Noisy-Channel Theorem: Forward Direction

With the tools from the previous section, we are ready to prove the forward direction of Shannon's noisy-channel coding theorem, which states that any rate strictly below the channel capacity is achievable:

Theorem: Shannon's noisy-channel coding theorem (forward direction)

For a discrete memoryless channel with capacity C, any rate $R < C$ is achievable. Concretely, for any $\varepsilon > 0$ and any rate $R < C$, for large enough n there exists a $(2^{n \cdot R}, n)$ code with **maximal error** $\lambda^{(n)} < \varepsilon$.

Proof

Given a channel $(\mathcal{X}, P_{Y|X}, \mathcal{Y})$ with capacity $C = \max P_X I(X;Y)$, let $R < C$ and $\epsilon > 0$. We will first show that for big enough n, a *randomly constructed* code with rate R has a low error probability. We will then argue that a low error probability on average of codes implies the existence of some *specific* code with low error probability.

Fix an input distribution P_X that maximizes $I(X;Y)$. For any n, construct a **$(2^{n \cdot R}, n)$-code** $\mathcal{C}$ by choosing a codebook at random according to P_X. That is, for every message $w \in [2^{n \cdot R}]$, sample n times from the distribution P_X, creating a codeword $\mathcal{C}(w) = (\mathcal{C}_1(w), \mathcal{C}_2(w), \ldots, \mathcal{C}_n(w))$ by concatenating the n independent samples $\mathcal{C}_i(w) \sim P_X$.

Since the channel is memoryless, if w is sent over the channel using $\mathcal{C}$, the output distribution Y^n is given by:

$$P_{Y^n|X^n}(y^n|\mathcal{C}(w)) = \prod_{i=1}^{n} P_{Y|X}(y_i \mid \mathcal{C}_i(w)).$$

What is the probability that the decoded message $\hat{w}$ is incorrect, i.e., not equal to w? This depends on the decoding method used by the receiver. The optimal decoding procedure is **maximum-likelihood decoding**, where the input message that is most likely with respect to $P_{X|Y}$ is selected as the decoding $\hat{w}$. However, it is hard to analyze the error probability for this decoding method. Instead, we will assume that the receiver applies **jointly typical decoding**, which has a slightly higher probability of decoding to the wrong message, but still small enough for our analysis. Jointly typical decoding works as follows: upon receiving an output y^n, the receiver looks for a *unique* message $\hat{w}$ such that the pair $(\mathcal{C}(\hat{w}), y^n)$ is jointly typical. If there exists no such message, or if it is not unique, the receiver declares a failure by decoding to $\hat{w} = 0$ (which is always wrong because $w \in [2^{n \cdot R}] = \{1, 2, \dots, 2^{n \cdot R}\}$).

With this decoding procedure in mind, we analyze the average error probability $P[\texttt{error}]$, where the average is taken over both the randomly constructed code $\mathcal{C}$ and the uniformly randomly selected message w. Defining $\lambda_w(\mathcal{C}) := P[\hat{w} \neq w \mid \mathcal{C}(w) \text{ was sent over the channel}]$ to be the probability that a message w (encoded using $\mathcal{C}$) is decoded incorrectly, we get:

$$\begin{aligned} P[\texttt{error}] &= \sum_{\mathcal{C}} P[\mathcal{C}] \cdot \left(\sum_{w=1}^{2^{n \cdot R}} \frac{1}{2^{n \cdot R}} \cdot \lambda_w(\mathcal{C}) \right) \\ &= \frac{1}{2^{n \cdot R}} \sum_{w=1}^{2^{n \cdot R}} \sum_{\mathcal{C}} P[\mathcal{C}] \cdot \lambda_w(\mathcal{C}). \end{aligned}$$

Since we average over all randomly constructed codes $\mathcal{C}$, and the codewords for all messages are sampled independently, the value $\sum_{\mathcal{C}} P[\mathcal{C}] \cdot \lambda_w(\mathcal{C})$ does not depend on the particular message w. Hence if we set, for example, $w_0 = 1$, then for all $w \in [2^{n \cdot R}]$,

$$\sum_{\mathcal{C}} P[\mathcal{C}] \lambda_w(\mathcal{C}) = \sum_{\mathcal{C}} P[\mathcal{C}] \lambda_{w_0}(\mathcal{C}).$$

This simplifies the calculation of $P[\texttt{error}]$ significantly:

$$\begin{aligned} P[\texttt{error}] &= \frac{1}{2^{n \cdot R}} \sum_{w=1}^{2^{n \cdot R}} \sum_{\mathcal{C}} P[\mathcal{C}] \cdot \lambda_{w_0}(\mathcal{C}) \\ &= \sum_{\mathcal{C}} P[\mathcal{C}] \cdot \lambda_{w_0}(\mathcal{C}). \end{aligned}$$

That is, the average probability of error is the probability (over the selection of the code $\mathcal{C}$, and over the randomness in the channel) that the message w_0 is decoded incorrectly. There are two possible reasons for an error in the decoding:

1. The output of the channel is not jointly typical with $\mathcal{C}(w_0)$. By the first item of the **joint AEP**, this probability approaches zero as n goes to infinity. Hence, for

big enough n, the probability of an error for this reason is smaller than ϵ.

2. There is some $w' \neq w_0$ such that the output of the channel is (also) jointly typical with $\mathcal{C}(w')$. Since $\mathcal{C}$ is a random code (and so $\mathcal{C}(w')$ is independent from the channel output y^n), by the third item of the joint AEP the probability that this occurs is at most

$$\sum_{w' \neq w_0} 2^{-n(I(X;Y)-3\epsilon)} = (2^{n \cdot R} - 1) 2^{-n(I(X;Y)-3\epsilon)}.$$

We can thus bound the average probability of error, using the union bound and the bounds in the above analysis, by

$$\begin{aligned} P[\texttt{error}] &\leq \epsilon + (2^{n \cdot R} - 1) 2^{-n(I(X;Y)-3\epsilon)} \\ &\leq \epsilon + 2^{n \cdot R} 2^{-n(I(X;Y)-3\epsilon)} \\ &= \epsilon + 2^{-n(I(X;Y)-R-3\epsilon)} \end{aligned}$$

As long as $R < I(X;Y)$, one can choose n large enough so that $P[\texttt{error}] \leq 2\epsilon$.

This analysis upper bounds the (expected) average error probability for a random code $\mathcal{C}$. However, if this expected probability is low, there must be some specific code $\mathcal{C}^*$ that also has low average error probability.

Finally, in $\mathcal{C}^*$, we aim to bound the *maximal* error probability, i.e., the probability of error for the worst message. We can do so by noting that at least half of the messages w has error probability $\lambda_w(\mathcal{C}^*) \leq 4\epsilon$: if not, then the total error probability of these messages would already exceed $2^{n \cdot R} \cdot 2\epsilon$, contradicting the upper bound of 2ϵ to the average error probability. Thus, we can construct a better code by discarding the worst half of the codewords, and using the remaining $2^{n \cdot R - 1}$ codewords to construct a new code, with rate

$$\frac{\log(2^{n \cdot R - 1})}{n} = \frac{n \cdot R - 1}{n} = R - \frac{1}{n}$$

and maximal probability of error $\lambda^{(n)} \leq 4\epsilon$.

In the above proof, we implicitly assumed that $2^{n \cdot R}$ is an integer. You can try to redo the proof for the case when it is not: construct $\mathcal{C}$ as a $(\lceil 2^{n \cdot R} \rceil, n)$ code, and verify that the average probability of error $P[\texttt{error}]$ is still sufficiently small. Also compute a lower bound on the rate of the final code.