Shannon's Secrecy System With Informed Receivers and its Application to Systematic Coding for Wiretapped Channels

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Abstract

Shannon's secrecy system is studied in a setting, where both the legitimate decoder and the wiretapper have access to side information sequences correlated to the source, but the wiretapper receives both the coded information and the side information via channels that are more noisy than the respective channels of the legitmate decoder, which in turn, also shares a secret key with the encoder. A single–letter characterization is provided for the achievable region in the space of five figures of merit: the equivocation at the wiretapper, the key rate, the distortion of the source reconstruction at the legitimate receiver, the bandwidth expansion factor of the coded channels, and the average transmission cost (generalized power). Beyond the fact that this is an extension of earlier studies, it also provides a framework for studying fundamental performance limits of systematic codes in the presence of a wiretap channel. The best achievable performance of systematic codes is then compared to that of a general code in several respects, and a few examples are given.

Index Terms: wiretap channel, encryption, Shannon's cipher system, separation theorem, systematic codes.

1   Introduction

Wyner, in his well–known paper on the wiretap channel [11], studied the problem of secure communication across a degraded broadcast channel, without using a secret key, where the legitimate receiver has access to the output of the good channel and the wiretapper receives the output of the bad channel. In that paper, Wyner characterized the optimum trade–off between reliable coding rates and the equivocation at the wiretapper, which was defined in terms of the conditional entropy of the source given the output of the bad channel, observed by the wire–tapper. Among other things, Wyner establised and characterized, in the same paper, the notion of the *secrecy capacity*, which is the maximum coding rate that still allows full secrecy, where the equivocation is equal to the (unconditional) entropy of the source, thus rendering the information available to the wiretapper, virtually useless for learning anything about the source. By applying good codes at rates close to the secrecy capacity, the channel is fully exploited in the sense that the "excess noise", that is sufferred at the bad channel output (beyond the noise at the good channel output), plays the role of securing the message with maximum efficiency. The idea behind the construction of a good code for the wiretapped channel is essentially similar to the idea of binning. One creates a relatively large code, which is reliably decodable at the legitimate receiver, and which is thought of as an hierarchy of randomized sub–codes, each of which being reliably decodable individually by the wiretapper. However, the bits that are decodable by the wiretapper are only those of the randomization, and thus carry information that is irrelevant with regard to the source.

Throughout the three decades that have passed since [11] was published, the results of that paper have been extended in quite many directions, and we mention here only a few. Csiszár and Körner [3] have generalized Wyner's setting to a broadcast channel that is not necessarily degraded (allowing also a common message to both receivers). Very shortly afterwards, Leung–Yan–Cheong and Hellman [4], studied the Gaussian wiretap channel, and have shown, among other things that its secrecy capacity is simply the difference between the capacities of the main (legitimate) channel and the wiretap channel. In [8], Ozarow and Wyner studied another model, referred to as the type II wiretap channel, where the main communication channel is noiseless, but the wiretapper has access to a subset of the coded bits, and optimal tradeoffs were characterized. In [13], the wiretap channel model was extended to have two parallel broadcast channels, connecting one encoder and

one legitmate decoder, where both channels are wiretapped by non–collaborating wiretappers, and again, optimum tradeoffs where given in terms of single–letter expressions. In [14], the scope of [11] was extended in two ways: First, by allowing a secret key to be shared between the encoder and the legitimate receiver, and secondly, by allowing a certain distortion in the reconstruction of the source at the legitimate receiver. The main coding theorem of [14] suggests a separation principle, which asserts that no asymptotic optimality is lost if the encoder, first, applies a rate–distortion source code, then encrypts the compressed bits, and finally, applies a good code for the wiretap channel. More recently, the Gaussian wiretap channel model of [4] was further extended in two directions: one is the Gaussian multiple access wiretap channel of [10], and the other is Gaussian intereference wiretap channel of [6], [7], where the encoder has access to the interference signal as side information, similarly as in Costa's dirty paper channel [1].

In this paper, we extend the setting of the wiretap channel in a different direction. For simplicity, we adopt the structure of a (stochastically) degraded broadcast channel, as in [11] (though it is plausible that the results are generalizable to more general broadcast channels), and similarly as in [14], we allow a secret key shared between the encoder and the authorized decoder, as well as lossy reconstruction of the source within a prescribed distortion level, but we, moreover, allow also side informations, correlated to the source, to be available both to the legitimate decoder and the wiretapper. We assume that the wiretapper receives its side information via a channel that is degraded relative to the side information channel of the innocent decoder (see Fig. 1). Our main result is a single–letter characterization of the optimum tradeoff among five figures of merit: the equivocation at the wiretapper, the distortion level in reconstructing the source at the authorized decoder, the bandwidth expansion factor of the coded channels, the rate of the secret key relative to the source, and the average tranmission cost.

One of the motivations for this study is that it establishes a framework for deriving performance limits of systematic codes for wiretapped channels and assessing their loss in performance compared to general codes (as was done in [9] in a different context): The side information channels ($P_{V|U}$ and $P_{W|V}$ in Fig. 1) can be thought of as conveying the systematic (uncoded) part of the codeword. We compare the best achievable performance of systematic codes to that of general codes at the same coding rates, in several aspects, like the maximum achievable equivocation in the absence of a secret key, the maximum achievable equivocation in the presence of a full–rate key, the key
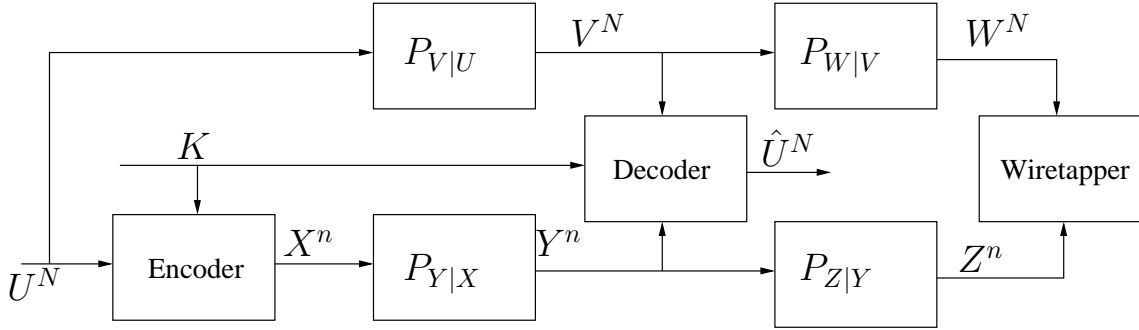
Figure 1: The wiretap channel with side information at the receivers.

rate needed to achieve the maximum achievable equivocation, and the distortion achieved when the channel is utilized at rate close to the secrecy capacity. A few examples are given for situations where systematic codes are as good as (and in some sense, sometimes even better than) general codes.

The outline of the remaining parts of this paper is as follows: In Section 2, we set up the notation, formulate the problem, present the main result, and make a few comments. In Section 3, we discuss the implications on systematic coding, and we make comparisons with general codes, as described in the previous paragraph. In Section 4, we prove the converse part of the main result, and finally, in Section 5, we prove the direct part.

2   Problem Formulation and Main Result

We begin by establishing some notation conventions. Throughout this paper, scalar random variables (RV's) will be denoted by capital letters, their sample values will be denoted by the respective lower case letters, and their alphabets will be denoted by the respective calligraphic letters. A similar convention will apply to random vectors and their sample values, which will be denoted with same symbols superscripted by the dimension, or by the bold face font, if there is no room for confusion regarding the dimension. Thus, for example, U^N (N – positive integer) or $\boldsymbol{U}$ will denote a random N–vector $(U_1, ..., U_N)$, and $u^N = (u_1, ..., u_N)$ is a specific vector value in $\mathcal{U}^N$, the N–th Cartesian power of $\mathcal{U}$.

Sources and channels will be denoted generically by the letter P, subscripted by the name of the RV and its conditioning, if applicable, e.g., $P_U(u)$ is the probability function of U at the point

$U = u$, $P_{Y|X}(y|x)$ is the conditional probability of $Y = y$ given $X = x$, and so on. Whenever clear from the context, these subscripts will be omitted. Information theoretic quantities like entropies and mutual informations will be denoted following the usual conventions of the Information Theory literature, e.g., $H(U^N)$, $I(X^n; Y^n)$, and so on. For single–letter information quantities (i.e., when $n = 1$ or $N = 1$), subscripts will be omitted, e.g., $H(U^1) = H(U_1)$ will be denoted by $H(U)$, similarly, $I(X^1; Y^1) = I(X_1; Y_1)$ will be denoted by $I(X; Y)$, and so on. For three random variables, generically denoted A, B, and C, the notation $A \ominus B \ominus C$ will designate the fact that they form, in this order, a Markov chain. The extension of this notation to longer Markov chains will be straightforward. The cardinality of a finite set $\mathcal{A}$ will be denoted by $|\mathcal{A}|$. The notation $[a]_+$ will stand for $\max\{0, a\}$. Finally, for $a, b \in \{0, 1\}$, $a \oplus b$ will denote the modulo 2 sum (XOR) of a and b, and for two general positive integers, a and b, the notation $a \oplus b$ will designate the positive integer whose binary representation is given by the bit–wise modulo 2 sum of the corresponding bits of the binary representations of a and b.

We now turn to the formal description of the model and the problem setting. A source P_U generates a sequence of N (N – positive integer) independent copies, $U^N = (U_1, \ldots, U_N)$, of a finite–alphabet[1] RV, $U \in \mathcal{U}$. At the same time, a discrete memoryless channel (DMC), symbolized by $P_{V|U}$ generates from U^N, another N-vector $V^N = (V_1, \ldots, V_N)$, with components in a finite–alphabet $\mathcal{V}$, and another DMC, denoted $P_{W|V}$, produces from V^N, yet another N-vector $W^N = (W_1, \ldots, W_N)$, with components in a finite–alphabet $\mathcal{W}$. Thus, the joint probability distribution of (u^N, v^N, w^N) is given by

$$P_{U^N}(u^N)P_{V^N|U^N}(v^N|u^N)P_{W^N|V^N}(w^N|v^N) = \prod_{i=1}^{N}[P_U(u_i)P_{V|U}(v_i|u_i)P_{W|V}(w_i|v_i)].$$

At the same time and independently, another source P_K, henceforth referred to as the *key source*, generates a random variable (or vector) K taking values in a finite alphabet $\mathcal{K}$.

Two additional cascaded DMC's operate at a bandwidth expansion factor of n/N channel uses per source symbol. This means that during the time that the source generates a block U^N of N symbols, the first channel receives a block X^n of n channel input symbols taking on values in a finite alphabet $\mathcal{X}$, and outputs a block Y^n of n channel output symbols in a finite alphabet $\mathcal{Y}$,

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[1]While we assume finite alphabet sources and channels throughout the derivations, some examples in the sequel pertain to Gaussian distributions, with the justification that similar results can be proven for this case as well.

according to

$$P_{Y^n|X^n}(y^n|x^n) = \prod_{j=1}^{n} P_{Y|X}(y_j|x_j),$$

whereas the second DMC receives Y^n as an input vector and outputs a block Z^n of n channel output symbols in a finite alphabet $\mathcal{Z}$, according to

$$P_{Z^n|Y^n}(z^n|y^n) = \prod_{j=1}^{n} P_{Z|Y}(z_j|y_j).$$

Given N and n, a block encoder is a (typically randomized) mapping $f_{n,N} : \mathcal{U}^N \times \mathcal{K} \to \mathcal{X}^n$, whose output is $X^n = (X_1, \ldots, X_n) = f_{n,N}(U^N, K) \in \mathcal{X}^n$. The channel input vector should satisfy an average transmission cost (generalized power) constraint:

$$\frac{1}{n} \sum_{j=1}^{n} \boldsymbol{E}\{\phi(X_j)\} \leq Q, \tag{1}$$

where $\phi : \mathcal{X} \to \mathbb{R}^+$ is the generalized power function and Q is a given positive real. Having defined the encoder, the definition of the joint distribution of the previously described RV's is completed by the Markov condition $W^N \ominus V^N \ominus (K, U^N) \ominus X^n \ominus Y^n \ominus Z^n$. The corresponding block decoder (of the authorized party) is a mapping $g_{n,N} : \mathcal{Y}^n \times \mathcal{V}^N \times \mathcal{K} \to \hat{\mathcal{U}}^N$, whose output is $\hat{U}^N = (\hat{U}_1, \ldots, \hat{U}_N) = g_{n,N}(Y^n, V^N, K) \in \hat{\mathcal{U}}^N$, where $\hat{\mathcal{U}}$ is the reproduction alphabet of the decoder output symbols. The decoding function $g_{n,N}$ will be taken as a deterministic mapping, since randomized decoding would be of no advantage.

Let $d : \mathcal{U} \times \hat{\mathcal{U}} \to \mathbb{R}^+$ denote a single–letter distortion measure between source symbols and reproduction symbols, and let the distortion between the vectors, $u^N \in \mathcal{U}^N$ and $\hat{u}^N \in \hat{\mathcal{U}}^N$, be defined additively across the corresponding components, as usual. Let $R_{U|V}(D)$ denote the Wyner–Ziv rate–distortion function [12] of the soure U with respect to the distortion measure d, and a decoder side information V, i.e.,

$$R_{U|V}(D) = \inf[I(U; A) - I(V; A)],$$

where the infimum is over all RV's A that form a Markov chain $A \ominus U \ominus V$ and that satisfy $\min_{\{\psi:\mathcal{A}\times\mathcal{V}\to\hat{\mathcal{U}}\}} \boldsymbol{E}\{d(U, \psi(A, V))\} \leq D$. This infimum does not change if the alphabet size of A is limited to $|\mathcal{U}| + 1$. Given the degraded broadcast channel $P_{YZ|X}(y, z|x) = P_{Y|X}(y|x)P_{Z|Y}(z|y)$, we

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will also define the function

$$\Gamma(r, q) = \sup_{\{P_X: \ I(X;Y) \geq r, \ \boldsymbol{E}\phi(X) \leq q\}} I(X;Y|Z) = \sup_{\{P_X: \ I(X;Y) \geq r, \ \boldsymbol{E}\phi(X) \leq q\}} [I(X;Y) - I(X;Z)] \quad (2)$$

which is similar to Wyner's Γ function [11], but with the additional generalized power constraint. The supremum over an empty set is understood to be $-\infty$.

An $(N, n, \lambda, D, \Delta, R, Q)$ codec is an encoder–decoder pair with parameters N and n, that satisfies the following requirements:

1. The bandwidth expansion factor is $n/N \leq \lambda$.

2. The expected distortion between the source and the reproduction satisfies

$$\sum_{i=1}^{N} \boldsymbol{E}\{d(U_i, \hat{U}_i)\} \leq ND. \quad (3)$$

3. The equivocation of the message source satisfies

$$H(U^N | W^N, Z^n) \geq N\Delta. \quad (4)$$

4. The rate of the secret key is $H(K)/N \leq R$.

5. The generalized transmission power satisfies $\sum_{j=1}^{n} \boldsymbol{E}\{\phi(X_j)\} \leq nQ$.

A quintuple $(\lambda, D, \Delta, R, Q)$ is said to be *achievable* if for every $\epsilon > 0$, there is a sufficiently large N and n for which $(N, n, \lambda + \epsilon, D + \epsilon, \Delta - \epsilon, R + \epsilon, Q + \epsilon)$ codecs exist. The *achievable region* of quintuples $\{(\lambda, D, \Delta, R, Q)\}$ is the set of all achievable quintuples $(\lambda, D, \Delta, R, Q)$.

The following theorem characterizes the region of achievable quintuples $(\lambda, D, \Delta, R, Q)$.

Theorem 1 *A quintuple* $(\lambda, D, \Delta, R, Q)$ *is achievable iff*

$$\Delta \leq \Delta^*(\lambda, R, D, Q) \triangleq H(U|W) - \left[R_{U|V}(D) - \lambda\Gamma\left(\frac{R_{U|V}(D)}{\lambda}, Q \right) - R \right]_+.$$

Discussion: A few comments are in order at this point.

As mentioned in the Introduction, Theorem 1 generalizes earlier results reported in [11], [9], and [14]. The generalization relative to [14, Theorem 1, "Case of LDBC"] is primarily in the presence

of side informations at the authorized decoder as well as the wiretapper. It should be also noted that in [14], there is no full proof of the direct part, but only an intuitive argument. Here, we provide complete proofs for both the converse part and the direct part, which are both based on the corresponding proofs in [11], but there are a few twists that are necessary in order to incorporate the secret key, K, the side informations, V^N and W^N, and the generalized power constraint. For example, one of the additional ingredients in the proof of the direct part, that is not present in the direct part of [11], is that we need to show, for the codec in the proof, that the key K can be estimated reliably from U^N, W^N, and Z^n, so that $H(K|U^N, W^N, Z^n)$ is small.

As in [14], Theorem 1 here suggests a *separation principle*, that guarantees no loss in asymptotically optimum performance, if one separates source coding, encryption, and channel coding. As will be seen in the proof of the direct part, the proposed achievability scheme consists of Wyner–Ziv rate-distortion source coding, followed by encryption of the compressed bits, followed in turn by good channel coding for the wiretapped channel, as in [11]. As is demonstrated in [5], the separation principle does not always hold in situations that involve source coding, encryption, and channel coding.

A few words about the intuition behind the achievable upper bound on the equivocation, $\Delta^*(\lambda, R, D, Q)$: For $R \geq R_{U|V}(D) - \lambda\Gamma(R_{U|V}(D)/\lambda, Q)$, there is enough randomness to achieve the maximum possible secrecy of $H(U^N|W^N) = NH(U|W)$, which cannot be exceeded even if the wiretapper did not have access to Z^n. For the more interesting case where $R_{U|V}(D) > \lambda\Gamma(R_{U|V}(D)/\lambda, Q)$ (which in turn means that $R_{U|V}(D)/\lambda$ is above the secrecy capacity), and $R < R_{U|V}(D) - \lambda\Gamma(R_{U|V}(D)/\lambda, Q)$, we can express $\Delta^*(\lambda, R, D, Q)$ as the sum of four terms:

$$\Delta^*(\lambda, R, D, Q) = [H(U|W) - H(U|V)] + [H(U|V) - R_{U|V}(D)] + \lambda\Gamma\left(\frac{R_{U|V}(D)}{\lambda}, Q\right) + R,$$

where we have added and subtracted $H(U|V)$. Now, the first bracketed term designates the fact that the wiretapper has side information whose quality is lower than that of the authorized user, a fact which contributes to the equivocation. The second bracketed term designates uncertainty due to the information loss at the source encoder (although a general coding scheme may not necessarily use a source encoder explicitly). Out of the $NR_{U|V}(D)$ bits of the description of the source, NR bits are covered by the key and another $n\Gamma(R_{U|V}(D)/\lambda, Q) = N\lambda\Gamma(R_{U|V}(D)/\lambda, Q)$ bits are covered by good channel coding for the wiretapped channel, as in [11], [14]. In designing a good coding

scheme, it should be kept in mind then, that there should be no overlap between the set of bits encrypted by the key and those that are "hidden" by coding. It is interesting to note that in the above decomposition of $\Delta^*(\lambda, R, D, Q)$, the first term depends solely on the joint distribution of (U, V, W), and not on any other factor of the problem, the second term depends only on the joint distribution of (U, V) and the allowed distortion (but no longer on the joint distribution with W), and the third term depends also the coded channels. Referring to the previous comment, it is interesting to note that even in the lossless case ($D = 0$) and even if the coded channels are clean (i.e., $X^n = Y^n = Z^n$ with probability one), the presence of side information at the legitimate decoder, which is of better quality than the one at the wiretapper, gives rise to "inherent secrecy," that is present even without a secret key. In such a case, the last three terms in the above representation of $\Delta^*(\lambda, R, D, Q)$ all vanish, but the first term is still positive. For example, a Slepian–Wolf encoder for a source U and side information V, which is based on random binning, has the maximum achievable inherent secrecy of $I(U; V)$ bits/symbol if a wiretapper that observes the compressed bits has no side information. This is in contrast to the case without side information, where there is no inherent secrecy at all.

An interesting question that arises is about optimum strategies and performance limits if one is interested to maximize the equivocation of $\hat{U}^N$ instead of, or in addition to that of U^N (see also [5]), which is reasonable because it is $\hat{U}^N$ that is the information conveyed from the source. In contrast to [5], where the problem was fully solved using ordinary rate–distortion coding considerations, here, because of the presence of side information, the problem remains open.

Finally, as mentioned already in the Abstract and the Introduction, Theorem 1 provides a framework for studying the fundamental performance limits of *systematic* (not necessarily linear) codes, in the same manner as in [9], for the wiretap channel. The next section is devoted to such a study.

3 Systematic Vs. Non–Systematic Codes

If $\mathcal{U} = \mathcal{X}$, $\mathcal{V} = \mathcal{Y}$, and the uncoded channel, $P_{V|U}$, is understood as an additional use of the same physical channel as the coded channel, $P_{Y|X}$, and if $\boldsymbol{E}\{\phi(U)\} \leq Q$, then the uncoded path $U^N \to V^N$ may be thought of as corresponding to the transmission and reception of the systematic

(uncoded) part of a systematic code, where the information symbols are sent directly to the channel. The total bandwidth expansion factor of this systematic code, when the uncoded part is viewed as part of the code, is then $(N + n)/N = 1 + \lambda$, assuming that $n/N = \lambda$. For a fully coded (general, non–systematic) system with the same bandwidth expansion factor, we can use the formula of $\Delta^*(\lambda, R, D, Q)$, but replace λ by $1 + \lambda$ and eliminate the side informations, V^N and W^N. The resulting maximum achievable equivocation of a general code, is therefore:

$$\Delta^*_{\text{gen}}(\lambda, R, D, Q) = H(U) - \left[R_U(D) - (1 + \lambda)\Gamma \left(\frac{R_U(D)}{1 + \lambda}, Q \right) - R \right]_+, \tag{5}$$

where $R_U(D)$ is the ordinary rate–distortion function of U (without side information), and we are interested to compare this to the original expression of $\Delta^*(\lambda, R, D, Q)$, given in Theorem 1, which will be denoted by $\Delta^*_{\text{sys}}(\lambda, R, D, Q)$ throughout this section. Quite obviously, $\Delta^*_{\text{sys}}(\lambda, R, D, Q)$ cannot exceed $\Delta^*_{\text{gen}}(\lambda, R, D, Q)$, but it is interesting to identify cases of equality, simply by comparing the two expressions. We will, however, focus here on a few specfic aspects of comparison between optimum systematic codes and optimum general codes:

1. The *full equivocation*, that is, the maximum equivocation that can be achieved in the absence of limitations on the key rate (in which case, the bracketed term of Δ^* vanishes).

2. The *zero key–rate equivocation*, which is defined as Δ^* for $R = 0$. This quantity manifests the "inherent" security that is already present in the system even without a key. It should be noted that whenever $\Delta^*_{\text{sys}}(\lambda, 0, D, Q) = \Delta^*_{\text{gen}}(\lambda, 0, D, Q)$, then, in general (as can be seen from the expressions of $\Delta^*_{\text{sys}}(\lambda, R, D, Q)$ and $\Delta^*_{\text{gen}}(\lambda, R, D, Q)$), there is a range of R, where $\Delta^*_{\text{sys}}(\lambda, R, D, Q) = \Delta^*_{\text{gen}}(\lambda, R, D, Q)$ since, in that range, both $\Delta^*_{\text{sys}}(\lambda, R, D, Q)$ and $\Delta^*_{\text{gen}}(\lambda, R, D, Q)$ grow linearly with a slope of 45 degrees, starting from their respective values at $R = 0$.

3. The *saturation key rate*, which is the smallest value of R, for which Δ^* achieves the full equivocation. When the saturation key rate is small, then so are the randomization resources required.

4. The *secrecy distortion*, which is the value of D for which the channel coding rate equals the secrecy capacity, in other words, the first argument of the function Γ agrees with the secrecy capacity. This is an interesting working point, because it is the point where the full

equivocation is achieved without using a key at all. In other words, using the terminology that we have already defined, the zero key–rate equivocation is equal to the full equivocation, and the saturation key rate vanishes.

While under the first two criteria, systematic codes can never be strictly better than general codes, this is not necessarily the case with the last two criteria, because codes that are optimum in the maximum equivocation sense may be suboptimal under other criteria. We next compare optimum systematic codes to optimum codes from the above four aspects.

1. The full equivocation: Obviously, this quantity is $H(U|W)$ for systematic codes and $H(U)$ for general codes, thus the difference, $I(U;W)$, depends only on the joint distribution of U and W. In this respect, optimum systematic codes are as good as optimum general codes only if the side information W is independent of U and hence useless.

2. The zero key–rate equivocation: For $R = 0$, we have

$$\Delta_{\text{gen}}^*(\lambda, 0, D, Q) = H(U) - \left[R_U(D) - (1 + \lambda)\Gamma\left(\frac{R_U(D)}{1 + \lambda}, Q\right) \right]_+ \tag{6}$$

for general codes, and

$$\Delta_{\text{sys}}^*(\lambda, 0, D, Q) = H(U|W) - \left[R_{U|V}(D) - \lambda\Gamma\left(\frac{R_{U|V}(D)}{\lambda}, Q\right) \right]_+ \tag{7}$$

for systematic codes. Let us assume that the bracketed terms in both expressions are positive (otherwise, we are back to the comparison of the previous paragraph). Comparing the two expressions, we see that equality is achieved if

$$(1 + \lambda)\Gamma\left(\frac{R_U(D)}{1 + \lambda}, Q\right) - \lambda\Gamma\left(\frac{R_{U|V}(D)}{\lambda}, Q\right) = R_U(D) - R_{U|V}(D) - I(U;W). \tag{8}$$

As is shown in [9, eqs. (2.12), (2.13)], the difference $R_U(D) - R_{U|V}(D)$ is never larger than $I(U;V)$, but there are cases of equality, most notably, the lossless case $D = 0$, as $R_U(0) = H(U)$ and $R_{U|V}(0) = H(U|V)$.[2] Thus, at least in the lossless case, eq. (8) boils down to

$$(1 + \lambda)\Gamma\left(\frac{H(U)}{1 + \lambda}, Q\right) - \lambda\Gamma\left(\frac{H(U|V)}{\lambda}, Q\right) = I(U;V) - I(U;W). \tag{9}$$

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[2]Another example is the Gaussian source U, the Gaussian channel $P_{V|U}$, and the squared error distortion measure, where $R_U(D) - R_{U|V}(D) = [h(U) - \frac{1}{2}\log(2\pi e D)] - [h(U|V) - \frac{1}{2}\log(2\pi e D)] = I(U;V)$ throughout the entire interesting range of distortion levels.

Now, in quite a few examples of interest, $\Gamma(r, q)$ is equal to a constant, Γ_0, throughout the entire interesting range of r. One such example occurs when $q = \infty$ (i.e., no generalized power constraint), $P_{Y|X}$ is the noiseless binary channel and $P_{Z|Y}$ is a binary symmetric channel (BSC) with crossover probability p_0 (cf. [11, p. 1362]), in which case, $\Gamma_0 = h(p_0)$, where $h(\cdot)$ is the binary entropy function. In this case, the left–hand side of eq. (9) becomes $h(p_0)$ independently of λ. Now, if $P_{V|U}$ has the same characteristics as $P_{Y|X}$, and similatry $P_{W|V}$ has the same characteristics as $P_{Z|Y}$ (which is indeed the case in systematic coding applications), and if P_U is the binary symmetric source (BSS), then it achieves the maximum of $I(U; V) - I(U; W)$, which is, again, $\Gamma_0 = h(p_0)$. In this case, therefore, the equality (9) is achieved. Similarly, if $P_{Y|X}$ is noiseless as before, but $P_{Z|Y}$ is an erasure channel with erasure probability p_0, then $\Gamma_0 = p_0$, and once again, equality is achieved if P_U is the BSS. Yet another example of this type occurs when both $P_{Y|X}$ and $P_{Z|Y}$ are independent Gaussian channels with an input power constraint defined in terms of $\phi(x) = x^2$ (and hence, so are $P_{V|U}$ and $P_{W|V}$). In this case, as was shown in [4], $\Gamma_0 = C_{X\to Y} - C_{X\to Z}$, the difference between the capacities of the channels $P_{Y|X}$ and $P_{Z|X}$. Here, equality in (9) is achieved if U is a zero–mean Gaussian random variable whose variance coincides with the maximum allowable input power, Q. Thus, we have demonstrated a few non–trivial examples where optimum systematic codes are as good as optimum codes in the absence of a secret key.

Note that the above examples are valid regardless of the value of λ. It is then interesting to examine the extreme case $\lambda \to 0$ (or even $\lambda = 0$), namely, a completely uncoded system, where the source is connected directly to the channel, without coding at all. In these examples, where the sources are matched to the channels, it is well known that optimum joint source–channel coding is achieved this way. Here we see that these systems are optimum also from the point of view of secrecy.

3. **The saturation key rate:** Here, we obtain

$$R^*_{\text{gen}} = R_U(D) - (1 + \lambda)\Gamma\left(\frac{R_U(D)}{1 + \lambda}, Q\right) \tag{10}$$

for general codes, and

$$R^*_{\text{sys}} = R_{U|V}(D) - \lambda\Gamma\left(\frac{R_{U|V}(D)}{\lambda}, Q\right) \tag{11}$$

for systematic codes. The condition for having a smaller saturation key rate for systematic codes is

$$(1+\lambda)\Gamma\left(\frac{R_U(D)}{1+\lambda},Q\right) - \lambda\Gamma\left(\frac{R_{U|V}(D)}{\lambda},Q\right) \le R_U(D) - R_{U|V}(D), \tag{12}$$

namely, the comparison is similar to the one made with regard to the zero key–rate equivocation criterion, but without the term $I(U;W)$. As we have previously shown examples of equality, even in the presence of the term $I(U;W)$, then the same examples can serve now for the desired inequality in the absence of this term. In these examples, as well as in many others, optimum systematic codes are advantageous over optimum codes in general. This apparently counterintuitive situation is possible because the requirement $\Delta^* = H(U)$ for general codes is more demanding than $\Delta_* = H(U|W)$ for systematic codes.[3]

4. The secrecy distortion: As mentioned earlier, Wyner [11] has established the notion of the secrecy capacity, C_s, which is the maximum coding rate for which full secrecy is still achieved even without a key. Here we ask how do systematic– and non–systematic codes compare in terms of the distortion, D, for which the rate of the channel code meets the secrecy capacity. For non–systematic codes, this distortion level is given by the solution to the equation

$$\frac{R_U(D)}{1+\lambda} = C_s, \tag{13}$$

which is

$$D^*_{\text{gen}} = D_U((1+\lambda)C_s), \tag{14}$$

where $D_U(\cdot)$ is the ordinary distortion–rate function of U (without side information). For systematic coding, on the other hand, it is the solution to the equation

$$\frac{R_{U|V}(D)}{\lambda} = C_s, \tag{15}$$

which is

$$D^*_{\text{sys}} = D_{U|V}(\lambda C_s), \tag{16}$$

where $D_{U|V}(\cdot)$ is the Wyner–Ziv distortion–rate function of U with side information V. The answer to the question: which class of codes is better in terms of the secrecy distortion, depends on the

---

[3]Similarly, in the other aspects of comparison in the sequel, systematic codes might sometimes be advantageous due to the different requirements.

parameters of the problem. One simple extreme example pertains to the case $C_s = 0$ (which happens, e.g., when the channel $P_{Z|Y}$ is clean and hence $Z^n = Y^n$ with probability one). In this case,

$$D^*_{\text{sys}} = D_{U|V}(0) = \min_{\{\psi : \mathcal{V} \to \hat{\mathcal{U}}\}} \boldsymbol{E}\{d(U, \psi(V))\} \tag{17}$$

is clearly smaller than

$$D^*_{\text{gen}} = D_U(0) = \min_{\hat{u} \in \hat{\mathcal{U}}} \boldsymbol{E}\{d(U, \hat{u})\}. \tag{18}$$

While the case where C_s is strictly zero, clearly trivializes the whole problem altogether, it is, of course, conceivable that for small enough positive values of C_s, continuity arguments imply that systematic codes still outperform non–systematic codes in the secrecy distortion sense.

As a somewhat less trivial example, consider the case where U is zero–mean, Gaussian, with variance σ_U^2, the channels are Gaussian and independent, and d is the squared error criterion. Then,

$$D^*_{\text{gen}} = \sigma_U^2 \cdot 2^{-2(1+\lambda)(C_{X\to Y} - C_{X\to Z})}, \tag{19}$$

whereas

$$D^*_{\text{sys}} = \sigma_{U|V}^2 \cdot 2^{-2\lambda(C_{X\to Y} - C_{X\to Z})}, \tag{20}$$

where $\sigma_{U|V}^2$ is the minimum mean squared error associated with optimum (linear) estimation of U based on V. Thus, $D^*_{\text{sys}} \leq D^*_{\text{gen}}$ whenever

$$C_{X\to Y} - C_{X\to Z} \leq \frac{1}{2} \log \frac{\sigma_U^2}{\sigma_{U|V}^2} = \frac{1}{2} \log \left(1 + \frac{\sigma_U^2}{\sigma^2}\right), \tag{21}$$

where σ^2 is the variance of the noise of the (Gaussian) channel from U to V. Note that the dependence upon λ disappeared. The last inequality is clearly met if, for example, the channel $P_{Y|X}$ is the same as the channel $P_{V|U}$ and $\sigma_U^2 = Q$ (in which case, the right–hand side becomes $C_{X\to Y}$).

Note that in this aspect of the secrecy distortion, our comparison between systematic codes and non–systematic codes is of the same spirit as in [9], in the sense that both are about equating rate–distortion functions to capacities. The only difference is that here, as opposed to [9], C_s replaces $C_{X\to Y}$ in the these equations (as there is only one coded channel and one uncoded channel in [9]). Obviously, in the comparisons carried out in [9], systematic codes can never outperform

non–systematic codes. By contrast, as we have seen here, when the secrecy capacity is the working point, this becomes possible.

Finally, one more comment is in order regarding systematic codes: In a real systematic code for the wiretap channel, there is, in principle, the freedom to use part of the secret key in order to encrypt the systematic symbols as well. This freedom has not been exploited thus far, and the question is whether there is any advantage in doing so. Suppose that the source U is binary and the key rate is R bits per source symbol ($R \leq 1$). Consider the following coding scheme. We select $0 \leq R' \leq R$, and for each block U^N, we use NR' key bits to encrypt the systematic part and $N(R - R')$ key bits to encrypt the Wyner–Ziv rate–distortion codeword before it is fed into the channel encoder of [11] (see also the proof of the direct part in Section 5). Then, by a slight extension of the analysis in Section 5 to follow, the resulting equivocation is essentially

$$\Delta \approx R'H(U) + (1 - R')H(U|W) + (R - R') + \lambda\Gamma(R_{U|V}(D)/\lambda, Q) - R_{U|V}(D). \qquad (22)$$

Since the coefficient of R', in this expression, is $I(U; W) - 1 < 0$, the best choice of R', in this example, is $R' = 0$, namely, secret key bits should better not be used for encrypting the systematic bits, but only the coded bits, as we assumed thus far.

4   Proof of the Converse Part of Theorem 1

Let an $(N, n, \lambda + \epsilon, D + \epsilon, \Delta - \epsilon, R + \epsilon, Q + \epsilon)$ codec be given. Consider first the following chain of inequalities, which will be used later on.

$$
\begin{aligned}
I(X^n; Y^n | K, V^N, W^N) \;&\overset{(a)}{\geq}\; I(U^N; Y^n | K, V^N, W^N) \\
&= \sum_{i=1}^N I(U_i; Y^n | K, V^N, W^N, U^{i-1}) \\
&= \sum_{i=1}^N [H(U_i | K, V^N, W^N, U^{i-1}) - H(U_i | Y^n, K, V^N, W^N, U^{i-1})] \\
&\overset{(b)}{\geq}\; \sum_{i=1}^N [H(U_i | V_i) - H(U_i | Y^n, K, V^N)] \\
&\overset{(c)}{=}\; \sum_{i=1}^N [H(U_i | V_i) - H(U_i | V_i, A_i)] \\
&= \sum_{i=1}^N I(U_i; A_i | V_i) \\
&= \sum_{i=1}^N [H(A_i | V_i) - H(A_i | U_i, V_i)] \\
&\overset{(d)}{=}\; \sum_{i=1}^N [H(A_i | V_i) - H(A_i | U_i)] \\
&= \sum_{i=1}^N [I(U_i; A_i) - I(V_i; A_i)] \\
&\overset{(e)}{\geq}\; \sum_{i=1}^N R_{U|V}(\boldsymbol{E} d(U_i, [g_{n,N}(A_i, V_i)]_i)) \\
&\overset{(f)}{\geq}\; N R_{U|V}\left(\frac{1}{N} \sum_{i=1}^N \boldsymbol{E} d(U_i, [g_{n,N}(A_i, V_i)]_i)\right) \\
&\overset{(g)}{\geq}\; N R_{U|V}(D + \epsilon),
\end{aligned}
\tag{23}
$$

where (a) follows from the fact that $U^N \ominus (K, V^N, W^N, X^n) \ominus Y^n$ is a Markov chain, (b) is because conditioning reduces entropy and because $U_i \ominus V_i \ominus (K, U^{i-1}, W^N, V^{i-1}, V_{i+1}^N)$ is a Markov chain, in (c) – A_i is defined as $(Y^n, K, V^{i-1}, V_{i+1}^N)$, (d) is because $A_i \ominus U_i \ominus V_i$ is a Markov chain, (e) is by definition of the Wyner–Ziv rate–distortion function, where $[g_{n,N}(A_i, V_i)]_i$ is the projection of

$g_{n,N}(A_i, V_i) = g_{n,N}(Y^n, V^N, K)$ to the i–th component, (f) is due to the convexity of the Wyner–Ziv rate–distortion function [12],[2, Lemma 14.9.1, p. 439], and (g) is due to its monotonicity, and the hypothesis that the codec achieves distortion $D + \epsilon$.

We next derive two upper bounds on Δ. The first one is trivial:

$$\Delta - \epsilon \le \frac{H(U^N|W^N, Z^n)}{N} \le \frac{H(U^N|W^N)}{N} = H(U|W), \tag{24}$$

and so,

$$\Delta \le H(U|W) \tag{25}$$

due to the arbitrariness of ϵ. The other, more interesting, upper bound on Δ is obtained as follows: First, we observe that

$$N(\Delta - \epsilon) \le H(U^N|W^N, Z^n) = I(U^N; V^N, K|W^N, Z^n) + H(U^N|V^N, K, W^N, Z^n). \tag{26}$$

Next, we bound from above each one of the terms on the right–most side. As for the first term, we have

$$
\begin{aligned}
I(U^N; V^N, K|W^N, Z^n) &= I(U^N; K|W^N, Z^n) + I(U^N; V^N|K, W^N, Z^n) \\
&\le H(K|W^N, Z^n) + H(V^N|K, W^N, Z^n) - H(V^N|U^N, K, W^N, Z^n) \\
&\le H(K) + H(V^N|W^N) - H(V^N|U^N, W^N) \\
&\le N(R + \epsilon) + NI(U; V|W) \\
&= N[R + I(U; V|W) + \epsilon], \tag{27}
\end{aligned}
$$

where in the second inequality we have used the fact that $V^N \ominus (U^N, W^N) \ominus (Z^n, K)$ is a Markov chain. As for the second term on the r.h.s. of (26), we have:

$$
\begin{aligned}
H(U^N|V^N, K, W^N, Z^n) &= H(U^N|W^N) - I(U^N; V^N, K, Z^n|W^N) \\
&= NH(U|W) - I(U^N; V^N|W^N) - I(U^N; K, Z^n|W^N, V^N) \\
&= N[H(U|W) - I(U; V|W)] - I(U^N; K, Z^n|W^N, V^N) \\
&= NH(U|V, W) - I(U^N; K, Y^n|W^N, V^N) + \\
&\quad [I(U^N; K, Y^n|W^N, V^N) - I(U^N; K, Z^n|W^N, V^N)]. \tag{28}
\end{aligned}
$$

We proceed by deriving a lower bound to $I(U^N; K, Y^n | W^N, V^N)$ and an upper bound to the bracketed term in the last expression. As for the former, we have:

$$I(U^N; K, Y^n | W^N, V^N) \geq I(U^N; Y^n | W^N, V^N, K) \geq N R_{U|V}(D + \epsilon), \tag{29}$$

where the second inequality has been proven above (compare the right–hand side of the first line of eq. (23) with the right–most side of that equation). As for the upper bound to the bracketed

term of the right–most side of (28), we have:

$$I(U^N; K, Y^n | W^N, V^N) - I(U^N; K, Z^n | W^N, V^N)$$

$$\overset{(a)}{=} I(U^N; K, Y^n, W^N, V^N) - I(U^N; K, Z^n, W^N, V^N)$$

$$\overset{(b)}{=} I(U^N, K; K, Y^n, W^N, V^N) - I(U^N, K; K, Z^n, W^N, V^N)$$

$$\leq I(U^N, K; K, Y^n, W^N, V^N) - I(U^N, K; K, Z^n, W^N, V^N) +$$

$$I(X^n; K, Y^n, Z^n, W^N, V^N | U^N, K) - I(X^n; K, Z^n, W^N, V^N | U^N, K)$$

$$\overset{(c)}{=} I(U^N, K; K, Y^n, W^N, V^N) - I(U^N, K; K, Z^n, W^N, V^N) +$$

$$I(X^n; K, Y^n, W^N, V^N | U^N, K) - I(X^n; K, Z^n, W^N, V^N | U^N, K)$$

$$= I(U^N, K, X^n; K, Y^n, W^N, V^N) - I(U^N, K, X^n; K, Z^n, W^N, V^N)$$

$$\overset{(d)}{=} I(U^N, K, X^n; Y^n | K, W^N, V^N) - I(U^N, K, X^n; Z^n | K, W^N, V^N)$$

$$\overset{(e)}{=} I(U^N, X^n; Y^n | K, W^N, V^N) - I(U^N, X^n; Z^n | K, W^N, V^N)$$

$$\overset{(f)}{=} I(X^n; Y^n | K, W^N, V^N) - I(X^n; Z^n | K, W^N, V^N)$$

$$= H(Y^n | K, W^N, V^N) - H(Z^n | K, W^N, V^N) +$$

$$H(Z^n | X^n, K, W^N, V^N) - H(Y^n | X^n, K, W^N, V^N)$$

$$= \sum_{i=1}^{n} [H(Y_i | Y^{i-1}, K, W^N, V^N) - H(Z_i | Z^{i-1}, K, W^N, V^N) +$$

$$H(Z_i | X_i, K, W^N, V^N) - H(Y_i | X_i, K, W^N, V^N)]$$

$$\overset{(g)}{\leq} \sum_{i=1}^{n} [H(Y_i | Y^{i-1}, K, W^N, V^N) - H(Z_i | Z^{i-1}, Y^{i-1}, K, W^N, V^N) +$$

$$H(Z_i | X_i, K, W^N, V^N) - H(Y_i | X_i, K, W^N, V^N)]$$

$$\overset{(h)}{=} \sum_{i=1}^{n} [H(Y_i | Y^{i-1}, K, W^N, V^N) - H(Z_i | Y^{i-1}, K, W^N, V^N) +$$

$$H(Z_i | X_i, Y^{i-1}, K, W^N, V^N) - H(Y_i | X_i, Y^{i-1}, K, W^N, V^N)]$$

$$= \sum_{i=1}^{n} [I(X_i; Y_i | Y^{i-1}, K, W^N, V^N) - I(X_i; Z_i | Y^{i-1}, K, W^N, V^N)]$$

$$= \sum_{i=1}^{n} [H(X_i | Z_i, Y^{i-1}, K, W^N, V^N) - H(X_i | Y_i, Y^{i-1}, K, W^N, V^N)]$$

$$\overset{(i)}{=} \sum_{i=1}^{n} [H(X_i | Z_i, Y^{i-1}, K, W^N, V^N) - H(X_i | Y_i, Z_i, Y^{i-1}, K, W^N, V^N)]$$

$$= \sum_{i=1}^{n} I(X_i; Y_i | Z_i, Y^{i-1}, K, W^N, V^N) \qquad (30)$$

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where (a) is by adding and subtracting $I(U^N; V^N, W^N)$, (b) is by adding $I(K; K, Y^n, W^N, V^N | U^N) = H(K|U^N)$ and subtracting $I(K; K, Z^n, W^N, V^N | U^N) = H(K|U^N)$, (c) is by the fact that $X^n \ominus (Y^n, K, U^N, V^N, W^N) \ominus Z^n$ is a Markov chain, (d) is by adding and subtracting $I(U^N, K, X^n; K, W^N, V^N)$, (e) is by the fact that K is degenerate as it appears in the conditioning, (f) is by the fact that $U^N \ominus (X^n, K, W^N, V^N) \ominus Y^n \ominus Z^n$ is a Markov chain, (g) is because conditioning reduces entropy, (h) is because $Z^{i-1} \ominus (Y^{i-1}, K, W^N, V^N) \ominus Z_i$ and $Z_i \ominus Y_i \ominus (X_i, K, V^N, W^N) \ominus Y^{i-1}$ are Markov chains, and (i) is because $X_i \ominus (Y_i, Y^{i-1}, K, W^N, V^N) \ominus Z_i$ is a Markov chain.

At this point, we are after an upper bound to $\sum_{i=1}^n I(X_i; Y_i | Z_i, Y^{i-1}, K, W^N, V^N)$, subject to the fact that

$$
\begin{aligned}
NR_{U|V}(D + \epsilon) &\leq I(X^n; Y^n | K, V^N, W^N) \\
&= \sum_{i=1}^n [H(Y_i | Y^{i-1}, K, V^N, W^N) - H(Y_i | X_i, Y^{i-1}, K, V^N, W^N)] \\
&= \sum_{i=1}^n I(X_i; Y_i | Y^{i-1}, K, V^N, W^N) \quad (31)
\end{aligned}
$$

where, once again, the first inequality has been proved already in (23). For given k, v^N, w^N, and y^{i-1}, $i = 1, 2, \ldots, n$, let

$$
\alpha_i(k, v^N, w^N, y^{i-1}) = I(X_i; Y_i | K = k, V^N = v^N, W^N = w^N, Y^{i-1} = y^{i-1}) \quad (32)
$$

and

$$
\beta_i(k, v^N, w^N, y^{i-1}) = \boldsymbol{E}\{\phi(X_i) | K = k, V^N = v^N, W^N = w^N, Y^{i-1} = y^{i-1}\}. \quad (33)
$$

Obviously, by definition of the function Γ,

$$
\begin{aligned}
& I(X_i; Y_i | Z_i, K = k, V^N = v^N, W^N = w^N, Y^{i-1} = y^{i-1}) \\
& \leq \Gamma(\alpha_i(k, v^N, w^N, y^{i-1}), \beta_i(k, v^N, w^N, y^{i-1})). \quad (34)
\end{aligned}
$$

Thus,

$$\sum_{i=1}^{n} I(X_i; Y_i | Z_i, Y^{i-1}, K, W^N, V^N)$$

$$= \sum_{i=1}^{n} \sum_{k, v^N, w^N, y^{i-1}} \Pr\{K = k, V^N = v^N, W^N = w^N, Y^{i-1} = y^{i-1}\} \times$$
$$I(X_i; Y_i | Z_i, K = k, V^N = v^N, W^N = w^N, Y^{i-1} = y^{i-1})$$

$$\leq \sum_{i=1}^{n} \sum_{k, v^N, w^N, y^{i-1}} \Pr\{K = k, V^N = v^N, W^N = w^N, Y^{i-1} = y^{i-1}\} \times$$
$$\Gamma(\alpha_i(k, v^N, w^N, y^{i-1}), \beta_i(k, v^N, w^N, y^{i-1}))$$

$$\overset{(a)}{\leq} n\Gamma\left(\frac{1}{n}\sum_{i=1}^{n} \sum_{k, v^N, w^N, y^{i-1}} \Pr\{K = k, V^N = v^N, W^N = w^N, Y^{i-1} = y^{i-1}\} \cdot \alpha_i(k, v^N, w^N, y^{i-1}),\right.$$

$$\left.\frac{1}{n}\sum_{i=1}^{n} \sum_{k, v^N, w^N, y^{i-1}} \Pr\{K = k, V^N = v^N, W^N = w^N, Y^{i-1} = y^{i-1}\} \cdot \beta_i(k, v^N, w^N, y^{i-1})\right)$$

$$= n\Gamma\left(\frac{1}{n}\sum_{i=1}^{n} I(X_i; Y_i | Y^{i-1}, K, W^N, V^N), \frac{1}{n}\sum_{i=1}^{n} \boldsymbol{E}\{\phi(X_i)\}\right)$$

$$\overset{(b)}{\leq} n\Gamma\left(\frac{N}{n} \cdot R_{U|V}(D + \epsilon), Q + \epsilon\right)$$

$$\overset{(c)}{\leq} n\Gamma\left(\frac{R_{U|V}(D + \epsilon)}{\lambda + \epsilon}, Q + \epsilon\right)$$

$$\overset{(d)}{\leq} N(\lambda + \epsilon)\Gamma\left(\frac{R_{U|V}(D + \epsilon)}{\lambda + \epsilon}, Q + \epsilon\right), \tag{35}$$

where (a) follows from the concavity of $\Gamma(r, q)$ jointly in both arguments,[4] together with its non–increasing monotonicity in r and non–decreasing monotonicity in q, (b)– from (31) and the non–increasing monotonicity of the function $\Gamma(\cdot)$, and (c) and (d) – from the postulate that the bandwidth expansion factor of the codec does not exceed $\lambda + \epsilon$. Combining eqs. (25), (26), (27), (28),

---

[4]This can readily be verified as a trivial extension of [11, Lemma 1] which accounts for the generalized power constraint.

(29), (30), and (35), and using the arbitrariness of ϵ with continuity considerations, we get

$$
\begin{aligned}
\Delta \;\leq\; & \min\left\{ H(U|W), R + I(U;V|W) + H(U|V,W) + \lambda\Gamma\left(\frac{R_{U|V}(D)}{\lambda}, Q\right) - R_{U|V}(D)\right\} \\
=\; & \min\left\{ H(U|W), H(U|W) + R + \lambda\Gamma\left(\frac{R_{U|V}(D)}{\lambda}, Q\right) - R_{U|V}(D)\right\} \\
=\; & H(U|W) - \left[R_{U|V}(D) - \lambda\Gamma\left(\frac{R_{U|V}(D)}{\lambda}, Q\right) - R\right]_{+} \\
=\; & \Delta^{*}(\lambda, R, D, Q),
\end{aligned}
\tag{36}
$$

which establishes the converse part of Theorem 1.

5 Proof of the Direct Part of Theorem 1

We begin with the following chain of equalities and inequalities:

$$
\begin{aligned}
N\Delta \;&=\; H(U^N|W^N,Z^n)\\
&=\; H(U^N,Z^n|W^N) - H(Z^n|W^N)\\
&=\; H(U^N,Z^n,X^n,K|W^N) - H(X^n,K|U^N,Z^n,W^N) - H(Z^n|W^N)\\
&=\; H(X^n,K,U^N|W^N) + H(Z^n|X^n,K,U^N,W^N) -\\
&\qquad H(X^n,K|U^N,Z^n,W^N) - H(Z^n|W^N)\\
&=\; H(Z^n|X^n,K,U^N,W^N) + H(U^N|W^N) +\\
&\qquad [H(X^n,K|U^N,W^N) - H(X^n,K|U^N,Z^n,W^N)] - H(Z^n|W^N)\\
&=\; NH(U|W) + H(Z^n|X^n,K,U^N,W^N) +\\
&\qquad I(X^n,K;Z^n|U^N,W^N) - H(Z^n|W^N)\\
&\overset{(a)}{\ge}\; NH(U|W) + H(Z^n|X^n) +\\
&\qquad I(X^n,K;Z^n|U^N,W^N) - H(Z^n)\\
&=\; NH(U|W) + I(X^n,K;Z^n|U^N,W^N) - I(X^n;Z^n)\\
&=\; NH(U|W) + I(K;Z^n|U^N,W^N) + I(X^n;Z^n|U^N,W^N,K) - I(X^n;Z^n)\\
&=\; NH(U|W) + H(K|U^N,W^N) - H(K|U^N,W^N,Z^n) +\\
&\qquad I(X^n;Z^n|U^N,W^N,K) - I(X^n;Z^n)\\
&\overset{(b)}{=}\; NH(U|W) + H(K) - H(K|U^N,W^N,Z^n) +\\
&\qquad I(X^n;Z^n|U^N,W^N,K) - I(X^n;Z^n)\\
&\overset{(c)}{=}\; NH(U|W) + NR - H(K|U^N,W^N,Z^n) +\\
&\qquad I(X^n;Z^n|U^N,W^N,K) - I(X^n;Z^n), \tag{37}
\end{aligned}
$$

where (a) follows from the fact that $(K,U^N,W^N) \ominus X^n \ominus Z^n$ is a Markov chain, (b) – from the fact that K is independent of (U^N,W^N), and (c) – by assuming that $H(K) = NR$. While this chain of equalities and inequalities holds for any codec, then in order to proceed, we will have to be specific, from now on, about the structure and the properties of the codec. In particular, referring to the right–most side of the above lower bound to $N\Delta$, then in order to prove the direct part, we will have

to prove that for our proposed codec (and as long as R is not too large): (i) $H(K|U^N, W^N, Z^n)$ is small, (ii) $I(X^n; Z^n)$ is essentially smaller than $nI(X; Z)$, and (iii) $I(X^n; Z^n|U^N, W^N, K)$ is essentially larger $[nI(X; Y) - NR_{U|V}(D)]$, where in (ii) and (iii) the distribution of the random variable X is the achiever of $\Gamma(R_{U|V}(D)/\lambda, Q)$.

Fix an arbitrarily small $\epsilon > 0$, and let D satisfy $R_{U|V}(D) \leq \lambda C_{X \to Y} - \epsilon$, where $C_{X \to Y}$ is the capacity of the channel $P_{Y|X}$ with power constraint Q (otherwise, the legitimate decoder cannot reconstruct U within distortion D [9]). Given such D and $\epsilon > 0$, let X^* denote the channel input variable that achieves $\Gamma((R_{U|V}(D) + \epsilon)/\lambda, Q)$. Let Y^* and Z^* denote the channel output variables induced by X^* and the channels $P_{Y|X}$ and $P_{Z|Y}$, respectively. Thus,

$$I(X^*; Y^*) - I(X^*; Z^*) = \Gamma\left(\frac{R_{U|V}(D) + \epsilon}{\lambda}, Q\right) \tag{38}$$

and

$$I(X^*; Y^*) \geq \frac{R_{U|V}(D) + \epsilon}{\lambda}. \tag{39}$$

Let us further suppose now that for the resulting optimal RV's X^*, Y^*, and Z^*, we have:

$$\frac{R_{U|V}(D)}{\lambda} > I(X^*; Y^*) - I(X^*; Z^*). \tag{40}$$

In the sequel, we will handle separately the case where (40) does not hold. Further, let $\mathcal{T}_n$ denote the set of $n^{-1/4}$–typical n–sequences with components in $\mathcal{X}$, i.e., the set of sequences for which the relative frequency of each $x \in \mathcal{X}$ differs from $P_{X^*}(x)$ by no more than $n^{-1/4}$. The following lemma, which is Lemma 8 of [11], guarantees that if the encoder is such that, with high probability $X^n \in \mathcal{T}_n$, then condition (ii) above is essentially satisfied:

Lemma 1 *[11, Lemma 8] Let X^n and Z^n be induced by an aribtrary encoder and the cascaded channel from X^n to Z^n:*

$$\frac{I(X^n; Z^n)}{n} \leq I(X^*; Z^*) + Pr\{X^n \in \mathcal{T}_n^c\} \cdot \log|\mathcal{X}| + f_1(n), \tag{41}$$

where $f_1(n) \to 0$ as $n \to \infty$.

Note that whenever $X^n \in \mathcal{T}_n$, the generalized power constraint is also essentially satisfied. It remains to handle conditions (i) and (iii). Consider next the encoder and the decoder of the
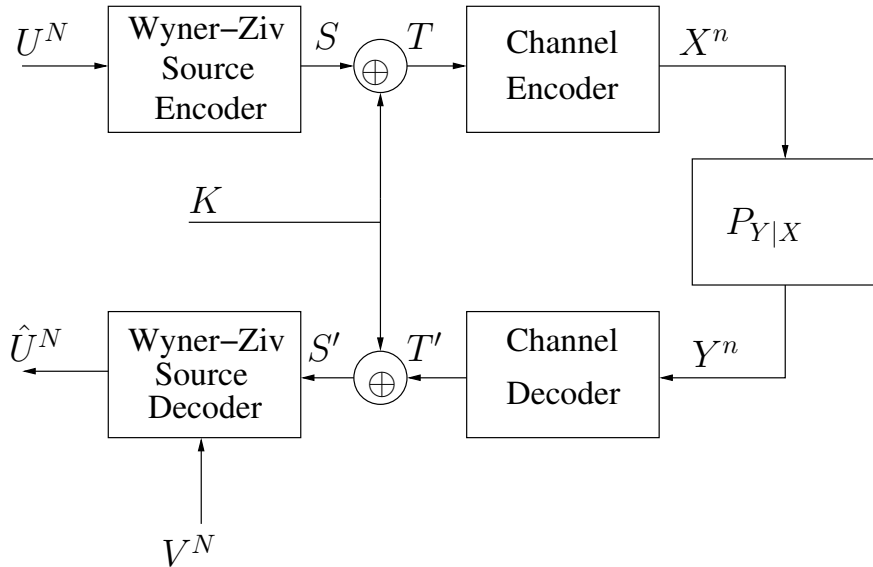
Figure 2: Encoder and decoder for the direct part.

legitimate receiver, depicted in Fig. 2. The source vector U^N is first compressed by a Wyner–Ziv encoder, designed for distortion level D and side information V^N, to a string of bits, $S = F_E(U^N)$, whose length does not exceed $N[R_{U|V}(D) + \epsilon/2] \leq n[I(X^*; Y^*) - \epsilon/(2\lambda)]$. Now, let us select R in the range

$$0 \leq R < R_{U|V}(D) - \lambda[I(X^*; Y^*) - I(X^*; Z^*)] - \epsilon, \tag{42}$$

where the right–most side is positive due to (40). The key K is a string of NR purely random bits, which are XORed with (the first) NR bits of S (one time pad). The resulting (partially) encrypted bit string, T, which will be represented by $T = S \oplus K = F_E(U^N) \oplus K$ (although it is possible that only some of the bits of S are XORed with those of K), is the message to be conveyed across the channel. Now, let

$$q_t \triangleq \Pr\{T = t\}, \quad t = 1, 2, \ldots, M \triangleq 2^{N[R_{U|V}(D) + \epsilon/2]}. \tag{43}$$

Next, let $M_1 = M_2 M$, where M_2 is a positive integer to be specified in the sequel. Let $\{\boldsymbol{x}_m\}_{m=1}^{M_1}$ be a subset of $\mathcal{X}^n$, which can be viewed as a code for the channel $P_{Y|X}$ or $P_{Z|X}$. The channel encoder and decoder in Fig. 2 work as follows. They both share a partition of $\{\boldsymbol{x}_m\}_{m=1}^{M_1}$ into M sub–codes, $\mathcal{C}_1, \mathcal{C}_2, \ldots, \mathcal{C}_M$, each of size M_2. Let $\mathcal{C}_t = \{\boldsymbol{x}_{(i-1)M_2+1}, \ldots, \boldsymbol{x}_{iM_2}\}$, $t = 1, 2, \ldots, M$. When $T = t$, the channel encoder outputs a vector X^n which is a (uniformly) randomly chosen member of sub–code

$\mathcal{C}_t$. Thus, for $t = 1, 2, \ldots, M$, $\tau = 1, 2, \ldots, M_2$,

$$\Pr\{X^n = \boldsymbol{x}_{(t-1)M_2+\tau} | T = t\} = \frac{1}{M_2} \tag{44}$$

and

$$\Pr\{X^n = \boldsymbol{x}_{(t-1)M_2+\tau}\} = \frac{q_t}{M_2}. \tag{45}$$

As mentioned earlier, the set $\{\boldsymbol{x}_m\}_{m=1}^{M_1}$ can be thought of as a code for the channel $P_{Y|X}$, where the prior probabilities of the codewords are given by (45). Let $T' = G(Y^n)$ denote the Bayes–optimal decoder for this code and these prior probabilities, which estimates the index t of the sub–code $\mathcal{C}_t$ that contains the transmitted codeword X^n. Let $\delta = \delta_Y(\boldsymbol{x}_1, \ldots, \boldsymbol{x}_{M_1}) \triangleq \Pr\{T' \neq T\}$. Obviously, if δ is small, namely, if $T' = T$ with high probability, then the Wyner–Ziv decoder $\hat{U}^N = F_D(S', V^N) = F_D(T' \oplus K, V^N)$ would output the "correct" reconstruction vector within distortion D, with the same probability.

Next, observe that each sub–code $\mathcal{C}_t$ may serve as a channel code for the degraded channel $P_{Z|X}$, provided that the corresponding decoder is informed of t. Let $\delta_t = \delta_Z(\mathcal{C}_t)$, $t = 1, 2, \ldots, M$, denote the error probability of code $\mathcal{C}_t$ w.r.t. the channel $P_{Z|X}$ when the decoder that observes Z^n is informed of t. Finally, let $\bar{\delta} = \sum_{t=1}^{M} q_t \delta_t$. With these definitions, we next make our first step to handle condition (iii).

Let U^N and K be such that $T = F_E(U^N) \oplus K = t$. Then, the channel input, given $T = t$, is distributed according to (44), that is, X^n is a randomly chosen member of $\mathcal{C}_t$, Thus, $H(X^n|T = t) = \log M_2$. Since δ_t is the probability of error associated with $\mathcal{C}_t$, Fano's inequality yields:

$$H(X^n|Z^n, T = t) \leq h(\delta_t) + \delta_t \log M_2 \leq 1 + \delta_t \log M_2 \tag{46}$$

where $h(\cdot)$ is the binary entropy function $h(a) = -a \log a - (1-a) \log(1-a)$. It follows then that

$$I(X^n; Z^n | T = t) \geq (1 - \delta_t) \log M_2 - 1, \tag{47}$$

which upon averaging over $\{t\}$ with weights $\{q_t\}$, yields

$$I(X^n; Z^n | T) \geq (1 - \bar{\delta}) \log M_2 - 1. \tag{48}$$

On the other hand,

$$
\begin{aligned}
I(X^n; Z^n | T) &= H(Z^n | T) - H(Z^n | X^n, T) \\
&= H(Z^n | T, U^N, W^N, K) - H(Z^n | X^n, T, U^N, W^N, K) \\
&\leq H(Z^n | U^N, W^N, K) - H(Z^n | X^n, U^N, W^N, K) \\
&= I(X^n; Z^n | U^N, W^N, K) \quad\quad\quad\quad\quad\quad\quad\quad\quad (49)
\end{aligned}
$$

where the second equality is due to the Markov relation $(U^N, W^N, K) \ominus T \ominus X^n \ominus Z^n$. Thus, we have established the inequality

$$
I(X^n; Z^n | U^N, W^N, K) \geq (1 - \bar{\delta}) \log M_2 - 1. \quad\quad (50)
$$

In the sequel, we will choose M_2 so as to meet condition (iii).

We next move on to handle condition (i). For every $S = s$, let $\mathcal{C}'_s$ denote the union of all 2^{NR} codebooks $\{\mathcal{C}_t, t = s \oplus k\}_{k=0}^{2^{NR}-1}$, and let $\delta'_s = \delta_Z(\mathcal{C}'_s)$ denote the error probability of $\mathcal{C}'_s$ w.r.t. the channel $P_{Z|X}$ when the decoder is informed of s. Let

$$
p_s \triangleq \Pr\{S = s\}, \quad s = 1, 2, \ldots, M. \quad\quad (51)
$$

Finally, let $\bar{\delta}' = \sum_s p_s \delta'_s$. With these definitions, let us now derive an upper bound on $H(K | U^N, W^N, Z^n)$:

$$
\begin{aligned}
H(K | U^N, W^N, Z^n) &\leq H(K | U^N, F_E(U^N), Z^n) \\
&= H(K | U^N, S, Z^n) \\
&\leq H(S \oplus T | S, Z^n) \\
&= H(T | S, Z^n) \\
&= \sum_{s=1}^{M} p_s H(T | S = s, Z^n) \\
&\leq \sum_{s=1}^{M} p_s [h(\delta'_s) + \delta'_s \log(2^{NR} M_2)] \\
&\leq 1 + \bar{\delta}'(NR + \log M_2), \quad\quad\quad\quad (52)
\end{aligned}
$$

where the third inequality is again Fano's inequality, and where we have also used the fact that δ'_s is an upper bound of the probability of error in estimating T, since T is only the index t of the codebook $\mathcal{C}_t$ to which the estimated codeword belongs.

To summarize our findings thus far, we substitute eqs. (52), (41) and (50) into (37), divide by N, and get:

$$
\begin{aligned}
\Delta \;\geq\; & H(U|W) + R + (1-\bar{\delta})\cdot\frac{\log M_2}{N} - \frac{2}{N} - \bar{\delta}'\left(R + \frac{\log M_2}{N}\right) - \\
& \lambda[I(X^*;Z^*) + \Pr\{X^n \in \mathcal{T}_n^c\}\cdot\log|\mathcal{X}| + f_1(n)].
\end{aligned}
\tag{53}
$$

Now, let us select

$$
M_1 = 2^{n[I(X^*;Y^*)-\epsilon/(2\lambda)]}
\tag{54}
$$

and

$$
M_2 = \frac{M_1}{M} = 2^{n[I(X^*;Y^*)-R_{U|V}(D)/\lambda-\epsilon/\lambda]}.
\tag{55}
$$

Applying this to (53), we get

$$
\begin{aligned}
\Delta \;\geq\; & H(U|W) + R + (1-\bar{\delta})[\lambda I(X^*;Y^*) - R_{U|V}(D) - \epsilon] - \frac{2}{N} - \bar{\delta}'(R + \log|\mathcal{X}|) - \\
& \lambda[I(X^*;Z^*) + \Pr\{X^n \in \mathcal{T}_n^c\}\cdot\log|\mathcal{X}| + f_1(n)] \\
\;\geq\; & H(U|W) + R + \lambda[I(X^*;Y^*) - I(X^*;Z^*)] - R_{U|V}(D) - \\
& \left\{\epsilon + \frac{2}{N} + (\bar{\delta}+\bar{\delta}')(R + \log|\mathcal{X}|) + \lambda[\Pr\{X^n \in \mathcal{T}_n^c\}\cdot\log|\mathcal{X}| + f_1(n)]\right\} \\
\;=\; & H(U|W) + R + \lambda\Gamma\left(\frac{R_{U|V}(D)+\epsilon}{\lambda}, Q\right) - R_{U|V}(D) - \\
& \left\{\epsilon + \frac{2}{N} + (\bar{\delta}+\bar{\delta}')(R + \log|\mathcal{X}|) + \lambda[\Pr\{X^n \in \mathcal{T}_n^c\}\cdot\log|\mathcal{X}| + f_1(n)]\right\}.
\end{aligned}
\tag{56}
$$

Finally, to prove that the expected distortion of $\hat{U}^N$ relative to U^N is essentially D, and to prove that Δ essentially meets the upper bound $\Delta^*(\lambda, R, D, Q)$ (namely, that the last term on the right–most side of (56) is arbitrarily small for large N), we have to prove the existence of a code $\{\boldsymbol{x}_m\}_{m=1}^{M_1}$ for which δ, $\bar{\delta}$, $\bar{\delta}'$ and $\Pr\{X^n \in \mathcal{T}_n^c\}$ are all simultaneously arbitrarily small for large N.

To this end, let us define $\mu(\boldsymbol{x}) = 1\{\boldsymbol{x} \in \mathcal{T}_n^*\}$, and for a given code $\{\boldsymbol{x}_m\}_{m=1}^{M_1}$, let $\delta_Y^m(\boldsymbol{x}_1,\ldots,\boldsymbol{x}_{M_1})$ denote the error probability w.r.t. the channel $P_{Y|X}$ with prior probabilities $\{q_t\}$ as given in (45), when $\boldsymbol{x}_m$ is transmitted. Then,

$$
\delta = \sum_{s=0}^{M-1} p_s \sum_{k=0}^{2^{NR}-1} \frac{1}{2^{NR}} \sum_{m=(s\oplus k)M_2+1}^{(s\oplus k+1)M_2} \frac{1}{M_2} \delta_Y^m(\boldsymbol{x}_1,\ldots,\boldsymbol{x}_{M_1}).
$$

Further, let $\delta_t = \delta_Z(\mathcal{C}_t)$, $\delta'_s = \delta_Z(\mathcal{C}'_s)$, $\bar{\delta}$, and $\bar{\delta}'$ be defined as above. Then,

$$
\begin{aligned}
\Phi(\boldsymbol{x}_1, \ldots, \boldsymbol{x}_{M_1}) \quad &\triangleq \quad \Pr\{X^n \in \mathcal{T}_n^c\} + \delta + \bar{\delta} + \bar{\delta}' \\
&= \sum_{s=0}^{M-1} p_s \sum_{k=0}^{2^{NR}-1} \frac{1}{2^{NR}} \sum_{m=(s\oplus k)M_2+1}^{(s\oplus k+1)M_2} \frac{1}{M_2} [\mu(\boldsymbol{x}_m) + \delta_Y^m(\boldsymbol{x}_1, \ldots, \boldsymbol{x}_{M_1}) + \\
&\qquad \delta_Z(\mathcal{C}_{s\oplus k}) + \delta_Z(\mathcal{C}'_s)].
\end{aligned}
\tag{57}
$$

Now, suppose that $\{\boldsymbol{x}_m\}_{m=1}^{M_1}$ are selected at random, with each $\boldsymbol{x}_m$ chosen independently according to $P_{X^*}^n(x^n) = \prod_{j=1}^n P_{X^*}(x_j)$. To prove that there exists a sequence of codes for which $\Phi \to 0$ as $n \to \infty$, all we have to show is that $\boldsymbol{E}\Phi \to 0$. But

$$
\boldsymbol{E}\Phi = \boldsymbol{E}\mu(\boldsymbol{X}) + \boldsymbol{E}\delta_Y^m(\boldsymbol{X}_1, \ldots, \boldsymbol{X}_{M_1}) + \boldsymbol{E}\delta_Z(\mathcal{C}_{s\oplus k}) + \boldsymbol{E}\delta_Z(\mathcal{C}'_s),
$$

where the indices s, k, and $m \in \{(s \oplus k)M_2 + 1, \ldots, (s \oplus k + 1)M_2\}$ are now immaterial. The first term tends to zero by the weak law of large numbers. The second term tends to zero by the ordinary random channel coding argument as the rate of the code $\{\boldsymbol{x}_m\}_{m=1}^{M_1}$ is less than $I(X^*; Y^*)$ (cf. the choice of M_1 above). By the same token, the fourth term vanishes with n, as $\mathcal{C}'_s$ is a random code of size $2^{NR}M_2$, and so its rate (cf. (42)) is

$$
R/\lambda + I(X^*; Y^*) - R_{U|V}(D)/\lambda < I(X^*; Z^*) - \epsilon/\lambda,
$$

which means that it is reliable for the channel $P_{Z|X}$ on the average. A–fortiori, the third term decays with n as $\mathcal{C}_t$ is even a smaller random code. By a simple application of the Chebychev inequality, with probability of at least $\frac{2}{3}$, the random of selection of the code yields $\Phi(\boldsymbol{X}_1, \ldots, \boldsymbol{X}_{M_1}) \leq 3\boldsymbol{E}\Phi$, which is still vanishingly small. On the other hand, since the codeword components are selected i.i.d. under P_{X^*}, then by the weak law of large numbers, for every $\epsilon > 0$ and large enough n and M_1, we have, with probability that tends to unity, and in particular, larger than $\frac{1}{2}$ from some point on:

$$
\sum_{t=1}^{M} \sum_{\tau=1}^{M_2} \frac{q_t}{M_2} \cdot \frac{1}{n} \sum_{j=1}^n \phi([\boldsymbol{X}_{(t-1)M_2+\tau}](j)) \leq Q + \epsilon
\tag{58}
$$

where $[\boldsymbol{X}_{(t-1)M_2+\tau}](j)$ is the j–th component of the codeword $\boldsymbol{X}_{(t-1)M_2+\tau}$. Since $\frac{2}{3} + \frac{1}{2} > 1$, it follows then that there exist codes for which both $\Phi(\boldsymbol{x}_1, \ldots, \boldsymbol{x}_{M_1}) \leq 3\boldsymbol{E}\Phi$ (and hence all components of Φ must be small) and the power constraint (58) holds at the same time.

The maximum secrecy of $H(U|W)$ is, of course, approached by letting R be arbitrarily close to (but strictly smaller than) $R_{U|V}(D) - \lambda\Gamma(R_{U|V}(D)/\lambda)$.

Finally, for completeness, we give a sketchy description of how the proof of the direct part should be slighlty modified in the (simpler) case where eq. (40) does not hold, namely,

$$\frac{R_{U|V}(D)}{\lambda} \leq I(X^*; Y^*) - I(X^*; Z^*). \tag{59}$$

Note that in this case, the achievable upper bound on Δ, asserted in Theorem 1, becomes $H(U|W)$ even for $R = 0$, as the bracketed term therein is non–positive. In the case, we will not use the key at all, i.e., $R = 0$ and K is degenerate. Thus, (37) becomes now:

$$N\Delta \geq NH(U|W) + I(X^n; Z^n|U^N, W^N) - I(X^n; Z^n).$$

As before, $I(X^n; Z^n)$ is essentially upper bounded by $nI(X^*; Z^*)$ using Lemma 1, and so, we only have to deal with the term $I(X^n; Z^n|U^N, W^N)$ and show that it is essentially lower bounded by $nI(X^*; Z^*)$. To this end, let us re–define M_1 as

$$M_1 = 2^{NR_{U|V}(D)+n[I(X^*;Z^*)-\epsilon/(2\lambda)]},$$

and M as before, so,

$$M_2 = \frac{M_1}{M} = 2^{n[I(X^*;Z^*)-\epsilon/\lambda]}.$$

Now, since $NR_{U|V}(D) + nI(X^*; Z^*) < nI(X^*; Y^*)$ (cf. (59)), the full codeword $\boldsymbol{x}_m$ can be reliably decoded at the legitimate decoder, as before. Also, since each sub–code $\mathcal{C}_t$ is, again, of rate less than $I(X^*; Z^*)$, then it can be decoded reliably by the wiretapper, provided that s/he is informed of t, thus $I(X^n; Z^n|U^N, W^N)$ is again, essentially lower bounded by $\log M_2 = n[I(X^*; Z^*) - \epsilon/\lambda]$. This completes the proof of direct part of Theorem 1.

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Neri Merhav - Biography

Neri Merhav (S'86–M'87–SM'93–F'99) was born in Haifa, Israel, on March 16, 1957. He received the B.Sc., M.Sc., and D.Sc. degrees from the Technion, Israel Institute of Technology, in 1982, 1985, and 1988, respectively, all in electrical engineering.

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