

Information Theory and Coding - Prof. Emere Telatar

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1 Data compression

Definition 1.1 (Information). *Abstractly, **information** can be thought of as the resolution of uncertainty.*

Given an alphabet \mathcal{U} (e.g. $\mathcal{U} = \{a, \dots, z, A, \dots, Z, \dots\}$), we want to assign binary sequences to elements of \mathcal{U} , i.e.

$$\mathcal{C} : \mathcal{U} \rightarrow \{0, 1\}^* = \{\emptyset, 0, 1, 00, 01, \dots\}$$

For \mathcal{X} a set

$$\begin{aligned}\mathcal{X}^n &\equiv \{(x_0 \dots x_n), x_i \in \mathcal{X}\} \\ \mathcal{X}^* &\equiv \bigcup_{n \geq 0} \mathcal{X}^n\end{aligned}$$

Definition 1.2. A code \mathcal{C} is called **singular** if

$$\exists (u, v) \in \mathcal{U}^2, u \neq v \quad \text{s.t.} \quad C(u) = C(v)$$

Non singular code is defined as opposite

Definition 1.3. A code \mathcal{C} is called **uniquely decodable** if

$$\forall u_1, \dots, u_n, v_1, \dots, v_n \in \mathcal{U}^* \quad \text{s.t.} \quad u_1, \dots, u_n \neq v_1, \dots, v_n$$

we have

$$\mathcal{C}(u_1) \dots \mathcal{C}(u_n) \neq \mathcal{C}(v_1) \dots \mathcal{C}(v_n)$$

i.e., \mathcal{C} is non-singular

Definition 1.4. Suppose $\mathcal{C} : \mathcal{U} \rightarrow \{0, 1\}^*$ and $\mathcal{D} : \mathcal{V} \rightarrow \{0, 1\}^*$ we can define

$$\mathcal{C} \times \mathcal{D} : \mathcal{U} \times \mathcal{V} \rightarrow \{0, 1\}^* \quad \text{as} \quad (\mathcal{C} \times \mathcal{D})(u, v) \rightarrow \mathcal{C}(u)\mathcal{D}(v)$$

Definition 1.5. Given $\mathcal{C} : \mathcal{U} \rightarrow \{0, 1\}^*$, define

$$\mathcal{C}^* : \mathcal{U}^* \rightarrow \{0, 1\}^* \quad \text{as} \quad \mathcal{C}^*(u_1 \dots u_n) = \mathcal{C}(u_1) \dots \mathcal{C}(u_n)$$

Definition 1.6. A code $\mathcal{U} \rightarrow \{0, 1\}^*$ is **prefix-free** if for no $u \neq v$ $\mathcal{C}(u)$ is a prefix of $\mathcal{C}(v)$.

Theorem 1.1. If \mathcal{C} is prefix-free then \mathcal{C} is uniquely decodable.

Definition 1.7. $l(\mathcal{C}(u))$ is the length of the code word $\mathcal{C}(u)$ and $l(\mathcal{C})$ is the expected length of the code:

$$l(\mathcal{C}) = \sum_u l(\mathcal{C}(u))p(u)$$

Definition 1.8 (Kraft sum). Given $\mathcal{C} : \mathcal{U} \rightarrow \{0, 1\}^*$

$$\text{kraftsum}(\mathcal{C}) = \sum_u 2^{-l(\mathcal{C}(u))}$$

Lemma 1.2. if $\mathcal{C} : \mathcal{U} \rightarrow \{0, 1\}^*$ and $\mathcal{D} : \mathcal{V} \rightarrow \{0, 1\}^*$ then

$$\text{kraftsum}(\mathcal{C} \times \mathcal{D}) = \text{kraftsum}(\mathcal{C}) \times \text{kraftsum}(\mathcal{D})$$

Proof.

$$\begin{aligned} \text{kraftsum}(\mathcal{C} \times \mathcal{D}) &= \sum_{u,v} 2^{-(l(\mathcal{C})+l(\mathcal{D}))} \\ &= \sum_u 2^{-l(\mathcal{C})} \sum_v 2^{-l(\mathcal{D})} \end{aligned}$$

□

Corollary 1.2.1. $\text{kraftsum}(\mathcal{C}^n) = (\text{kraftsum}(\mathcal{C}))^n$

Proposition 1.1. if \mathcal{C} is non-singular, then

$$\text{kraftsum}(\mathcal{C}) \leq 1 + \max_u l(\mathcal{C}(u))$$

In coding theory, the **Kraft-McMillan inequality** gives a necessary and sufficient condition for the existence of a uniquely decodable code for a given set of codeword lengths.

Wikipedia. Kraft's inequality limits the lengths of codewords in a prefix code: if one takes an exponential of the length of each valid codeword, the resulting set of values must look like a probability mass function, that is, it must have total measure less than or equal to one. Kraft's inequality can be thought of in terms of a constrained budget to be spent on codewords, with shorter codewords being more expensive.

Theorem 1.3. if \mathcal{C} is uniquely decodable, then $\text{kraftsum}(\mathcal{C}) \leq 1$

Proof. \mathcal{C} is uniquely decodable $\equiv \mathcal{C}^*$ is non singular

$$\begin{aligned} &\Rightarrow \text{kraftsum}(\mathcal{C}^n) \leq 1 + \max_{u_1, \dots, u_n} l(\mathcal{C}^n) \\ &\Rightarrow \text{kraftsum}(\mathcal{C})^n \leq 1 + nL, \quad L = \max l(\mathcal{C}(n)) \end{aligned}$$

A growing exp cannot be bounded by a linear function

$$\Rightarrow \text{kraftsum}(\mathcal{C}) \leq 1$$

□

TODO: something fuzzy with the following theorem

Theorem 1.4. Suppose $\mathcal{C} : \mathcal{U} \rightarrow \mathcal{N}$ is such that $\sum_u i^{\mathcal{C}(u)} \leq 1$, then, there exists a prefix-free code $\mathcal{C}' : \mathcal{U} \rightarrow \{0, 1\}$ s.t. $\forall u, l(\mathcal{C}(u)) = l(\mathcal{C}'(u))$

Proof. Let $\mathcal{U} = \{u_1, \dots, u_n\}$ and $\mathcal{C}(u_1) \leq \mathcal{C}(u_2) \leq \dots \leq \mathcal{C}(u_k) = \mathcal{C}_{max}$. Consider the complete binary tree up to depth \mathcal{C}_{max} initially all nodes are available to be used as codewords. For $i = 1, 2, \dots, n$, place $\mathcal{C}(u_i)$ at an available node at level $\mathcal{C}(u_i)$ remove all descendant of $\mathcal{C}(u_i)$ from the available list.

Corollary 1.4.1. Suppose $\mathcal{C} : \mathcal{U} \rightarrow \{0, 1\}^*$ is uniquely decodable, then there exist an $\mathcal{C}' : \mathcal{U} \rightarrow \{0, 1\}^*$ which is prefix-free and $l(\mathcal{C}'(n)) = l(\mathcal{C}(n))$

□

Example 1. $\mathcal{U} = \{a, b, c, d\}$, $\mathcal{C} : \{0, 01, 011, 111\}$ and $\mathcal{C}' : \{0, 10, 110, 111\}$

In this case, decoding \mathcal{C} may require delay, while decoding \mathcal{C}' is instantaneous.

2 Alphabet with statistics

Suppose we have an alphabet \mathcal{U} , and suppose we have a random variable U taking values in \mathcal{U} . We denote by $p(u) = \Pr(U = u)$, $u \in \mathcal{U}$ with $p(u) \geq 0$ and $\sum_u p(u) = 1$.

Suppose we have a code $\mathcal{C} : \mathcal{U} \rightarrow \{0, 1\}^*$. We then have $\mathcal{C}(u)$ a random binary string and $l(\mathcal{C}(u))$ a random integer.

Example 2. $\mathcal{U} = \{a, b, c, d\}$
 $p : \{0.5, 0.25, 0.125, 0.125\}$
 $\mathcal{C} : \{0, 01, 110, 111\}$

then we have

$$l(\mathcal{C}(u)) = \begin{cases} 1, & p = 0.5 \\ 2, & p = 0.25 \\ 3, & p = 0.125 + 0.125 + 0.25 \end{cases}$$

We can measure how efficient \mathcal{C} represents \mathcal{U} by considering

$$E[l(\mathcal{C}(u))] = \sum_u p(u) l(\mathcal{C}(u)) \quad \text{with} \quad \mathcal{C}(u) = l(\mathcal{C}(u))$$

Theorem 2.1. *if \mathcal{C} is uniquely decodable, then*

$$E[l(\mathcal{C}(u))] \geq \sum_u p(u) \log\left(\frac{1}{p(u)}\right)$$

Proof. let $\mathcal{C}(u) = l(\mathcal{C}(u))$, we know $\sum_u 2^{-\mathcal{C}(u)} \leq 1$ because \mathcal{C} is uniquely decodable

$$\begin{aligned} E[l(\mathcal{C}(u))] &= \sum_u p(u) \mathcal{C}(u) = \sum_u p(u) \log_2\left(\frac{1}{q(u)}\right) \\ &\equiv \sum_u p(u) \log\left(\frac{q(u)}{p(u)}\right) \leq 0 \\ &\equiv \sum_u p(u) \ln\left(\frac{q(u)}{p(u)}\right) \leq 0 \\ &\leq \sum_u p(u) \left[\frac{q(u)}{p(u)} - 1\right] = \underbrace{\sum_u q(u)}_{\leq 1} - \underbrace{\sum_u p(u)}_{=1} \leq 0 \end{aligned}$$

TODO: what's $q(u)$ in the proof above?

□

Theorem 2.2. *For any \mathcal{U} , there exists a prefix-free code \mathcal{C} s.t.*

$$E[l(\mathcal{C}(u))] < 1 + \sum_{u \in \mathcal{U}} p(u) \log\left(\frac{1}{p(u)}\right)$$

Proof. Given \mathcal{U} , let

$$\begin{aligned} \mathcal{C}(u) &= \log\left(\frac{1}{p(u)}\right) < 1 + \log\left(\frac{1}{p(u)}\right) \\ \Rightarrow \sum_u 2^{-\mathcal{C}(u)} &\leq \sum_u p(u) = 1 \\ \Rightarrow \sum_u p(u) \mathcal{C}(u) &< \sum_u p(u) \log\left(\frac{1}{p(u)}\right) + \underbrace{1}_{\sum p(u)} \end{aligned}$$

□

Definition 2.1 (Entropy). *Entropy quantifies the amount of uncertainty involved in the value of a random variable or the outcome of a random process.*

Theorem 2.3. *The entropy of a random variable $U \in \mathcal{U}$ is*

$$H(U) = \sum_{u \in \mathcal{U}} p(u) \log\left(\frac{1}{p(u)}\right)$$

with $p(u) = \Pr(U = u)$

Note that $H(U)$ is a function of the distribution $\mathcal{C}_u(\cdot)$ of the random variable U , it isn't a function of U .

$$H(U) = E[f(U)] \quad \text{where} \quad f(U) = \log\left(\frac{1}{p(u)}\right)$$

How to design optimal codes (in the sense of minimizing $E[l(\mathcal{C}(u))]$)?
Formally, given a random variable U , find $\mathcal{C}(u) \rightarrow \mathcal{N}$ s.t.

$$\sum_{u \in \mathcal{U}} 2^{\mathcal{C}(u)} \leq 1 \quad \text{that minimizes} \quad \sum_{u \in \mathcal{U}} p(u) \mathcal{C}(u)$$

Properties of optimal prefix-free codes

- if $p(u) < p(v)$ then $\mathcal{C}(u) \geq \mathcal{C}(v)$
- The two longest codewords have the same length
- There is an optimal code such that the two least probable letters are assigned codewords that differ in the last bit.

Observe that if $\{\mathcal{C}(u_1), \dots, \mathcal{C}(u_{k-1}), \mathcal{C}(u_k)\}$ is a prefix-free collection of the property that

$$\begin{aligned} \mathcal{C}(u_{k-1}) &= \alpha 0 \\ \mathcal{C}(u_k) &= \alpha 1 \end{aligned} \quad \text{with} \quad \alpha \in \{0, 1\}^*$$

then $\{\mathcal{C}(u_1), \dots, \mathcal{C}(u_{k-2}), \alpha\}$ is also a prefix-free collection.

Also

$$\begin{aligned} \sum_{u \in \mathcal{U}} p(u) l(\mathcal{C}(u)) &= p(u_1) l(\mathcal{C}(u_1)) + \dots + p(u_{k-2}) l(\mathcal{C}(u_{k-2})) + [p(u_{k-1}) + p(u_k)] (l(\alpha) + 1) \\ &= (p(u_{k-1}) + p(u_k)) + \sum_{v \in \mathcal{V}} p(v) l(\mathcal{C}'(v)) \end{aligned}$$

So we have shown that with

$$E[l(\mathcal{C}(U))] = p(u_{k-1}) + p(u_k) + E[l(\mathcal{C}'(V))]$$

if \mathcal{C} is optimal for U , then \mathcal{C}' is optimal for V

3 Entropy and mutual information

Definition 3.1 (Joint entropy). Suppose U, V are random variables with $p(u, v) = \Pr\{U = u, V = v\}$, the joint entropy is

$$H(UV) = \sum_{u,v} p(u, v) \log \frac{1}{p(u, v)}$$

Theorem 3.1.

$$H(UV) \leq H(U) + H(V)$$

with equality iff U and V are independants.

Proof. We want to show that

$$\sum_{u,v} p(u, v) \log \frac{1}{p(u, v)} \leq \sum_u p(u) \log \frac{1}{p(u)} + \sum_v p(v) \log \frac{1}{p(v)} \iff \sum_{u,v} p(u, v) \log \frac{p(u)p(v)}{p(u, v)} \leq 0$$

We use $\{\ln z \leq z - 1, \forall z\}$ (with equality iff $z = 1$):

$$\sum_{u,v} p(u, v) \log \frac{p(u)p(v)}{p(u, v)} \leq \sum_{u,v} p(u, v) \left[\frac{p(u)p(v)}{p(u, v)} - 1 \right] = \sum_{u,v} p(u)p(v) - \sum_{u,v} p(u, v) = 1 - 1 = 0$$

□

Same definitions of entropy holds for n symbols.

Definition 3.2 (Joint Entropy). Suppose U_1, U_2, \dots, U_n are RVs and we are given $p(u_1 \dots u_n)$, the joint entropy is

$$H(U_1, \dots, U_n) = \sum_{u_1 \dots u_n} p(u_1 \dots u_n) \log \frac{1}{p(u_1 \dots u_n)}$$

Theorem 3.2.

$$H(U_1 \dots U_n) \leq \sum_{i=1}^n H(U_i)$$

with equality iff U s are independants

Corollary 3.2.1. if U_1, \dots, U_n are i.i.d. then $H(U_1 \dots U_n) = nH(U_1)$

Definition 3.3 (Conditional entropy).

$$H(U|V) = \sum_{u,v} p(u, v) \log \frac{1}{p(u|v)}$$

Theorem 3.3.

$$H(UV) = H(U) + H(V|U) = H(V) + H(U|V)$$

Theorem 3.4.

$$H(U) + H(V) \geq H(UV) = H(V) + H(U|V)$$

Definition 3.4 (Mutual information). Mutual information measures the amount of information that can be obtained about one random variable by observing another.

$$\begin{aligned} I(U; V) &= I(V; U) = H(U) - H(U|V) \\ &= H(V) - H(V|U) \\ &= H(U) + H(V) - H(UV) \end{aligned}$$

We can apply the chain rule on the entropy as follow

$$H(U_1 U_2 \dots U_n) = H(U_1) + H(U_2|U_1) + \dots + H(U_n|U_1 U_2 \dots U_{n-1})$$

Definition 3.5 (Conditional mutual information).

$$\begin{aligned} I(U; V|W) &= H(U|W) - H(U|VW) \\ &= H(V|W) - H(V|UW) \\ &= \mathbb{E}_{u,v,w} \left[\log \frac{p(uv|w)}{p(u|w)p(v|w)} \right] \end{aligned}$$

Theorem 3.5.

$$I(V; U_1 \dots U_n) = I(V; U_1) + I(V; U_2|U_1) + \dots + I(V; U_n|U_1 \dots U_{n-1})$$

We can apply the chain rule on the mutual information as follows

$$I(U_1, U_2, \dots; V) = I(U_1; V) + I(U_2; V|U_1) + \dots$$

Notation 1.

$$U^n \triangleq (U_1 U_2 \dots U_n)$$

Theorem 3.6.

$$I(U; V|W) \geq 0$$

equality iff conditioned on w , u and v are independent, that is iff $U - V - W$ is a Markov chain.

Proof.

$$\begin{aligned} I(U; V|W) &= \frac{1}{\ln 2} \sum_{u,v,w} p(uvw) \ln \frac{p(u|w)p(v|w)}{p(uv|w)} \\ &\geq \frac{1}{\ln 2} \sum_{u,v,w} p(uvw) \left[\frac{p(u|w)p(v|w)}{p(uv|w)} - 1 \right] \\ &= \frac{1}{\ln 2} \sum_{u,v,w} (p(w)p(u|w)p(v|w) - p(uvw)) \\ &= \frac{1}{\ln 2} (1 - 1) \\ &= 0 \end{aligned}$$

□

4 Data processing

Theorem 4.1. $U - V - W$ is a MC $\iff I(U; W|V) = 0$

Corollary 4.1.1. $I(U; V) \geq I(U; W)$ and by symmetry of MC $I(W; V) \geq I(U; W)$

Proof.

$$I(U; VW) = I(U; V) + I(U; W|V) = I(U; V)$$

and

$$I(U; VW) = I(U; W) + I(U; V|W) \geq I(U; W)$$

□

Theorem 4.2. Given U a RV taking values in \mathcal{U} then $0 \leq H(U) \leq \log |\mathcal{U}|$. $H(U) = 0$ iff U is constant, $H(U) = \log |\mathcal{U}|$ iff U is $p(u) = 1/|\mathcal{U}|$ for all u .

Proof. For the lower bound,

$$H(U) = \sum_u \underbrace{p(u)}_{\geq 0} \underbrace{\log \frac{1}{p(u)}}_{\geq 0} \geq 0$$

For the upper bound,

$$\begin{aligned} H(U) - \log |\mathcal{U}| &= \sum_u p(u) \log \frac{1}{p(u)} - \sum_u p(u) \log |\mathcal{U}| \\ &= \frac{1}{\ln 2} \sum_u p(u) \ln \frac{1}{|\mathcal{U}| p(u)} \\ &\leq \frac{1}{\ln 2} \sum_u p(u) \left(\frac{1}{|\mathcal{U}| p(u)} - 1 \right) \\ &= \frac{1}{\ln 2} \left[\sum_u \frac{1}{|\mathcal{U}|} - \sum_u p(u) \right] \\ &= 0 \end{aligned}$$

□

Theorem 4.3. $I(U; V) = 0 \iff U \perp V$

Definition 4.1 (Entropy rate of a stochastic process).

$$r = \lim_{n \rightarrow \infty} \frac{1}{n} H(U^n) \quad \text{if the limit exists}$$

Theorem 4.4. For stationary stochastic process U^n , the sequences

$$a_n = \frac{1}{n} H(U^n) \text{ and } b_n = H(U_n | U^{n-1})$$

are positive and non increasing. Then $a = \lim_{n \rightarrow \infty} a_n$ and $b = \lim_{n \rightarrow \infty} b_n$ exists and $a = b$.

Proof.

$$\begin{aligned} b_{n+1} &= H(U_{n+1} | U_1, U_2, \dots, U_n) \\ &\leq H(U_{n+1} | U_2, \dots, U_n) \\ &= H(U_n | U_1, U_2, \dots, U_{n-1}) \\ &= b_n, \text{ because } U_1 \dots U_n \sim U_2 \dots U_{n+1} \text{ (Stationarity).} \end{aligned}$$

Hence, it is non-increasing.

For the $\{a_n\}$, observe that

$$\begin{aligned} a_n &= \frac{1}{n}H(U^n) = \frac{1}{n} \left[H(U_1) + H(U_2|U_1) + H(U_3|U^2) + \cdots + H(U_n|U^{n-1}) \right] \\ &= \frac{1}{n} \left[b_1 + b_2 + \cdots + b_n \right] \end{aligned}$$

and by the "Lemma", whenever $b_n \rightarrow b$, $a_n \rightarrow b$ □

Lemma 4.5 (Cesaro). *Suppose $b_n \rightarrow b$,*

then,

$$a_n = \frac{1}{n} \left[b_1 + b_2 + \cdots + b_n \right] \text{ also converges and to } b.$$

Proof. Since $b_n \rightarrow b$, $\left(\equiv \forall \epsilon > 0, \exists n(\epsilon) \text{ s.t. } \forall n > n(\epsilon) |b_n - b| < \epsilon \right)$

$\exists B$ s.t. $|b_n| < B$ for all n .

Take $n > n_1(\epsilon) \triangleq \dots$ then

$$\begin{aligned} |a_n - b| &\leq \frac{|b_1 - b| + |b_2 - b| + |b_3 - b| + \cdots + |b_n - b|}{n} \\ \text{so } |a_n - b| &\leq \frac{1}{n} \left[\sum_{i=1}^{n_0(\epsilon)} \underbrace{|b_i - b|}_{2B} + \sum_{i=n_0(\epsilon)+1}^n \underbrace{|b_i - b|}_{\leq \epsilon} \right] \leq \frac{n_0(\epsilon)2B}{n} + \epsilon < 2\epsilon \\ &\text{for } n > n_1(\epsilon) \triangleq \max, \left\{ n_0(\epsilon) \frac{1}{\epsilon} n_0(\epsilon) 2B \right\} \end{aligned}$$

□

Theorem 4.6. *Given a stationary process with entropy rate r :*

$$r = \lim_{n \rightarrow \infty} \frac{1}{n} H(U^n)$$

then

1. *for every source coding scheme*

$$\mathcal{C}_n : U^n \rightarrow \{0, 1\}^*$$

the expected number of bits / letter is given by

$$\frac{1}{n} E[l(\mathcal{C}(U^n))] \geq r$$

2. *for any $\epsilon > 0$, there exists a source coding scheme $\mathcal{C}_n : U^n \rightarrow \{0, 1\}^*$ s.t.*

$$\frac{1}{n} E[l(\mathcal{C}_n(U^n))] < r + \epsilon$$

Proof. 1. we already know

$$\frac{1}{n}E[l(\mathcal{C}_n(U^n))] \geq \frac{1}{n}H(U^n)$$

and the right term is decreasing

2. we also know that for each $n, \exists \mathcal{C}_n$ that is prefix-free s.t.

$$E[l(\mathcal{C}_n(U^n))] < \underbrace{\frac{1}{n}H(U^n)}_r + \underbrace{\frac{1}{n}}_0$$

we can find n large enough s.t. the right hand side $< r + \epsilon$

□

5 Typicality and typical set

Suppose we have a sequence U_1, U_2, \dots of i.i.d. random variables taking values in an alphabet \mathcal{U} . Suppose we observe u_1, u_2, \dots, u_n . We will call it to be *typical* (ϵ, p) if

$$p(u)(1 - \epsilon) \leq \frac{\# \text{ of times } u \text{ appears in } u_1, \dots, u_n}{n} \leq p(u)(1 + \epsilon)$$

Theorem 5.1. u^n is (ϵ, p) -typical then

$$2^{-nH(u)(1+\epsilon)} \leq Pr(U^n = u^n) \leq 2^{-nH(u)(1-\epsilon)}$$

Proof.

$$Pr(U^n = u^n) = \prod_{i=1}^n Pr(U_i = u_i) = \prod_{i=1}^n p(u_i) = \prod_{u \in \mathcal{U}} p(u)^{\#_u}$$

with $\#_u$ the number of times u appears in u_1, \dots, u_n where

$$n(1 - \epsilon)p(u) \leq \#_u \leq n(1 + \epsilon)p(u)$$

consequently

$$p(u)^{np(u)(1-\epsilon)} \geq p(u)^{\#_u} \geq p(u)^{np(u)(1+\epsilon)}$$

then

$$\left(\prod_n p(u)^{p(u)}\right)^{(1-\epsilon)n} \geq Pr(U^n = u^n) \geq \left(\prod_n p(u)^{p(u)}\right)^{(1+\epsilon)n}$$

but

$$p(u)^{p(u)} = 2^{-p(u) \log(\frac{1}{p(u)})} \Rightarrow \prod p(u)^{p(u)} = 2^{-H(u)}$$

□

Definition 5.1 (Typical set).

$$T(n, \epsilon, p) = \{u^n \in \mathcal{U}^n : u^n \text{ is } (\epsilon, p)\text{-typical}\}$$

Wikipedia. Typical sets provide a theoretical means for compressing data, allowing us to represent any sequence X^n using $nH(X)$ bits on average, and, hence, justifying the use of entropy as a measure of information from a source.

Theorem 5.2. 1. if $u^n \in T(n, \epsilon, p)$ then

$$p(u^n) = \Pr(U^n = u^n) = 2^{-nH(u)(1 \pm \epsilon)}$$

when U_i i.i.d.

2.

$$\lim_{n \rightarrow \infty} \Pr(U^n \in T(n, \epsilon, p)) = 1$$

3.

$$|T(n, \epsilon, p)| \leq 2^{n(H(u)(1+\epsilon))}$$

4.

$$|T(n, \epsilon, p)| \geq (1 - \epsilon)2^{nH(u)(1-\epsilon)}$$

Proof. 1. Fix $u \in \mathcal{U}$ let $X_i = 1$ if $U_i = u$ and 0 otherwise

$$\frac{\# \text{ of times } u \text{ appears in } U_1 \dots U_n}{n} = \frac{1}{n} \sum_{i=1}^n X_i$$

observe that $\{X_i\}$ are i.i.d.

$$X_i = \begin{cases} 1 & \text{w.p. } p(u) \\ 0 & \text{w.p. } 1 - p(u) \end{cases}$$

$$\Rightarrow E[X_i] = p(u) \quad \text{and} \quad \text{Var}[X_i] = p(u) - p(u)^2$$

$$\Pr \left\{ \underbrace{\left| \frac{1}{n} \sum_{i=1}^n X_i - p(u) \right|}_{u^n \text{ fails the test for letter } u} \geq \epsilon p(u) \right\} \leq \frac{\text{Var}(\frac{1}{n} \sum X_i)}{(\epsilon p(u))^2} = \frac{(1 - p(u))}{\epsilon^2 p(u)}$$

2.

$$\begin{aligned} \Pr \{U^n \notin T(n, \epsilon, p)\} &= \Pr \left\{ \bigcup_{u \in \mathcal{U}} \{u^n \text{ fails the test for } u\} \right\} \\ &\leq \sum_{u \in \mathcal{U}} \Pr \{U^n \text{ fails the test for } u\} \\ &\leq \frac{1}{n} \sum_{u \in \mathcal{U}} \frac{(1 - p(u))}{p(u)\epsilon} \quad \text{which goes to 0 as } n \text{ gets large} \end{aligned}$$

3.

$$\begin{aligned} 1 &\geq \Pr \{U^n \in T(n, \epsilon, p)\} = \sum_{u^n \in T(n, \epsilon, p)} \Pr \{U^n = u^n\} \\ &\geq \sum_{u^n \in T(n, \epsilon, p)} 2^{-n(1+\epsilon)H(u)} \\ &= 2^{n(1+\epsilon)H(u)} |T(n, \epsilon, p)| \end{aligned}$$

4.

$$\begin{aligned}
1 - \epsilon &\leq \Pr \{U^n \in T(n, \epsilon, p)\} = \sum_{u^n \in T(n, \epsilon, p)} \Pr \{U^n = u^n\} \\
&\geq \sum_{u^n \in T(n, \epsilon, p)} 2^{nH(u)(1-\epsilon)} \\
&= 2^{nH(u)(1-\epsilon)} |T(n, \epsilon, p)|
\end{aligned}$$

□

Definition 5.2 (Kullback-Leiber divergence (information gain)). *If we compress data in a manner that assumes $q(u)$ is the distribution underlying some data, when, in reality, $p(u)$ is the correct distribution, the Kullback-Leiber divergence is the number of average additional bits per datum necessary for compression.*

Lemma 5.3. *if $U_1 \dots U_n$ are i.i.d. with distribution q and $u_1 \dots u_n$ is (ϵ, p) -typical, then*

$$\begin{aligned}
\Pr \{U^n = u^n\} &= \left[\prod_u q(u)^{p(u)} \right]^{n(1+\epsilon)} \\
&= 2^{-n(1+\epsilon)} \sum_u p(u) \log \frac{1}{q(u)}
\end{aligned}$$

We know that $\Pr \{U^n \in T(n, \epsilon, p)\} \rightarrow 1$ as $n \rightarrow \infty$ and

$$(1 - \epsilon) 2^{nH(U)(1-\epsilon)} \leq |T(n, \epsilon, p)| \leq 2^{nH(U)(1+\epsilon)}$$

Observation 5.1. *Suppose $U_1 \dots U_n$ are i.i.d. following q and $u^n \in T(n, \epsilon, p)$*

$$\left[\prod_u q(u)^{p(u)} \right]^{n(1+\epsilon)} \leq \Pr \{U^n = u^n\} \leq \left[\prod_u q(u)^{p(u)} \right]^{n(1-\epsilon)}$$

and

$$\begin{aligned}
\prod_u q(u)^{p(u)} &= 2^{-\sum p(u) \log \frac{1}{q(u)}} \\
\sum_u p(u) \log \frac{1}{q(u)} &= \underbrace{\sum_u p(u) \log \frac{1}{p(u)}}_{H(p)} + \underbrace{\sum_u p(u) \log \frac{p(u)}{q(u)}}_{D(p||q)}
\end{aligned}$$

Corollary 5.3.1. *if $U_1 \dots U_n$ are i.i.d. following distribution q , then*

$$2^{-n[(1+\epsilon)D(p||q)+2\epsilon H(p)]} \leq \Pr \{U^n \in T(n, \epsilon, p)\} \leq 2^{-n[(1-\epsilon)D(p||q)-2\epsilon H(p)]}$$

Proof.

$$\Pr \{U^n \in T(n, \epsilon, p)\} = \sum_{u^n \in T(n, \epsilon, p)} \Pr \{U^n = u^n\}$$

We have

$$\begin{aligned}
2^{-n[H(p)+D(p||q)](1+\epsilon)} &\leq \Pr \{U^n = u^n\} \leq 2^{-n[H(p)+D(p||q)](1-\epsilon)} \\
2^{nH(p)(1-\epsilon)} &\leq |T(n, \epsilon, p)| \leq 2^{nH(p)(1+\epsilon)}
\end{aligned}$$

□

Observation 5.2.

$$D(p||q) = \sum_u p(u) \log \left(\frac{p(u)}{q(u)} \right) \geq 0 \text{ with equality iff } p = q$$

Example 3. $U \in \{0, 1\}$, $p = \frac{1}{2}, \frac{1}{2}$, $q = \frac{1}{2} - \delta, \frac{1}{2} + \delta$

$$D(p||q) = \frac{1}{2} \log \frac{1}{1-2\delta} + \frac{1}{2} \log \frac{1}{1+2\delta} = \frac{1}{2} \log \frac{1}{1-4\delta^2} = -\frac{1}{2} \log(1-4\delta^2) \approx \frac{1}{2} 4\delta^2 + o(\delta^4)$$

So if we want $2^{-nD(p||q)}$ small $n = \Omega(1/\delta^2)$

Example 4. Suppose we are told that U is p distributed and $p(u)$ are powers of 2. We design a prefix-free code \mathcal{C} to minimize $\sum_u p(u)l(\mathcal{C}(u))$. We have been misinformed and $U \sim q$, then:

$$\begin{aligned} E[l(\mathcal{C}(u))] &= \sum_u q(u) \log \frac{1}{p(u)} \\ &= \underbrace{H(q)}_{\text{length for optimal code}} + \underbrace{D(q||p)}_{\text{penalty for misbelief}} \end{aligned}$$

5.1 Universal data compression

Suppose we know that the distribution p of U is either $p_1, p_2 \dots p_k$, can we design a code $\mathcal{C} : U \rightarrow \{0, 1\}^*$

$$E[l(\mathcal{C}(U))] \leq H(U) + \text{small for every } p$$

$$E \left[\frac{1}{n} l(\mathcal{C}(U)) \right] \leq o(n) + E \left[h_2 \left(\frac{K}{n} \right) \right]$$

with $K = \sum_{i=1}^n u_i$

We have $\frac{E[K]}{n} = \theta_1$ and $E \left[h_2 \left(\frac{K}{n} \right) \right] \leq h_2 \left(E \left[\frac{K}{n} \right] \right) = h_2(\theta)$

Design \mathcal{C} Because the probability of a bit string is only dependant of the number of 1s (or 0s), it makes sense to encode two strings with the same numbers of 1 with code words of same lengths. Given $u_1 \dots u_n \in \{0, 1\}^n$, first count the number of 1, call it k .

$$\mathcal{C}(u_1 \dots u_n) = \underbrace{\text{describe } k}_{\lceil \log(n+1) \rceil} \underbrace{\text{describe } u_1 \dots u_n}_{\lceil \log \binom{n}{k} \rceil}$$

We now want to evaluate

$$\frac{1}{n} E[l(\mathcal{C}(U))]$$

when $U_1 \dots U_n$ are i.i.d with $p_1 = \theta$ and $p_0 = 1 - p_1$

Observation 5.3. for any $0 \leq \alpha \leq 1$

$$\begin{aligned} 1 = 1^n &= (\alpha + (1 - \alpha))^n = \sum_{i=0}^n \binom{n}{i} \alpha^i (1 - \alpha)^{n-i} \\ &\geq \binom{n}{k} \alpha^k (1 - \alpha)^{n-k} \end{aligned}$$

Then for all α

$$\binom{n}{k} \leq \alpha^{-k} (1 - \alpha)^{-(n-k)} = 2^{-n(\frac{k}{n} \log \frac{1}{\alpha} + (1 - \frac{k}{n}) \log \frac{1}{1-\alpha})}$$

We pick $\alpha = \frac{k}{n}$, and we get

$$\binom{n}{k} < 2^{nh_2(\frac{k}{n})}$$

Using this bound we have

$$\frac{1}{n}l(\mathcal{C}(u_1 \dots u_n)) \leq \frac{2}{n} + \frac{\log(n+1)}{n} + h_2\left(\frac{k}{n}\right)$$

$$E\left[\frac{1}{n}l(\mathcal{C}(U))\right] \leq o(n) + E\left[h_2\left(\frac{k}{n}\right)\right]$$

Claim 5.1. Suppose U_i are i.i.d. with $\Pr\{U_1 = 1\} = \theta$. We have $E\left[\frac{k}{n}\right] = \theta$ and $E\left[h_2\left(\frac{k}{n}\right)\right] \leq h_2(E\left[\frac{k}{n}\right]) = h_2(\theta)$. So

$$\lim_{n \rightarrow \infty} \frac{1}{n} E[l(\mathcal{C}(u_1 \dots u_n))] \leq h_2(\theta)$$

consequently this scheme is asymptotically optimal.

Proof. To prove the claim we need to show that if $\beta_1 \dots \beta_k$ are in $[0, 1]$ and $q_1 \dots q_k$ are non negative numbers that sum to 1 then

$$\sum_{i=1}^k q_i h_2(\beta_i) \leq h_2\left(\sum_{i=1}^k q_i \beta_i\right)$$

Let U and V be random variables with $U \in \{0, 1\}$ and $V \in \{1, \dots, k\}$ with

$$\begin{aligned} \Pr\{V = i\} &= q_i \\ \Pr\{U = 1|V = i\} &= \beta_i \\ \Pr\{U = 0|V = i\} &= 1 - \beta_i \end{aligned}$$

Then,

$$\begin{aligned} \Pr\{U = 1\} &= \sum_i q_i \beta_i \\ H(U) &= h_2\left(\sum_i q_i \beta_i\right) \\ H(U|V) &= \sum_i q_i h_2(\beta_i) \end{aligned}$$

And we already know that $H(U) \geq H(U|V)$ □

TODO: THOMAS SCRIBE ICI

Suppose we have an infinite string $u_1 u_2 \dots, u_i \in \mathcal{U}$, and

$$u_1 u_2 \dots = v_1 v_2 \dots \text{ with } v_i \in \mathcal{U}^*, v_i \neq v_j \text{ when } i \neq j$$

for any k we have

$$\lim_{m \rightarrow \infty} \frac{\text{length}(v_1 \dots v_m)}{m} \geq k \Rightarrow \lim_{m \rightarrow \infty} \frac{\text{length}(v_1 \dots v_m)}{m} = \infty$$

Definition 5.3. Given an infinite string $u_1 u_2 \dots$ and a machine M , let

$$\rho_M(u_1 u_2 \dots) = \lim_{n \rightarrow \infty} \frac{\text{length of the output } M \text{ after reading } u_1 u_2 \dots}{n}$$

also given $s > 0$, define

- The compressibility of \mathcal{U}^* by s -state machines

$$\rho_s(u_1 u_2 \dots) = \min_M \rho_M(u_1 u_2 \dots)$$

with M an s' -state machine with $s' \leq s$

- Compressibility of \mathcal{U}^* by finite state machines

$$\rho_{FSM}(u_1 u_2 \dots) = \lim_{s \rightarrow \infty} \rho_s(u_1 u_2 \dots)$$

Definition 5.4. Suppose $u_1 u_2 \dots$ an infinite sequence, define $m(n)$ as the largest m for which $u_1 \dots u_n = v_1 \dots v_m$ with distinct $v_1 \dots v_m$

Example 5.

$$u = aaaaaaaaaa, \quad \underbrace{\emptyset}_{v_1} \underbrace{a}_{v_2} \underbrace{aa}_{v_3} \underbrace{aaa}_{v_4} \underbrace{aaaa}_{v_5} \Rightarrow m(10) = 5$$

So far we know that

$$\frac{\text{length of the output of any } s\text{-state IL machine when it reads } u_1 u_2 \dots}{n} \geq \frac{m(n) \log(\frac{m(n)}{8s^2})}{n}$$

with

$$\frac{m(n) \log(\frac{m(n)}{8s^2})}{n} = \frac{m(n) \log(m(n))}{n} - \frac{m(n) \log(8s^2)}{\text{length}(v_1 \dots v_m)}$$

hence if M is a s -state machine

$$\rho_M(u_1 u_2 \dots) \geq \overline{\lim}_{n \rightarrow \infty} \frac{m(n) \log(m(n))}{n} \quad \text{then} \quad \rho_{FSM}(u_1 u_2 \dots) \geq \overline{\lim}_{n \rightarrow \infty} \frac{m(n) \log m(n)}{n}$$

6 Lemple-Ziv data compression method

Given some alphabet \mathcal{U} to both encoder and decoder, they also agree an order on \mathcal{U} :

1. Start with a dictionary $\mathcal{D} = \mathcal{U}$
2. To each word $w \in \mathcal{D}$, assign a $\lceil \log |\mathcal{D}| \rceil$ -bit binary description in the dictionary order
3. Parse the first word w in $u_1 u_2 \dots$ in the dictionary, output its binary description
4. replace w in \mathcal{D} by $\{wu, \forall u \in \mathcal{U}\}$.
5. Go to 2.

Example 6. Define an alphabet $\mathcal{U} = \{a, b, c\}$ with $a \leq b \leq c$ and an input message

$$u = abacac$$

- Create the dictionary $\mathcal{D} = \{a, b, c\}$ and its corresponding binary description $\mathcal{D}_{bin} = \{00, 01, 10\}$
- The first word in the message is $'a'$, output its binary description

$$output = 01$$

- Update the dictionary:

$$\mathcal{D} = \{a, ba, bb, bc, c\} \quad \mathcal{D}_{bin} = \{000, 001, 010, 011, 100\}$$

- Parse the next word $'ba'$ and output its binary description

$$output = 01001$$

- Update the dictionary

$$\mathcal{D} = \{a, baa, bab, bac, bb, bc, c\} \quad \mathcal{D}_{bin} = \{000, 001, \dots\}$$

- Continue until the end of the input data...

The decoder can proceed in a similar way to iteratively update the dictionary while decoding the message.

6.1 Analysis of LZ

Observe that LZ parses the string $u_1 u_2 \dots$ into $v_1 v_2 \dots$ with $v_i \in \mathcal{U}^*$ or $v_i \in \mathcal{D}_i$ where \mathcal{D}_i is the dictionary at step i .

When going from iteration $i \rightarrow i + 1$, v_i is removed from \mathcal{D} , consequently v_1, v_2, v_3 are distinct.

The length of the output of LZ after reading $u_1 \dots u_m$ is given by

$$\text{LZ output's length} = \lceil \log |\mathcal{U}| \rceil + \lceil \log(2|\mathcal{U}| - 1) \rceil + \lceil \log(3|\mathcal{U}| - 2) \rceil + \dots + \lceil \log(m|\mathcal{U}| - m + 1) \rceil$$

we observe that

$$\text{LZ output's length} < m(\log(m|\mathcal{U}|) + 1) = m \log(2m|\mathcal{U}|)$$

Also we have

$$\begin{aligned} \# \text{ bits / letter} &< \frac{m \log(2m|\mathcal{U}|)}{\text{length}(u_1 \dots u_m)} \\ &= \frac{m \log(m)}{\text{length}(u_1 \dots u_m)} + \frac{m \log(2|\mathcal{U}|)}{\text{length}(u_1 \dots u_m)} \end{aligned}$$

therefore

$$\rho_{LZ}(u_1 u_2 \dots) = \lim_{m \rightarrow \infty} \frac{\# \text{ bits}}{\text{letter}} \leq \lim_{m \rightarrow \infty} \frac{m \log(m)}{\text{length}(u_1 \dots u_m)} \leq \lim_{n \rightarrow \infty} \frac{m(n) \log(m(n))}{n} \leq \rho_{FSM}(u_1 u_2 \dots)$$

So we have proved the following theorem:

Theorem 6.1. for every $u_1 u_2 \dots$

$$\rho_{LZ}(u_1 u_2 \dots) \leq \rho_{FSM}(u_1 u_2 \dots)$$

Corollary 6.1.1. if $u_1 u_2 \dots$ is stationary

$$\rho_{LZ}(u_1 u_2 \dots) = \text{entropy rate of } u_1 u_2 \dots$$

7 Transmission of data

Interesting in the case of unreliable transmission media.

Definition 7.1 (Communication channel). A communication channel W is a device with an input alphabet \mathcal{X} and an output alphabet \mathcal{Y} . Its behavior is described by

$$W_i(y_i|x^i, y^{i-1}) = \Pr \{Y_i = y_i | X^i = x^i, Y^{i-1} = y^{i-1}\}$$

Definition 7.2 (Memoryless channel). a channel W is said to be memoryless if

$$W_i(y_i|x^i, y^{i-1}) = W(y_i|x_i)$$

Definition 7.3 (Stationary channel). a channel W is said to be stationary if

$$W_i(y|x) = W(y|x)$$

Example 7 (Binary erasure channel - BEC). $\mathcal{X} = \{0, 1\}$ and $\mathcal{Y} = \{0, 1, ?\}$, then

$$W(0|0) = 1 - p$$

$$W(?|0) = p$$

$$W(1|0) = 0$$

and same for $x_i = 1$.

Example 8 (Binary symmetric channel - BSC).

$$W(0|0) = 1 - p = W(1|1)$$

$$W(1|0) = p = W(0|1)$$

The input $X_1, X_2 \dots X_n$ to a channel might have memory

$$\Pr \{X^n = x^n\} = p(x_1)p(x_2|x_1) \dots p(x_i|x^{i-1}) \dots p(x_n|x^{n-1})$$

$$\begin{aligned} \Pr \{X^n = x^n, Y^n = y^n\} &= p(x_1)W_1(y_1|x_1)p(x_2|x_1, y_1)W(y_2|x_1, x_2, y_1) \dots \\ &= \prod_i p(x_i | \underbrace{x^{i-1}}_{\text{feedback}} \underbrace{y^{i-1}}_{\text{memory}}) W_i(y_i|x^i y^{i-1}) \end{aligned}$$

Lemma 7.1. if there is no feedback and the channel is memoryless and stationary, then

$$\Pr \{Y^n = y^n | X^n = x^n\} = \prod_{i=1}^n W(y_i|x_i)$$

Proof.

$$\begin{aligned} \Pr \{Y^n = y^n, X^n = x^n\} &= \prod_{i=1}^n p(x_i|x^{i-1} y^{i-1}) W_i(y_i|x^i y^{i-1}) \\ &= \prod_{i=1}^n p(x_i|x^{i-1}) W(y_i|x_i) \\ &= \prod_{i=1}^n W(y_i|x_i) \Pr \{X^n = x^n\} \end{aligned}$$

□

Example 9. Suppose W is BSC(1/2) but we have feedback, defined by $X_1 = 0$ and $X_i = Y_{i-1}$.

$$\begin{aligned} \Pr \{Y^2 = 00 | X^2 = 01\} &= 0 \\ W(0|0)W(0|1) &= \frac{1}{4} \end{aligned}$$

Lemma 7.2. *if W is stationary memoryless and there is no feedback, then*

$$H(Y^n | X^n) = \sum_{i=1}^n H(Y_i | X_i)$$

Proof.

$$H(Y^n | X^n) = E \left[\log \frac{1}{\Pr \{Y^n | X^n\}} \right] = E \left[\log \prod_{i=1}^n \frac{1}{\Pr \{Y_i | X_i\}} \right] = \sum_{i=1}^n E \left[\log \frac{1}{\Pr \{Y_i | X_i\}} \right] = \sum_{i=1}^n H(Y_i | X_i)$$

□

For a memoryless stationary channel $W(Y|X)$ we can compute, for any distribution $p(x)$, $p(x, y) = p(x)W(y|x)$ and $I(X; Y)$, we can also compute

$$C(W) = \max_{p(x)} I(X; Y)$$

TODO: is c a typo in the formula below? should it be n ?

Lemma 7.3. *for a stationary memoryless W without feedback, we have*

$$I(X^n; Y^n) \leq cC(W)$$

Proof.

$$\begin{aligned} I(X^n; Y^n) &= H(Y^n) - H(Y^n | X^n) \\ &= H(Y^n) - \sum_i H(Y_i | X_i) \\ &\leq \sum_i H(Y_i) - \sum_i H(Y_i | X_i) \\ &= \sum_i I(X_i; Y_i) \end{aligned}$$

Note that the joint distribution $\Pr \{X_i, Y_i\}$ is of the form $p(x)W(y|x)$, then $I(X_i; Y_i) \leq C(W)$

□

Notation 2. *for simplicity*

$$p * q = (1 - q)p + q(1 - p)$$

Example 10. Let W be a BSC(p), $\Pr \{X = 0\} = 1 - q$ and $\Pr \{X = 1\} = q$. Then

$$\begin{aligned} \Pr \{Y = 0\} &= (1 - q)(1 - p) + qp \\ \Pr \{Y = 1\} &= (1 - q)p + q(1 - p) \end{aligned}$$

$$\begin{aligned}
H(Y|X=0) &= p \log \frac{1}{p} + (1-p) \log \frac{1}{1-p} \\
H(Y|X=1) &= p \log \frac{1}{p} + (1-p) \log \frac{1}{1-p} \\
H(Y|X) &= p \log \frac{1}{p} + (1-p) \log \frac{1}{1-p}
\end{aligned}$$

$$\begin{aligned}
I(X;Y) &= H(Y) - H(Y|X) \\
&= (p * q) \log \frac{1}{p * q} + (1 - (p * q)) \log \frac{1}{1 - (p * q)} - \left[p \log \frac{1}{p} + (1-p) \log \frac{1}{1-p} \right]
\end{aligned}$$

We maximize $I(X;Y)$ for $q = 1/2$

$$C(W) = \log 2 - h_2(p)$$

Example 11. Let W be BEC(p) and $Pr\{X=1\} = q$

$$\begin{aligned}
H(X) &= h_2(q) \\
H(X|Y=0) &= 0 \\
H(X|Y=1) &= 0 \\
H(X|Y=?) &= h_2(q)
\end{aligned}$$

$$\begin{aligned}
I(X;Y) &= h_2(q) - p h_2(q) = (1-p) h_2(q) \\
C(W) &= (1-p) \log 2
\end{aligned}$$

7.1 Fano's inequality

Suppose U and V take values in the same alphabet \mathcal{U} , then

$$H(U|V) \leq p_e \log(|\mathcal{U}| - 1) + h_2(p_e)$$

with

$$p_e = Pr\{U \neq V\} \quad \text{and} \quad h_2(p) = p \log\left(\frac{1}{p}\right) + (1-p) \log\left(\frac{1}{1-p}\right)$$

Proof. Define

$$Z = \begin{cases} 1 & U \neq V \\ 0 & U = V \end{cases}, \quad H(Z) = h_2(p_e)$$

$$\begin{aligned}
H(UZ|V) &= H(U|V) + H(Z|UV) \\
&= H(Z|V) + H(U|VZ) \\
&\leq H(Z) + H(U|VZ)
\end{aligned}$$

but

$$H(U|VZ) = \underbrace{H(U|V, Z=0)}_0 Pr\{Z=0\} + \underbrace{H(U|V, Z=1)}_{\leq \log(|\mathcal{U}|-1)} \underbrace{Pr\{Z=1\}}_{p_e}$$

□

So if $H(U|V) > \lambda \Rightarrow \exists f(\lambda) > 0, p_e > f(\lambda)$

Corollary 7.3.1. Suppose U^L, V^L are random sequences with common alphabet \mathcal{U} , define :

$$p_{e,i} = \Pr \{U_i \neq V_i\}, \quad \bar{p}_e = \frac{1}{L} \sum_{i=1}^L p_{e,i}$$

then

$$\frac{1}{L} H(U^L|V^L) \leq h_2(\bar{p}_e) + \bar{p}_e \log(|\mathcal{U}| - 1)$$

Proof.

$$\begin{aligned} \frac{1}{L} H(U^L|V^L) &= \frac{1}{L} \sum_{i=1}^L H(U_i|U^{i-1}V^L) \\ &\leq \frac{1}{L} \sum_{i=1}^L H(U_i|V_i) \\ &\leq \frac{1}{L} \sum_{i=1}^L (p_{e,i} \log(|\mathcal{U}| - 1) + h_2(p_{e,i})) \\ &= \bar{p}_e \log(|\mathcal{U}| - 1) + \frac{1}{L} \sum_{i=1}^L h_2(p_{e,i}) \\ &\leq \bar{p}_e \log(|\mathcal{U}| - 1) + h_2\left(\frac{1}{L} \sum_{i=1}^L p_{e,i}\right) \\ &= \bar{p}_e \log(|\mathcal{U}| - 1) + h_2(\bar{p}_e) \end{aligned}$$

□

Theorem 7.4. "Bad news" theorem, converse to the coding theorem

- Suppose we have a stationary source $U_1U_2\ldots$ with entropy rate H and produces a letter every τ_s seconds.
- Suppose also that we have a channel W that accepts input $X_1X_2\ldots$ once every τ_c seconds.
- Suppose also

$$\frac{H}{\tau_s} > \frac{C(W)}{\tau_c}$$

then there is a $\lambda > 0$ such that $\bar{p}_e > \lambda$

Definition 7.4. stable suppose the encoder works by taking blocks of L letters

$$(U_1 \dots U_L)(U_{L+1} \dots U_{2L}) \dots$$

and outputs

$$(X_1 \dots X_n)(X_{n+1} \dots U_{2n}) \dots$$

then the encoder is stable if

$$L\tau_s \geq n\tau_c$$

Proof. Recall that for a stationary source $\frac{H(U_1 \dots U_L)}{L}$ tends to H so

$$H(U_1 \dots U_L) \geq LH$$

We also have

$$I(U^2; V^2) \leq nC(W)$$

therefore, since $\frac{n}{L} \leq \frac{\tau_s}{\tau_c}$

$$\begin{aligned} H(U^2|V^2) &= \frac{1}{L}(H(U^2) - I(U^2; V^2)) \geq H - \frac{n}{L}C(W) \\ &\geq H - \frac{\tau_s}{\tau_c}C(W) \\ &= \tau_s\left(\frac{H}{\tau_s} - \frac{C(W)}{\tau_c}\right) \end{aligned}$$

The right hand side is

$$\epsilon(\tau_c, \tau_s, H, C) > 0$$

so for every stable encoder, decoder, we have

$$\bar{p}_e \log(|\mathcal{U}| - 1) + h_2(\bar{p}_e) > \epsilon(\tau_s, \tau_c, H, C)$$

then

$$\bar{p}_e \geq \epsilon(\tau_s, \tau_c, H, C, |\mathcal{U}|)$$

□

Example 12. Suppose $\mathcal{U} = \{0, 1\}$ and $U_1 U_2 \dots$ is a Markov process with

$$U_1 = \begin{cases} 0 & \text{with } p = 0.5 \\ 1 & \text{with } p = 0.5 \end{cases}, \quad p(U_{n+1}|U_n) = \begin{cases} 1-p & u_{n+1} = u_n \\ p & u_{n+1} \neq u_n \end{cases},$$

$$\begin{aligned} H &= \lim_{n \rightarrow \infty} H(U_n|U^{n-1}) \\ &= \lim_{n \rightarrow \infty} H(U_n|U_{n-1}) \\ &= H(U_2|U_1) = h_2(p) \end{aligned}$$

suppose $w = BEC(q)$, $c(w) = (1-q)\log(2)$ and $\tau_s = \tau_c = 1$

$$h_2(\bar{p}_e) \geq h_2(p) - (1-q)\log(2) \Rightarrow \bar{p}_e \geq \lambda$$

What we want to do next is to show a matching "Good news" theorem:

We could show that if $\frac{H}{\tau_s} \leq \frac{c(w)}{\tau_c}$ then for any $\lambda > 0$, we can find a stable encoder and decoder such that $p_e < \lambda$. Instead, we will show stronger results:

1. **Separation theorem** The encoder can be designed in a modular way:

- A **source encoder** which encodes message words in bits. The design of this encoder is strongly dependent of the type of the input.
- A **channel encoder** which encodes the bits to maximize the performance with a specific channel.

2. We will show that

$$Pr \{U^L \neq V^L\} < \lambda$$

using the fact that

$$(U_i \neq V_i) \Rightarrow (U^L \neq V^L) \quad \text{so} \quad p_{e,i} \leq Pr \{U^L \neq V^L\} \Rightarrow \bar{p}_e \leq Pr \{U^L \neq V^L\}$$

We will now show that good channel encoders and channel decoders exist

Definition 7.5. *Given a channel W with input alphabet \mathcal{X} , a block encoder is a function*

$$Enc : \{1, \dots, M\} \rightarrow \mathcal{X}^n$$

with n the block length.

$Enc(1), \dots, Enc(M)$ are each called codewords and M is equal to the number of codewords.

The rate of the code can be defined by

$$R = \frac{\log M}{n}$$

Definition 7.6. *Given a channel W with output alphabet \mathcal{Y} , a block decoder is a function*

$$Dec : \mathcal{Y}^n \rightarrow \{?, 1, \dots, M\}$$

Definition 7.7.

$$p_{error}(m) = Pr \{\hat{m} \neq m | m\}$$

$$\bar{p}_{error}(m) = \frac{1}{M} \sum_{m=1}^M p_{error}(m)$$

$$\hat{p}_{error}(m) = \max_m p_{error}(m)$$

Appendices

A Markov chains

$U_1 - U_2 - \dots - U_n$ forms a Markov chain if the joint probability distribution of the RVs is

$$p(a, b, c, d) = p(a)p(b|a)p(c|b)p(d|c)$$

which is equivalent to (U_1, \dots, U_{k-1}) are independent of (U_{k+1}, \dots, U_n) when conditioned on U_k for any k .

Theorem A.1. *The reverse of a MC is a MC*

B Stochastic processes

A stochastic process is a collection $U_1, U_2 \dots U_n$ of RVs each taking values in \mathcal{U} . It is described by its joint probability

$$p(u^n) = P(U_1 \dots U_n = u_1 \dots u_n) = P(U^n = u^n)$$

Definition B.1 (Stationary stochastic process). *A process U_1, U_2, \dots is called stationary if for every n and k and $u_1 \dots u_n$, we have*

$$p(u^n) = p(U_1 \dots U_n = u_1 \dots u_n) = p(U_{1+k} \dots U_{n+k} = u_1 \dots u_n)$$

In other words, the process is time shift invariant.