e$  is defined in Section IV. Hence the total number of channel uses needed to transmit  $\{s, x, h\}$  bits is given by  $8m + \frac{dm}{R_e}$ . This means the overall transmission rate converges to  $d/(8 + d/R_e)$ , which can be made arbitrarily close to  $R_e$  by making  $d$  sufficiently large.

For a given rate and hence fixed  $d$ , the total number of channel uses is proportional to  $m$ . From (23), we see that the probability that the adversary wins decreases exponentially fast with the total number of channel uses.

Finally, we check the secrecy constraint. The wire-tap code from [9] offers the following strong notion of secrecy:

$$I(s; Y_r(2)) < \exp(-\beta L') \quad (24)$$

where  $L' = md/R_e$  is the length of the wire-tap code. Hence:

$$I(s; Y_r(1), Y_r(2)) \leq I(s, x, h; Y_r(1), Y_r(2)) \quad (25)$$

$$\leq I(x, h; Y_r(1)) + I(s; Y_r(2)) \quad (26)$$

$$= m \exp(-8m\alpha) + \exp(-\beta md/R_e) \quad (27)$$

(26) is due to the fact that the channel and the signal transmitted by node 2 are both memoryless. (27) follows from Theorem 2 and (24). Hence for a given rate, the mutual information between the confidential message and the observation of the eavesdropper also decreases exponentially fast with the total number of channel uses.

### B. The Gaussian Model (Theorem 4)

For the Gaussian model, (22) is given by

$$\frac{d+1}{2^m} + \sqrt{(2 \ln 2) m (L \exp(-N\gamma) + \exp(-L\alpha))} \quad (28)$$

Let  $N_2$  denotes the number of channel uses during the second stage. Then the total number of channel uses is  $N_T = N_2 + LN$ . Let  $R'_e$  is the rate of the randomly generated wire-tap code per channel use. To match the rate with an AMD tuple of  $(d+2)m$  bits, We require  $R'_e(N_2) = dm$  and  $\frac{L}{4} = 2m$ , where 4 in the denominator comes from the fact we are using a rate 1/4 linear wiretap code. Let  $N_L$  be the dimension of the lattice code used in stage two. Let  $\gamma_L$  is the error exponent of the lattice code decoder. Then, with the wire-tap code from [9], it can be shown that  $R'_e$  can be chosen to be any value smaller than  $(1 - e^{-\gamma_L N_L})R_e - e^{-\gamma_L N_L} - 1/N_L$  [16]. Transmission below this rate will allow the decoding error probability at node 2 to decrease exponentially fast with respect to  $N_2/N_L = dm/(N_L R'_e)$ . The overall rate per channel use is given by

$$R_{\text{overall}} = \frac{dm}{N_2 + LN} = \frac{d/N}{\frac{d/N}{R'_e} + 8} \quad (29)$$

When  $d/N$  and  $N_L$  is sufficiently large,  $R_{\text{overall}}$  can be made arbitrarily close to  $R_e$ . If we fix  $d/N$  and let  $k = d/N$ , then the total number of channel uses is given by  $mN \left(8 + \frac{k}{R'_e}\right)$ . Let  $m = N$ . Then  $N_T \sim O(m^2)$ . It can be verified by checking (28) that  $\Pr(A \text{ wins})$  decreases exponentially fast with respect to  $m$ . It remains to check the secrecy constraints: The  $L'$  in (24) is given by  $N_2/N_L = km^2/(N_L R'_e)$ . Hence  $I(s; Y_r(1), Y_r(2)) = me^{-8m\alpha} + e^{-\frac{\beta km^2}{N_L R'_e}}$ . Hence if we choose  $N_L = m$ , then it decreases exponentially fast with respect to  $m$ . The decoding error probability of the wire-tap code used to transmit  $s$  decreases exponentially fast with respect to  $O(dm/(N_L R'_e)) = O(m)$ .

## VI. CONCLUSION

In this work, we have considered a source-destination pair that communicates via a relay node. The relay node is an eavesdropper and a Byzantine attacker. For the adder model, we showed that the probability that the adversary wins decreases exponentially fast with respect to the total number of channel uses  $N_T$ . For the Gaussian two-hop model, we showed that that the probability that the adversary wins decreases exponentially fast with respect to  $\sqrt{N_T}$ . For both models, the loss in rate due to Byzantine detection is arbitrarily small.

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