Efficient Codes for Adversarial Wiretap Channels

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Abstract—In [13] we proposed a \((\rho_r, \rho_w)\)-adversarial wiretap channel model (AWTP) in which the adversary can adaptively choose to see a fraction \(\rho_r\) of the codeword sent over the channel, and modify a fraction \(\rho_w\) of the codeword by adding arbitrary noise values to them. In this paper we give the first efficient construction of a capacity achieving code family that provides perfect secrecy for this channel.

I. INTRODUCTION

In Wyner’s wiretap model [15] channel noise in the channel is used as a resource for the system designer to provide (asymptotic) perfect secrecy against a computationally unbounded adversary without the need for a shared key. In this model, a sender and a receiver communicate over a noisy channel referred as the main channel, and their communication is eavesdropped by an adversary through a second noisy channel, referred to as the adversary channel. The goal is to provide (asymptotic) perfect reliable communication from sender to receiver with (asymptotic) perfect secrecy against the adversary. In this model adversary is passive and obstruction of its view by noise is probabilistic.

Recently a number of models [1], [4], [11] that include a stronger adversary that can modify communication have been introduced. These models primarily use arbitrarily varying channel approach and assume eavesdropper and jammer (who modifies communication) do not communicate. We introduced [13] an adversarial model for wiretap channel in which the adversary can adaptively choose a fraction of the communicated codeword to see and a fraction to modify. The modification of each component is by adding (algebraic) an arbitrary value (adversary’s choice) to the component. The adversary’s choice of observation and tampering components is unrestricted, as long as the total number of observation and tampering symbols are within specific limits. An Adversary Wiretap Channel (AWTP) is specific by two parameters \((\rho_r, \rho_w)\) and is denoted by \((\rho_r, \rho_w)\)-AWTP channel. An \((\epsilon, \delta)\)-AWTP code guarantees that the information leaked about the message (measured using statistical distance) and the probability of decoding failure are upper bounded by \(\epsilon\) and \(\delta\), respectively. The information rate of a code \(C\) is \(R(C) = \frac{\log |M|}{N \log |\Sigma|}\), where \(N\) is the length of the code and \(M\) is the message space. The code provides perfect secrecy if \(\epsilon = 0\).

We derived an upper bound on the rate of codes for \((\rho_r, \rho_w)\)-AWTP channels as \(R(C) \leq 1 - \rho_r - \rho_w + 2\epsilon \log |\Sigma| \frac{1}{\delta}\), and code family with perfect secrecy is \(R(C) \leq 1 - \rho_r - \rho_w\). An explicit and inefficient construction of AWTP code is also given in [13].

A. Our Result

We give an efficient construction of a code family \(C = \{C^N; N \in \mathbb{Z}\}\) in which every code \(C^N\) of length \(N\), provides perfect secrecy for a \((\rho_r, \rho_w)\)-AWTP channel. The construction uses three building blocks: an Algebraic Manipulate Detection Code (AMD code) [5], a Subspace Evasive Sets (SES) [7], and a Folded Reed-Solomon code (FRS code) [8]. AMD code detects algebraic manipulation assuming the adversary is oblivious and does not have access to the codeword. SES are subsets with the property that their intersection with any subset of certain dimension is bounded. FRS code is a special class of Reed-Solomon code that achieve list decoding capacity, and have efficient encoding and decoding. Encoding of a message uses the three building blocks in order: the message is encoded using AMD code, then using a SES and finally an FRS code. In decoding, first the FRS decoder outputs a list of possible codewords. This list for the decoding algorithm in [8], is a function of \(N\), the code length. Using the intersection algorithm of SES the list can be pruned to a shorter list which is independent of the code length. The final step is to use the AMD code to find the correct message. The decoder always outputs the correct message. We prove with appropriate choice of parameters, each code in the family is perfectly secure, satisfies the upper bound on rate for \((\rho_r, \rho_w)\)-AWTP channels with equality and so is capacity achieving, and finally the probability of decoding error reduces exponentially in \(N\).

B. Related Work

Wiretap channels have been an active area of research for a number of years with excellent progress on extending the model and strengthening security against passive adversary [2], [3], [6], [9]–[12]. More recently active adversary for these channels have been considered [1], [4], [11], [13]. The active adversary in [3], [11] is modeled using arbitrarily varying channels, and is assumed that there is no communication between the eavesdropper and the wiretapper. In [1] the wiretap II model is extended to active adversary. The adversary however is restricted to flip the codeword components that they have chosen to read. In [13] we proposed a model for adversarial channel called limited view adversarial channel (LVAC), which is the same as the adversarial channel considered here. The goal of communication however was reliability only. \((\rho_r, \rho_w)\)-AWTP channels have the same adversary power as LVAC channel, but the goal of communication is reliability and privacy both.

Paper organization: In section [1] we recall the model and capacity results for \((\rho_r, \rho_w)\)-AWTP channels. In section [11]...
II. MODEL AND DEFINITIONS

We consider the following scenario. Alice (Sender $S$) wants to send messages $m \in \mathcal{M}$ securely and reliably to Bob (Receiver $R$), over a communication channel that is partially controlled by Eve (Adversary). Let $[N] = \{1, \ldots, N\}$ and $S_r = \{i_1, \ldots, i_{p_r}, N\} \subseteq [N]$ and $S_w = \{j_1, \ldots, j_{p_w}, N\} \subseteq [N]$ denote two subsets of the $N$ coordinates. For a vector $x$, $\text{SUPP}(x)$ denotes the set of coordinates where $x_i$ is non-zero. Let $\Sigma$ denote the code alphabet, with an underlying group operation.

Definition 1: [13] A $(p_r, p_w)$-Adversarial Wiretap channel $((\rho_r, \rho_w))$-AWTP channel, is an adversarially corrupted communication channel between Alice and Bob such that it is (partially) controlled by an adversary Eve, with two capabilities: Reading and Writing. In Reading (or Eavesdropping), Eve selects a subset $S_r \subseteq [N]$ of size at most $\rho_r N$ and sees the components of the sent codeword $c$ on $S_r$. Eve’s view of the codeword is the set of all read components: $\text{View}_A(\text{AWTPenc}(m, r_s), r_A) = \{c_{i_1}, \ldots, c_{i_{p_r}, N}\}$. In Writing (or Jamming), Eve chooses a subset $S_w \subseteq [N]$ of size at most $\rho_w N$ and adds an error vector $e$ to $c$, where the addition is component-wise and over $\Sigma$. We require $\text{SUPP}(e) = S_w$. The corrupted components of $c$ are $\{y_{j_1}, \ldots, y_{j_{p_w}, N}\}$ and $y_{j_i} = c_{j_i} + e_{j_i}$. The error $e$ is generated according to the Eve’s best strategy to make Bob’s decoder fail.

The adversary is adaptive and selects components of $c$ for reading and writing, one by one and at each step using its knowledge of the codeword at that time.

Alice and Bob will use an Adversarial Wiretap channel to provide security and reliability for communication over Adversary wiretap channel.

Definition 2: [13] An $(\mathcal{M}, N, \Sigma, \epsilon, \delta)$-AWTP Code $((\epsilon, \delta))$-AWTP code for short) for a $(\rho_r, \rho_w)$-AWTP channel consists of a randomized encoding $\text{AWTPenc} : \mathcal{M} \times \mathcal{U} \rightarrow \mathcal{C}$, from the message space $\mathcal{M}$ to a code $\mathcal{C}$, and a deterministic decoding algorithm $\text{AWTPdec} : \Sigma^N \rightarrow \{\text{m} \uparrow \}$, such that $\text{AWTPdec}(\text{AWTPenc}(m, r_S)) = m$ for all $m \in \mathcal{M}$. The code guarantees secrecy and reliability as defined below.

i) Secrecy: For any two messages $m_1, m_2 \in \mathcal{M}$, we have

\[
\text{Adv}^{\text{Sec}}(\text{AWTPenc}, \text{View}_A) \triangleq \max_{m_0, m_1} \text{SD}(\text{View}_A(\text{AWTPenc}(m_1), r_A), \\
\text{View}_A(\text{AWTPenc}(m_2), r_A)) \leq \epsilon
\]

Here we assume the adversary uses the same random coins $r_A$ for the encoding of two messages.

ii) Reliability: For any message $m$ that is encoded to $e$ by the sender, and corrupted to $y = e + c$ by the $(\rho_r, \rho_w)$-AWTP channel, the probability that the receiver outputs the correct information $m$ is at least $1 - \delta$. Receiver will output $\perp$ with probability no more than $\delta$ and will never output an incorrect message. That is,

\[
P[\text{AWTPdec}(\text{AWTPenc}(m) + e) = \perp] \leq \delta
\]

An AWTP code is perfectly secure if $\epsilon = 0$.

Definition 3: For a fixed $\epsilon > 0$, an $\epsilon$-secure AWTP code family is a family $\mathcal{C} = \{\mathcal{C}_N\}_{N \in \mathbb{N}}$ of $(\epsilon, \delta_N)$-AWTP codes indexed by $N \in \mathbb{N}$, for a $(\rho_r, \rho_w)$-AWTP channel. When $\epsilon = 0$, the family is called a perfectly secure AWTP code family.

Definition 4: For a family $\mathcal{C}$ of $(\epsilon, \delta)$-AWTP codes the rate $R(\mathcal{C})$ is achievable if for any $\xi > 0$, there exists $N_0$ such that for any $N \geq N_0$, we have, $\frac{1}{N} \log|\mathcal{C}_N| \geq R(\mathcal{C}) - \xi$, and the probability of decoding error is $\delta \leq \xi$.

We use the achievable rate of a code family for an AWTP channel to define secrecy capacity of the channel.

Definition 5: The $\epsilon$-secrecy (perfect secrecy) capacity of a $(\rho_r, \rho_w)$-AWTP channel denoted by $C^\epsilon, \epsilon_0^0$ (perfect secrecy) capacity of an AWTP channel, is the largest achievable rate of all $(\epsilon, \delta)$-AWTP codes $((0, \delta))$-AWTP code families $\mathcal{C}$ for the channel.

The following upper bounds are derived in [13].

Lemma 1: [13] The $\epsilon$-secrecy capacity of a $(\rho_r, \rho_w)$-AWTP channel satisfies the upper bound,

\[
C^\epsilon, \epsilon_0^0 \leq 1 - \rho_r - \rho_w + 2\epsilon \rho_w N \log|\Sigma| (1 + \frac{1}{\epsilon})
\]

The upper bound for the perfect secrecy capacity of a $(\rho_r, \rho_w)$-AWTP channel is, $C^{\epsilon_0^0} \leq 1 - \rho_r - \rho_w$.

III. AN EFFICIENT CAPACITY ACHIEVING AWTP-CODE

The general approach to the construction was outlined in Section I. A. Below we recall the definition of the building blocks, and give our instantiations, and construction of the code.

1) Algebraic Manipulation Detection Code (AMD code): Consider a storage device $\Sigma(\mathcal{G})$ that holds an element $x$ from a group $\mathcal{G}$. The storage $\Sigma(\mathcal{G})$ is private but can be manipulated by the adversary by adding $\Delta \in \mathcal{G}$. AMD code allows the manipulation to be detected.

Definition 6 (AMD-code [3]): An $(\mathcal{X}, \mathcal{G}, \delta)$-Algebraic Manipulation Detection code $((\mathcal{X}, \mathcal{G}, \delta))$-AMD code consists of two algorithms (AMDenc, AMDdec). Encoding given by, $\text{AMDenc} : \mathcal{X} \rightarrow \mathcal{G}$, is probabilistic and maps an element of a set $\mathcal{X}$ to an element of an additive group $\mathcal{G}$. Decoding, $\text{AMDdec} : \mathcal{G} \rightarrow \mathcal{X} \cup \{\perp\}$, is deterministic and we have $\text{AMDdec}(\text{AMDenc}(x)) = x$, for any $x \in \mathcal{X}$. Security of AMD codes is defined by requiring,

\[
P[\text{AMDdec}(\text{AMDenc}(x) + \Delta) \in \{x, \perp\}] \leq \delta, \quad (1)
\]

for all $x \in \mathcal{X}, \Delta \in \mathcal{G}$.

An AMD code is systematic if the encoding has the form $\text{AMDenc} : \mathcal{X} \rightarrow \mathcal{X} \times \mathcal{G}_1 \times \mathcal{G}_2$, $x \rightarrow (x, r, t = f(x, r))$ for some function $f$ and $r \in \mathcal{G}_1$. The decoding function $\text{AMDdec}(x, r, t)$ is $x$ if and only if $t = f(x, r)$ and $\perp$ otherwise.

We use a systematic AMD-code that is based on the extension of the construction in [3] to extension fields. Let $\phi$ be a bijection between vectors $v$ of length $N$ over $\mathbb{F}_q$, and elements in $\mathbb{F}_{q^N}$, and let $\ell$ be an integer such that $\ell + 2$ is not divisible
by $q$. Define the function $AMDenc : \mathbb{F}_q^t \rightarrow \mathbb{F}_q^r \times \mathbb{F}_q^N \times \mathbb{F}_q^N$ by

$$f(x, r) = \phi^{-1} \left( \phi(r)^{t+2} + \sum_{i=1}^t \phi(x_i) \phi(r)^i \right) \mod q^N$$

Lemma 2: For the AMD-code above, given a codeword $(x, r, t)$, the success chance of an adversary that has no information about $(x, r, t)$, in constructing a new codeword $(x', r', t') = (x' = x + \Delta x, r' = r + \Delta r, t' = t + \Delta t)$, that passes the verification $t' = f(x', r')$ is at most \(\frac{\epsilon}{\ell}\).

2) Subspace Evasive Sets: We briefly introduce subspace evasive sets. More details can be found in Appendix A.

Definition 7 (Subspace Evasive Sets [27, 38]): Let $S \subset \mathbb{F}_q^n$. We say $S$ is $(v, \ell \cdot \varepsilon)$-subspace evasive if for all $v$-dimensional affine subspaces $H \subset \mathbb{F}_q^n$, we have $|S \cap H| \leq \ell \cdot \varepsilon$. Dvir et al. [27] show that there is an efficient construction for subspace evasive sets $S \subset \mathbb{F}_q^n$, and an efficient intersection algorithm to compute $S \cap H$ for any $v$-dimensional subspace $H \subset \mathbb{F}_q^n$.

Lemma 3: Let $v, n_1 \in \mathbb{N}, w = v^2, n = \frac{w}{\ell N}$, and $\mathbb{F}_q$ be a finite field. Then there is a $(v, w \cdot C \log \log v)$-subspace evasive set $S \subset \mathbb{F}_q^n$. For any vector $v \in \mathbb{F}_q^n$, there is a bijection which maps $v$ into an elements of the subspace evasive set. That is

$$SE : v \rightarrow v' \in S$$

Lemma 4: Let $S \subset \mathbb{F}_q^n$ be the $(v, \ell \cdot \varepsilon)$-subspace evasive set. There exists an algorithm that, given a basis for any $H$, output $S \cap H$ in $O(v \cdot C \log \log v)$ time.

3) Folded Reed-Solomon Code (FRS code): A error correcting code $C$ is a subspace of $\mathbb{F}_q^N$. The rate of the code is $\log_2 \frac{|C|}{N}$. A code $C$ of length $N$ and rate $R$ is $(\rho, \ell \cdot \varepsilon)$-list decodable if the number of codewords within distance $\rho N$ from any received word is at most $\ell \cdot \varepsilon$. List decodable codes can potentially correct up to $1 - R$ fraction of errors, which is twice that of unique decoding. This is however at the cost of outputting a list of possible sent codewords (messages). Construction of good code with efficient list decoding algorithms is an important research question. An explicit construction of a list decodable code that achieves the list decoding capacity $\rho = 1 - R - \varepsilon$ is given by Guruswami et al. [38]. The code is called Folded Reed-Solomon codes (FRS codes), defined by Guruswami et al. [38], gives an explicit construction for list decodable codes that achieve the list decoding capacity $\rho = 1 - R - \varepsilon$. The code has polynomial time encoding and decoding algorithms.

Definition 8: [38] A $u$-Folded Reed-Solomon code is an error correcting code with block length $N$ over $\mathbb{F}_q$ and $q > N u$. The message of an FRS code is written in the form of a polynomial $f(x)$ of degree $k$ over $\mathbb{F}_q$. The FRS codeword corresponding to the message is a vector over $\mathbb{F}_q^n$, where each component is a $u$-tuple $(f(\gamma^j), f(\gamma^{j+u}), \ldots, f(\gamma^{j+u^{t-1}}))$, $0 \leq j < N$, where $\gamma$ is a generator of $\mathbb{F}_q^*$, the multiplicative group of $\mathbb{F}_q$. A codeword of a $u$-folded Reed-Solomon code of length $N$ is in one-to-one correspondence with a codeword $c$ of a Reed-Solomon code of length $uN$, and is obtained by grouping together $u$ consecutive components of $c$. We use $FRSenc$ to denote the encoding algorithm of the FRS code. $u$ is called the folding parameter of the FRS code.

We will use the linear algebraic FRS decoding algorithm of these codes [38] (Appendix B-A). The following Lemma gives the decoding capability of linear algebraic FRS code.

Lemma 5: [38] For a Folded Reed-Solomon code of block length $N$ and rate $R = \frac{k}{q^N}$, the following holds for all integers $1 \leq v \leq u$. Given a received word $y \in \mathbb{F}_q^N$ agreeing with $c$ in at least a fraction,

$$N - \rho N > N\left(\frac{1}{v+1} + \frac{v}{v+1} \frac{uR}{u - v + 1}\right)$$

one can compute a matrix $M \in \mathbb{F}_q^{k \times (v - 1)}$ and a vector $z \in \mathbb{F}_q^v$ such that the message polynomials $f \in \mathbb{F}_q[X]$ in the decoded list are contained in the affine space $Mb + z$ for $b \in \mathbb{F}_q^v$ in $O((Nu log q)^2)$ time.

A. An Explicit Capacity Achieving $(0, \delta)$-AWTP Code Family

Let $M$ denote the message space, $N$ denote the code length and the encoding and decoding algorithms be, $AWTPenc_N$ and $AWTDec_N$, respectively. The message, also referred to as the information block of the AWTP code, is $m = \{m_1, \ldots, m_{uRN}\} \in M$ where $m_i \in \mathbb{F}_q$. Let $S$ be a $(v, w \cdot C \log \log v)$-subspace evasive set in $\mathbb{F}_q^n$. Let $u$ and $v$ denote the folding and the interpolation parameters of the FRS code, respectively. Let $s$ be a prime number larger than $Nu$, $\gamma$ be a primitive element of $\mathbb{F}_q$, $\ell = [uR]$, $w = v^2$, $b = \left[\frac{Nu + 2N}{\ell}\right]$, $n_1 = (w - v)b$, $n = nb$, $SE : \mathbb{F}_q^n \rightarrow S$ be the bijection of subspace evasive set.

The construction of encoder and decoder for $C_N$ is given in Figure III-A.

Figure III-A

Encoding: For a code rate $R$, the sender $S$ does the following.

1) Start with the information block $m$ of length $uRN$. Append sufficient zeros $N(\ell - uR)$ to construct a vector $x$ of length $N \ell$; that is, $x = \{m_0, \ldots, 0\}$.

2) Generate a random vector $r$ with length $N$ over $\mathbb{F}_q$. Use the AMD construction in section III-C to construct the AMD codeword $(x, r, t)$. That is, AMDenc$(x) = (x, r, t)$. The length of AMD code is $\ell N + 2N$.

3) Extend the AMD codeword to length $n_1$ by appending zeros. Encode AMD code into an element $s$ of the subspace evasive set $S$. The length of $s$ is $n$. That is

$$s = SE(x, r, t)[0, \ldots, 0]$$

4) Append a random vector $a = \{a_1 \cdots a_{u\rho}, N\} \in \mathbb{F}_q^{u\rho, N}$ to $s$ to form a vector that will be the message of the FRS code, and interpret that as
coefficients of the polynomial $f(x)$ over $\mathbb{F}_q$. That is $\{f_0, \cdots, f_{k-1}\} = \{s||a\}$. We have $k = \text{deg}(f) + 1 = u\rho_r N + n$.

5) Use $\text{FRSenc}$ to construct the FRS codeword $c = \text{FRSenc}(f(X)) = \{c_1, \cdots, c_N\}$, and $c_i = \{f(\gamma_i(u-1)), \cdots, f(\gamma_i u^{-1})\} \in \mathbb{F}_q^u$, $i = 1, \cdots, N$.

**Decoding:** The receiver $R$ does the following:

1) Let $y = c + e$, and $w_H(e) \leq \rho_d N$. The $i$-th component of $y$ is $y_i = \{y_{i,1}, \cdots, y_{i,u}\}$ for $i = 1, \cdots, N$.

2) Use the FRS decoding algorithm $\text{FRSdec}(y)$ to output a matrix $M \in \mathbb{F}_q^{k \times u}$ and a vector $z \in \mathbb{F}_q^u$. Such that the codewords in the output list are $L_{\text{FRS}} = Mb + z$, $M$ has $k$ rows each giving a component of the output vector as a linear combination of $\{b_1, \cdots, b_u\}$. Let $H$ denote the space which is generated by the first $n$ equations. That is $H = M_{n \times k}b + z_n$, $b \in \mathbb{F}_q^u$.

where $M_{n \times k}$ is the first $n$ rows of the submatrix of $M_{n \times u}$ and $z_n$ is the first $n$ elements of $z$.

3) The decoder calculates the intersection $S \cap H$ and outputs a list $L$ with size at most $(rCv^{\log \log v})^{uN}$. Each $s_i \in L$ corresponds to an AMD codeword $\{x_i, r_i, t_i\}$.

4) For each AMD codeword $\{x_i, r_i, t_i\}$, the decoder verifies $t_i = f(x_i)$. If there is a unique valid AMD codeword, the decoder outputs the first $uRN$ components of $x$ as the correct message $m$. Otherwise, outputs $\perp$.

It is easy to see that $s$ together with the randomness $a$ uniquely determines $c^{[r]}$. This gives

$$P(C^{[r]} = c^{[r]} | \{S, A\} = \{s, a\}) = 1. \quad (2)$$

Conversely, for given values of $s$ and $\{c_{1,1}, \cdots, c_{1,n}\}$, and noting that the coefficient matrix is Vandermonde, there exists a unique solution for the $u \rho_r N$ unknown components of $a = \{a_1, \cdots, a_{u \rho_r N}\} \in \mathbb{F}_q^{u \rho_r N}$. That is

$$P(A = a | \{S, C^{[r]}\} = \{s, c^{[r]}\}) = 1 \quad (3)$$

Since $a$ is chosen uniformly and independent of $s$, we have

$$P(A = a | S = s) = \frac{1}{q^{u \rho_r N}} \quad (4)$$

From (2), (3), and (4) we have,

$$P(C^{[r]} = c^{[r]}, A = a | S = s) = \frac{1}{q^{u \rho_r N}} \left| P(C^{[r]} = c^{[r]} | S = s) \right| P(A = a | S = s)$$

which implies for any $s$, $\text{SD}(\text{View}_{\mathcal{A}} | s_1, \text{View}_{\mathcal{A}} | s_2)$ = 0

**Lemma 7 (Reliability):** The failure probability of AWTPdec$_N$ is bounded by $\delta_N \leq \frac{e^{Cv^{\log \log v}}}{q^{uN}}$.

**Proof:** The FRS decoder outputs a list of elements of the subspace evasive set $s_i \in L$ with list size at most $\ell_{SE} \leq uCv^{\log \log v}$. Each element corresponds to a unique AMD codeword $\{x_i, r_i, t_i\} = \text{SE}^{-1}(s_i)$.

We first show that the correct message $m$ will always be output by the receiver. Denote the AMD codeword corresponding to the message $m$ as $[x, r, t] = \text{AMDenc}(m)[0, \cdots, 0]$. The list decoding algorithm outputs codewords that are at distance at most $\rho_d N$ of the received word and so include the original codeword. The bijection function $\text{SE}$ encodes the AMD codeword into an element of the subspace evasive set $s \in S$ that belongs to the decoded list $s \in H$ that passes AMD verification. That is,

$$\text{SE}(x, r, t)[0, \cdots, 0] \in L = S \cap H \quad \text{and} \quad t = f(x, r)$$

Second, we show that the probability that any other codeword in the list is a valid AMD codeword is small. That is we will show that,

$$P(\{x', r', t'\} = \text{SE}^{-1}(s') \cap S \subseteq L \cap t' = f(x', r')) \leq \frac{\ell}{q^N}$$

From Lemma 6 the adversary has no information about the encoded subspace evasive sets element $s$ and the AMD.
Finally, we show the unique correct message output by receiver with probability at least $1 - \frac{v!^{(C+2)\cdot v \cdot \log \log v}}{q^N}$. The list size is at most $v!^{(C+2)\cdot v \cdot \log \log v}$ and $t \leq v^2$. So the probability that any $\{x', r', t\} \neq \{x, r, t\}$ in decoding list pass the verification $t' = f(x', r')$, is no more than $\frac{v!^{(C+2)\cdot v \cdot \log \log v}}{q^N}$. That is

\[
P(\bigcup_{s' \in L} \{x', r', t'\} = SE^{-1}(s') \land t' = f(x', r')) \\
\leq \sum_{s' \in L} P(\{x', r', t'\} = SE^{-1}(s') \land t' = f(x', r')) \\
\leq \sum_{s' \in L} P(t' = f(x', r')) \leq \frac{\ell |L|}{q^N} \leq \frac{v!^{(C+2)\cdot v \cdot \log \log v}}{q^N}
\]

We first find the information rate of the code $C_N$, and then find the achievable rate of the code family $C$. 

**Lemma 8 (Rate of $C_N$):** The AWTP code $C_N$ described above provides reliability for a $(\rho_r, \rho_w)$-AWTP channel if the following holds:

\[
\rho_w < \frac{v}{v + 1} - \frac{w(uR + 3) + u\rho_r}{u - v + 1}.
\]

Proof is in Appendix C

**Lemma 9 (Achievable Rate of $C$):** The information rate of the $(0, \delta)$-AWTP code family $C = \{C_N\}_{N \in \mathbb{N}}$ for a $(\rho_r, \rho_w)$-AWTP channel is $R(C) = 1 - \rho_r - \rho_w$.

**Proof:** For a given small $\frac{1}{2} > \xi > 0$, let code parameters be chosen as, $\xi_1 = \frac{\xi}{\xi_2}$, $v = \frac{1}{\xi_1}$ and $u = \frac{1}{\xi_2}$. Finally let, $N_0 > (1/\xi)^{C/\xi \log \log 1/\xi}$ where $C > 0$ is constant. From

\[
\left(1 - R - \rho_r - 12\xi_1\right) \leq \left(\frac{1}{\xi_1 + 1} - \frac{1}{\xi_1 + 1} - 1 - \xi_1\right) + \rho_r
\]

the decoding condition of AWTP code is satisfied if,

\[
\rho_w < 1 - R - \rho_r - 12\xi_1.
\]

We choose $R = 1 - \rho_r - \rho_w - 12\xi_1$, the decoding condition of AWTP code will be satisfied. Now since $\xi = 13\xi_1$, for any $N > N_0$, the rate of the AWTP code $C_N$ is

\[
\frac{1}{N} \log |\Sigma| \cdot |M_N| = R = 1 - \rho_r - \rho_w - 12\xi_1 \\
> 1 - \rho_r - \rho_w - \xi = R(C) - \xi
\]

and the probability of decoding error,

\[
\delta_N \leq (1/\xi)^{C/\xi \log \log 1/\xi} q^{-N} \leq Nq^{-N} \leq \xi
\]

So the information rate of AWTP code family $C$ is $R(C) = 1 - \rho_r - \rho_w$.

The computational time for encoding is $O((N \log q)^2)$. The decoding of FRS code and intersection algorithm of the subspace evasive set is $O((1/\xi)^{C/\xi \log \log 1/\xi})$. The AMD verification is $O((1/\xi)^{C/\xi \log \log 1/\xi} (N \log q)^2)$. So the total computational time of decoding is $O((N \log q)^2)$.

**Theorem 1:** For any small $\xi > 0$, there is $(0, \delta)$-AWTP code $C_N$ of length $N$ over $(\rho_r, \rho_w)$-AWTP channel such that the information rate is $R(C_N) = 1 - \rho_r - \rho_w - \xi$, the size of alphabet is $|\Sigma| = O(q^{1/\xi^2})$ and decoding error $\delta < q^{-\Omega(N)}$. The computational time is $O((N \log q)^2)$. The AWTP code family $C = \{C_N\}_{N \in \mathbb{N}}$ achieves secrecy capacity $R(C) = 1 - \rho_r - \rho_w$ for $(\rho_r, \rho_w)$-AWTP channel.

**IV. CONCLUDING REMARKS**

$(\rho_r, \rho_w)$-AWTP extends Wyner wiretap models to include active corruption at physical layer of communication channel. Although corruption in our general model is additive, for $S_w \subset S_r$, it is equivalent to arbitrary replacement of code components. We proposed an efficient construction for a capacity achieving code family for $(\rho_r, \rho_w)$-AWTP channels. The alphabet size for the code is $q^{\frac{1}{\xi}}$ where for $\delta < \xi$, $u = \Omega(1/\xi)$. That is for small failure probability, larger size alphabet must be used. Constructing capacity achieving codes over small (fixed) size alphabets remains an open problem.

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Appendix A
Subspace Evasive Sets

Recently, Guruswami et al. [8] showed that the subspace evasive sets can be used to reduce the list size of list decoding algorithm. Dvir et al. [11] gives a explicit and efficient construction of subspace evasive sets. We briefly introduce Dvir et al. [11]’s construction of subspace evasive sets. In detail, we give the definition of subspace evasive set, the construction, the encoding function, and the bound of the size of intersection between subspace evasive sets and any v-dimensional space.

Definition 9: [8, 11] Let $S \subset \mathbb{F}^n$. We say $S$ is $(v, \ell SE)$-subspace evasive sets if for all $v$-dimensional affine subspaces $\mathcal{H} \subset \mathbb{F}^n$, there is $|S \cap \mathcal{H}| \leq \ell SE$.

A. Construction of Subspace Evasive Set

Let $\mathbb{F}$ be a field and $\mathbb{F}$ be its algebraic closure. A variety in $\mathbb{F}^w$ is the set of common zeros of one or more polynomials. Given $v$ polynomials $f_1, \ldots, f_v \in \mathbb{F}[x_1, \ldots, x_w]$, we denote the variety as

$$V(f_1, \ldots, f_v) = \{x \in \mathbb{F}^w : f_1(x) = \cdots = f_v(x) = 0\}$$

where $x = (x_1, \ldots, x_w)$.

For a polynomials $f_1, \ldots, f_v \in \mathbb{F}[x_1, \ldots, x_w]$, we define the common solutions in $\mathbb{F}^w$ as

$$V_\mathbb{F}(f_1, \ldots, f_v) = V(f_1, \ldots, f_v) \cap \mathbb{F}^w = \{x \in \mathbb{F}^w : f_1(x) = \cdots = f_v(x) = 0\}$$

We say that a $v \times w$ matrix is strongly-regular if all its $r \times r$ minors are regular for all $1 \leq r \leq v$. For instance, if $\mathbb{F}$ is a field with at least $w$ distinct nonzero elements $\gamma_1, \ldots, \gamma_w$, then $A_{i,j} = \gamma_j^i$ is strongly-regular.

Lemma 10: (Theorem 3.2 [2]) Let $v \geq 1, \varepsilon > 0$ and $\mathbb{F}$ be a finite field. Let $w = v/\varepsilon$ and $w$ divides $n$. Let $A$ be a $v \times w$ matrix with coefficients in $\mathbb{F}$ which is strongly-regular. Let $d_1 \times \cdots \times d_w$ be integers. For $i \in [v]$ let

$$f_i(x_1, \ldots, x_w) = \sum_{j=1}^w A_{i,j}x_{d_j}$$

and define the subspace evasive sets $S \subset \mathbb{F}^n$ to be $(n/w)$-times cartesian product of $V_\mathbb{F}(f_1, \ldots, f_v) \subset \mathbb{F}^w$. That is

$$S = V_\mathbb{F}(f_1, \ldots, f_v) \times \cdots \times V_\mathbb{F}(f_1, \ldots, f_v) = \{x \in \mathbb{F}^n : f_i(x_{tw+1}, \ldots, x_{tw+w}) = 0, \forall 0 \leq t < n/w, 1 \leq i \leq v\}$$

Then $S$ is $(v, (d_1, \ldots, d_v))$ subspace evasive sets.

Moreover, if at least $v$ of the degrees $d_1, \ldots, d_w$ are co-prime to $|\mathbb{F}| - 1$, then $|S| = \left|\mathbb{F}\right|^{(1-\varepsilon)n}$.

The size of list is bounded by $d_1$ and $v$. If we can bound $d_1$ by $v$, the list size can be only bounded by the $v$-dimensional subspace $\mathcal{H}$.

Lemma 11: (Claim 4.3 [2]) There exists a constant $C > 0$ such that the following holds: There is a deterministic algorithm that, given integer inputs $v, N$ so that in $\text{Poly}(N)$ time there is prime $q$ and $v$ integers $vC\log \log v > d_1 > d_2 > \cdots > d_w > 1$ such that:

1) For all $i \in [v]$,

$$\text{gcd}(q - 1, d_i) = 1$$

2) $N < q \leq vC\log \log v$.

Because we only need to choose $w$ integer $d_1 > \cdots > d_w$ and $v$ of the integers are co-prime to $q$, the bound of $d_1$ is $d_1 \leq \max(v, vC\log \log v)$.

B. Encoding Vector as Elements in $S$

We show the encoding map $SE : v \rightarrow s$. Assuming there is a vector $v$ of length $n_1$ and $(w-v)\mid n_1$. First we divide the vector into $\frac{n_1}{w}$ blocks. Then for each block $v_i$ for $i = 1, \cdots, \frac{n_1}{w}$, we encode into a block $s_i$ using bijection $\varphi$. Then we concatenate each block $s_i$ for $i = 1, \cdots, \frac{n_1}{w}$ and generate $s$ in $S$. We give the function $\varphi$ in the following.

Lemma 12: (Claim 4.1) Assume that at least $v$ of the degree $d_1, \ldots, d_v$ are co-prime to $|\mathbb{F}| - 1$ and then there is an easy to compute bijection $\varphi : \mathbb{F}^{w-v} \rightarrow V_{\mathbb{F}} \subset \mathbb{F}^w$. Moreover, there are $w-v$ coordinates in the output of $\varphi$ that can be obtained from the identity mapping $Id : \mathbb{F}^{w-v} \rightarrow \mathbb{F}^{w-v}$.

Let $d_1, \ldots, d_v$ be the degree among $d_1, \ldots, d_w$ co-prime to $|\mathbb{F}| - 1$ and let $J = \{j_1, \ldots, j_v\}$ and $d_{j_i} = q_i$. On the positions $[w] \setminus J$, the map $\varphi$ takes the elements from $\mathbb{F}^{w-v} \to \mathbb{F}^w \setminus J$. For the elements on $J$, there is

$$\sum_{j \in J} A_{i,j}x_{d_j} = -\sum_{j \in J} A_{i,j}x_{d_j}$$

Let $A'$ be the $v \times v$ minor of $A$ given by restricting $A$ to columns in $J$ and $b_i = -\sum_{j \in J} A_{i,j}x_{d_j}$. Then

$$A'y = b$$

and for each $y$, there is unique solution of $x_{d_{j_i}} = y_i \mod q$ because $d_{j_i}$ is co-prime to $q - 1$.

C. Computing Intersection

We show how to compute the intersection $S \cap \mathcal{H}$ given $(v, \ell SE)$ subspace evasive sets $S$ and $v$-dimensional subspace $\mathcal{H}$. The subspace evasive sets $S$ will filter out the elements in $\mathcal{H}$ and output a set of elements $S \cap \mathcal{H}$ with size no more than $\ell SE$.

Lemma 13: (Claim 4.2 [2]) Let $S \subset \mathbb{F}^n$ be the $(v, \ell SE)$-subspace evasive sets. There exists an algorithm that, given a basis of $\mathcal{H}$, output $S \cap \mathcal{H}$ in $\text{Poly}(d_1 \cdots d_v)$ time.

Because $\mathcal{H}$ is $v$-dimensional subspace and $\mathcal{H} \subset \mathbb{F}^n$, there exists a set of affine maps $\ell \{\ell_1, \ldots, \ell_h\}$ such that for any elements $x = (x_1, \ldots, x_m) \in \mathcal{H}$, there is $x_i = \ell_1(s_1, \ldots, s_v)$. We show the result by induction of the number of blocks $i = 1, \ldots, n/w$. If $i = 1$, let $\mathcal{H}_1 := \{(x_1, \ldots, x_w) : (x_1, \ldots, x_w) \in \mathcal{H}\}$, the dimension of $\mathcal{H}_1$ is $r_1 \leq v$ and $\mathcal{H}_{x_1, \ldots, x_w} = \{(x_1, \ldots, x_w) \in \mathcal{H} : (x_1, \ldots, x_w)\}$ such that $\mathcal{H} = \cup_{(x_1, \ldots, x_w) \in \mathcal{H}_{x_1, \ldots, x_w}}$. And the dimension of $\mathcal{H}_{x_1, \ldots, x_w}$ is $v - r_1$. There is

$$V_\mathbb{F}(f_1, \ldots, f_v) \cap \mathcal{H}$$

$$= \{(x_1, \ldots, x_w) : (\ell_1(s_1, \ldots, s_v), \ldots, \ell_w(s_1, \ldots, s_v)) = 0, \ldots, f_v(\ell_1(s_1, \ldots, s_v), \ldots, \ell_w(s_1, \ldots, s_v)) = 0\}$$
We can solve the $v$ equations to get $(s_1, \cdots, s_v)$ and then obtain $(x_1, \cdots, x_w)$. Since $\mathcal{H}_1 \subset \mathbb{F}^w$,
\[
V_{\mathcal{F}}(f_1, \cdots, f_v) \cap \mathcal{H}_1 = V(f_1, \cdots, f_v) \cap \mathcal{H}_1
\]
By Bezout’s theorem, there is $|V(f_1, \cdots, f_v)\cap \mathcal{H}_1| \leq (d_1)^r_1$. So there are at most $(d_1)^r_1$ solutions for $(x_1, \cdots, x_w) \in \mathcal{H}_1$.

The computational time of solving the equation system follows from powerful algorithms that can solve a system of polynomial equations (over finite fields) in time polynomial in the size of the output, provided that the number of solutions is finite in the algebraic closure (i.e. the zero-dimensional case). So for $i = 1$, the computational time is at most $\text{Poly}((d_1)^r_1)$ and there are $(d_1)^r_1$ solutions for $(x_1, \cdots, x_w)$.

For every fixed of the first $w$ coordinates, we reduce the dimension of $\mathcal{H}$ by $r_1$ and obtained a new subspace $\mathcal{H}_2$ on the remaining coordinates. Continuing in the same fashion with $\mathcal{H}_2$ on the second block we can compute all the solutions in time $\text{Poly}((d_1)^{r_1}) \cdot \text{Poly}((d_1)^{r_2}) \cdots \text{Poly}((d_1)^{r_{n/w}})$, where $r_1 + r_2 + \cdots + r_{n/w} = v$. So the total running time is $\text{Poly}((d_1)^v)$.

**APPENDIX B**

**LIST DECODABLE CODE**

A. Decoding algorithm of FRS code

Linear algebraic list decoding [8] has two main steps: interpolation and message finding as outlined below.

- Find a polynomial, $Q(X, Y_1, \cdots, Y_v) = A_0(X) + A_1(X)Y_1 + \cdots + A_v(X)Y_v$, over $\mathbb{F}_q$ such that $\deg(A_i(X)) \leq D$, for $i = 1 \cdots v$, and $\deg(A_0(X)) \leq D + k - 1$, satisfying $Q(a_1, X, y_2, \cdots, y_v) = 0$ for $1 \leq i \leq n_0$, where $n_0 = (u - v + 1)N$.

- Find all polynomials $f(X) \in \mathbb{F}_q[X]$ of degree at most $k - 1$, with coefficients $f_0, f_1, \cdots, f_{k-1}$, that satisfy, $A_0(X) + A_1(X)f(X) + A_2(X)f(\gamma X) + \cdots + A_v(X)f(\gamma^{v-1}X) = 0$, by solving linear equation system.

The two above requirements are satisfied if $f \in \mathbb{F}_q[X]$ is a polynomial of degree at most $k - 1$ whose FRS encoding agrees with the received word $y$ in at least $t$ components:

$$t > N\left(\frac{1}{v + 1} + \frac{v}{v + 1} \frac{uR}{u - v + 1}\right)$$

This means we need to find all polynomials $f(X) \in \mathbb{F}_q[X]$ of degree at most $k - 1$, with coefficients $f_0, f_1, \cdots, f_{k-1}$, that satisfy,

$$A_0(X) + A_1(X)f(X) + A_2(X)f(\gamma X) + \cdots + A_v(X)f(\gamma^{v-1}X) = 0$$

Let us denote $A_i(X) = \sum_{j=0}^{D+k-1} a_{i,j} X^j$ for $0 \leq i \leq v$. $(a_{i,j} = 0$ when $i \geq 1$ and $j \geq D)$. Define the polynomials,

$$B_0(X) = a_{1,0} + a_{2,0}X + a_{3,0}X^2 + \cdots + a_{v,0}X^{v-1},$$

$$B_{k-1}(X) = a_{1,k-1} + a_{2,k-1}X + a_{3,k-1}X^2 + \cdots + a_{v,k-1}X^{v-1}$$

So for $0 \leq i \leq k - 1$, we have

$$B_i(X)f(X) = 0$$

We examine the condition that the coefficients of $X^i$ of the polynomial $Q(X) = A_0(X) + A_1(X)f(X) + A_2(X)f(\gamma X) + \cdots + A_v(X)f(\gamma^{v-1}X) = 0$ equals $0$, for $i = 0, \cdots, k - 1$. This is equivalent to the following system of linear equations for $f_0 \cdots f_{k-1}$.

$$
\begin{bmatrix}
B_0(\gamma^0) & 0 & 0 & \cdots & 0 \\
B_1(\gamma^0) & B_0(\gamma^1) & 0 & \cdots & 0 \\
B_2(\gamma^0) & B_1(\gamma^1) & B_0(\gamma^2) & \cdots & 0 \\
\vdots & \vdots & \vdots & \ddots & \vdots \\
B_{k-1}(\gamma^0) & B_{k-2}(\gamma^1) & B_{k-3}(\gamma^2) & \cdots & B_0(\gamma^{k-1})
\end{bmatrix} 
\begin{bmatrix}
f_0 \\
f_1 \\
f_2 \\
\vdots \\
f_{k-1}
\end{bmatrix} = 
\begin{bmatrix}
-a_{0,0} \\
-a_{0,1} \\
-a_{0,2} \\
\vdots \\
-a_{0,k-1}
\end{bmatrix}
$$

The rank of the matrix of (Eqs. 8) is at least $k - v + 1$ because there are at most $v - 1$ solutions of equation $B_0(X) = 0$ so at most $v - 1$ of $\gamma^i$ that makes $B_0(\gamma^i) = 0$. The dimension of solution space is at most $v - 1$ because the rank of matrix of (Eqs. 8) is at least $k - v + 1$. So there are at most $q^{v-1}$ solutions to (Eqs. 8) and this determines the size of the list which is equal to $q^{v-1}$.

**APPENDIX C**

**PROOF OF LEMMA**

**Proof:** FRS decoding algorithm FRSdec requires,

$$N - \rho_u N > N\left(\frac{1}{v + 1} + \frac{v}{v + 1} \frac{uR_{\text{FRS}}}{u - v + 1}\right)$$

The dimension of the FRS code is bounded by,

$$k = uR_{\text{FRS}} N = w\left[\ell N + 2N\left(\frac{w}{w - uR}\right) + uR_{\text{FRS}} N\right]$$

The (10) holds because $\ell \leq uR + 1$. So the decoding condition for FRS code holds if,

$$N - \rho_u N > N\left(\frac{1}{v + 1} + \frac{v}{v + 1} \frac{w}{w - uR + 3N} + uR_{\text{FRS}} N\right)$$

From $w = v^2$, it is equivalent to,

$$\rho_u < \frac{v}{v + 1} - \frac{v}{v + 1} \frac{w}{w - uR + 3N} + uR_{\text{FRS}} N$$

[8]