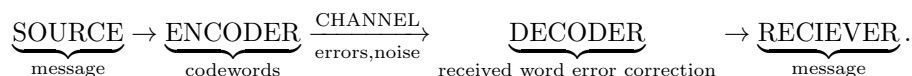


Introduction

We model communication:



Examples: optical signals, electrical telegraph, SMS (compression), postcodes, CDs (error correction), zip/gz files (compression).

Given a source and a channel, modelled probabilistically, the basic problem is to design an encoder and decoder to transmit messages economically (noiseless coding; compression) and reliably (noisy coding).

Examples:

- Noiseless coding: Morse code: common letters are assigned shorter code-words, e.g. $A \mapsto \bullet-$, $E \mapsto \bullet$, $Q \mapsto --\bullet-$, $S \mapsto \bullet\bullet\bullet$, $O \mapsto --$, $Z \mapsto --\bullet\bullet$. Noiseless coding is adapted to source.
- Noisy coding: Every book has an ISBN $a_1, a_2, \dots, a_9, a_{10}$, $a_i \in \{0, 1, \dots, 9\}$ for $1 \leq i \leq 9$ and $a_{10} \in \{0, 1, \dots, 9, X\}$ with $\sum_{j=1}^{10} ja_j \equiv 0 \pmod{11}$. This detects common errors - e.g one incorrect digit, transposition of two digits. Noisy coding is adapted to the channel.

Plan:

- (I) Noiseless coding - entropy
- (II) Error correcting codes - noisy channels
- (III) Information theory - Shannon's theorems
- (IV) Examples of codes
- (V) Cryptography

Books: [GP], [W], [CT], [TW], Buchmann, Körner. Online notes: Carne, Körner.

Basic Definitions

Definition (Communication channel). A *communication channel* accepts symbols from a alphabet $\mathcal{A} = \{a_1, \dots, a_r\}$ and it outputs symbols from alphabet $\mathcal{B} = \{b_1, \dots, b_s\}$. Channel modelled by the probabilities $\mathbb{P}(y_1 \dots y_n \text{ received} | x_1 \dots x_n \text{ sent})$. A *discrete memoryless channel* (DMC) is a channel with

$$p_{ij} = \mathbb{P}(b_j \text{ received} | a_i \text{ sent})$$

the same for each channel use and independent of all past and future uses. The channel matrix is $P = (b_{ij})$, a $r \times s$ stochastic matrix.

Definition (Binary symmetric channel). The *binary symmetric channel* (BSC) with error probability $p \in [0, 1)$ from $\mathcal{A} = \mathcal{B} = \{0, 1\}$. The channel matrix is

$$\begin{pmatrix} 1-p & p \\ p & 1-p \end{pmatrix}.$$

A symbol is transmitted correctly with probability $1 - p$. Usually assume $p < 1/2$.

The *binary erasure channel* (BEC) has $\mathcal{A} = \{0, 1\}$, $\mathcal{B} = \{0, 1, *\}$. The channel matrix is

$$\begin{pmatrix} 1-p & 0 & p \\ 0 & 1-p & p \end{pmatrix}.$$

So $p = \mathbb{P}(\text{symbol can't be read})$.

Definition. We model n uses of a channel by the n th extension, with input alphabet \mathcal{A}^n and output alphabet \mathcal{B}^n . A *code* C of length n is a function $\mathcal{M} \rightarrow \mathcal{A}^n$ where \mathcal{M} is the set of possible messages. Implicitly we also have a decoding rule $\mathcal{B}^n \rightarrow \mathcal{M}$. The *size* of C is $m = |\mathcal{M}|$. The *information rate* is $\rho(C) = \frac{1}{n} \log_2 m$. The *error rate* is $\hat{e}(C) = \max_{x \in \mathcal{M}} \mathbb{P}(\text{error} | x \text{ sent})$.

Remark. For the remainder of the course we write \log instead of \log_2 .

Definition. A channel can *transmit reliably at rate* R if there exists $(C_n)_{n=1}^\infty$ with each C_n a code of length n such that

$$\lim_{n \rightarrow \infty} \rho(C_n) = R \text{ \& } \lim_{n \rightarrow \infty} \hat{e}(C_n) = 0.$$

The *capacity* is the supremum of all reliable transmission rates. We'll see in Chapter 9 that a BSC with error probability $p < 1/2$ has non-zero capacity.

1 Noiseless coding

1.1 Prefix-free codes

For an alphabet \mathcal{A} , $|\mathcal{A}| < \infty$, let $\mathcal{A}^* = \bigcup_{n \geq 0} \mathcal{A}^n$, the set of all finite strings from \mathcal{A} . The *concatenation* of strings $x = x_1 \dots x_r$ and $y = y_1 \dots y_s$ is $xy = x_1 \dots x_r y_1 \dots y_s$.

Definition. Let \mathcal{A}, \mathcal{B} be alphabets. A code is a function $c : \mathcal{A} \rightarrow \mathcal{B}^*$. The strings $c(a)$ for $a \in \mathcal{A}$ are called *codewords* or *words* (CWS).

Example 1.1 (Greek fire code). $\mathcal{A} = \{\alpha, \beta, \dots, \omega\}$ (greek alphabet), $\mathcal{B} = \{1, 2, 3, 4, 5\}$, $c : \alpha \mapsto 11, \beta \mapsto 12, \dots, \psi \mapsto 53, \omega \mapsto 54$. xy means hold up x torches and another y torches nearby.

Example 1.2. \mathcal{A} = words in a dictionary, $\mathcal{B} = \{A, B, \dots, Z, \omega\}$. $c : \mathcal{A} \rightarrow \mathcal{B}$ splits the word and follows with a space. Send message $x_1 \dots x_n \in \mathcal{A}^*$ as $c(x_1) \dots c(x_n) \in \mathcal{B}^*$. So c extends to a function $c^* : \mathcal{A}^* \rightarrow \mathcal{B}^*$.

Definition. c is said to be *decipherable* if the induced map c^* (as in the previous example) is injective. In other words, each string from \mathcal{B} corresponds to at most one message.

Clearly if c is decipherable, it is necessary for c to be injective. However it is not sufficient:

Example 1.3. $\mathcal{A} = \{1, 2, 3, 4\}$, $\mathcal{B} = \{0, 1\}$. Define $c : 1 \mapsto 0, 2 \mapsto 1, 3 \mapsto 00, 4 \mapsto 01$. Then $c^*(114) = 0001 = c^*(312) = c^*(144)$ yet c is injective.

Notation: $|\mathcal{A}| = m$, $|\mathcal{B}| = a$, call c an a -ary code of size m . For example a 2-ary code is a binary one, and a 3-ary code is a ternary code.

Our aim is to construct decipherable codes with short word lengths. Assuming c is injective, the following codes are always decipherable:

- (i) A block code has all codewords of the same length (e.g Greek fire code);
- (ii) A comma code reserves a letter from \mathcal{B} to signal the end of a word (e.g Example 1.2);
- (iii) A prefix-free code is a code where no codeword is a prefix of any other distinct word (if $x, y \in \mathcal{B}^*$ then x is a prefix of y if $y = xz$ for some string $z \in \mathcal{B}^*$).

(i) and (ii) are special cases of (iii). As we can decode the message as it is received, prefix-free codes are sometimes called *instantaneous*.

Exercise: find a decipherable code which is not prefix-free.

Definition (Kraft's inequality). $|\mathcal{A}| = m$, $|\mathcal{B}| = a$, $c : \mathcal{A} \rightarrow \mathcal{B}^*$ has word lengths l_1, \dots, l_m . Then Kraft's inequality is

$$\sum_{i=1}^m a^{-l_i} \leq 1. \quad (*)$$

Theorem 1.1. A prefix-free code exists if and only if Kraft's inequality $(*)$ holds.

Proof. Rewrite $(*)$ as

$$\sum_{l=1}^s n_l a^{-l} \leq 1, \quad (**)$$

where n_l is the number of codewords with length l , and $s = \max_{1 \leq i \leq m} l_i$.

Now if $c : \mathcal{A} \rightarrow \mathcal{B}^*$ is prefix-free,

$$n_1 a^{s-1} + n_2 a^{s-2} + \dots + n_{s-1} a + n_s \leq a^s.$$

Indeed the LHS is the number of strings of length s in B with some codeword of c as a prefix, and the RHS is the total number of strings of length S . Dividing through by a^s we get (**).

Now given n_1, \dots, n_s satisfying (**), we try to construct a prefix-free code c with n_l codewords of length l , $\forall l \leq s$. Proceed by induction on s , $s = 1$ is clear (since (**) gives $n_1 \leq a$ so can construct code).

By the induction hypothesis, there exists a prefix-code \hat{c} with n_l codewords of length l for all $l \leq s - 1$. Then (**) implies

$$n_1 a^{s-1} + n_2 a^{s-2} + \dots + n_{s-1} a + n_s \leq a^s.$$

The first $s - 1$ terms on the LHS sum to the number of strings of length s with a codeword of \hat{c} as a prefix and the RHS is the number of strings of length s . Hence we can add at least n_s new codewords of length s to \hat{c} and maintain the prefix-free property. □

Remark. This proof is constructive: just choose codewords in order of increasing length, ensuring that no previous codeword is a prefix.

Theorem 1.2 (McMillan). *Any decipherable code satisfies Kraft's inequality.*

Proof (Karush, 1961). Let $c : \mathcal{A} \rightarrow \mathcal{B}^*$ be a decipherable code with word lengths l_1, \dots, l_m . Set $s = \max_{1 \leq i \leq m} l_i$. For $R \in \mathbb{N}$

$$\left(\sum_{i=1}^m a^{-l_i} \right)^R = \sum_{l=1}^{Rs} b_l a^{-l}, \quad (\dagger)$$

where b_l is the number of ways of choosing R codewords of total length l . Since c is decipherable, any string of length l formed from codewords must correspond to at most one sequence of codewords, i.e $b_l \leq |\mathcal{B}^l| = a^l$. Subbing this into (†)

$$\left(\sum_{i=1}^m a^{-l_i} \right)^R \leq \sum_{l=1}^{Rs} a^l a^{-l} = Rs,$$

so

$$\sum_{i=1}^m a^{-l_i} \leq (Rs)^{1/R} \rightarrow 1 \text{ as } R \rightarrow \infty.$$

Hence $\sum_{i=1}^m a^{-l_i} \leq 1$. □

Corollary 1.3. *A decipherable code with prescribed word lengths exists if and only if a prefix-free code with the same word lengths exists.*

Proof. Combine previous two theorems. □

Therefore we can restrict our attention to prefix-free codes.

2 Shannon's Noiseless Coding Theorem

Entropy is a measure of 'randomness' or 'uncertainty'. Suppose we have a random variable X taking a finite set of values x_1, \dots, x_n with probabilities p_1, \dots, p_n respectively. The *entropy* $H(X)$ of X is the expected number of fair coin tosses needed to simulate X (roughly speaking).

Example 2.1. Suppose $p_1 = p_2 = p_3 = p_4 = 1/4$. Identify (x_1, x_2, x_3, x_4) with (HH, HT, TH, TT) . Then the entropy is 2.

Example 2.2. Suppose $(p_1, p_2, p_3, p_4) = (1/2, 1/4, 1/8, 1/8)$. Identify (x_1, x_2, x_3, x_4) with (H, TH, TTH, TTT) . Then the entropy is

$$\frac{1}{2} \times 1 + \frac{1}{4} \times 2 + \frac{1}{8} \times 3 + \frac{1}{8} \times 3 = \frac{7}{4}.$$

In a sense, the previous example (2.1) was 'more random' than this.

Definition (Entropy). The *entropy* of X is

$$H(X) = - \sum_{i=1}^b p_i \log p_i.$$

(Recall that $\log =: \log_2$ here.) Note $H(X) \geq 0$. It is measured in *bits* (binary digits). Conventionally, we take $0 \log 0 = 0$.

Example 2.3. Take a biased coin $\mathbb{P}(H) = p, \mathbb{P}(T) = 1 - p$. Write $H(p, 1 - p) := H(p)$. Then

$$H(p) = -p \log p - (1 - p) \log(1 - p).$$

Note that $H'(p) = \log \frac{1-p}{p}$. Hence the entropy is maximised for $p = 1/2$ (giving entropy 1).

Proposition 2.1 (Gibbs' inequality). Let $(p_1, \dots, p_n), (q_1, \dots, q_n)$ be probability distributions. Then

$$- \sum_{i=1}^n p_i \log p_i \leq - \sum_{i=1}^n p_i \log q_i.$$

(The RHS is sometimes called the *cross entropy* or *mixed entropy*) Furthermore we have equality iff $p_i = q_i$ for all i .

Proof. Since $\log x = \frac{\ln x}{\ln 2}$, we may replace \log with \ln . Put $I = \{1 \leq i \leq n : p_i \neq 0\}$. Now $\ln x = x - 1$ for all $x > 0$ with equality iff $x = 1$. Hence $\ln \frac{q_i}{p_i} \leq \frac{q_i}{p_i} - 1$ for all $i \in I$. So

$$\begin{aligned} \sum_{i \in I} p_i \ln \frac{q_i}{p_i} &\leq \underbrace{\sum_{i \in I} q_i}_{\leq 1} - \underbrace{\sum_{i \in I} p_i}_{=1} \leq 0 \\ \implies - \sum_{i \in I} p_i \ln p_i &\leq - \sum_{i \in I} p_i \ln q_i \end{aligned}$$

$$\implies -\sum_{i=1}^n p_i \ln p_i \leq -\sum_{i=1}^n p_i \ln q_i.$$

If equality holds, then $\sum_{i \in I} q_i = 1$ and $\frac{p_i}{q_i} = 1$ for all $i \in I$. So $q_i = p_i$ for all $1 \leq i \leq n$. \square

Corollary 2.2. $H(p_1, p_2, \dots, p_n) \leq \log n$ with equality iff $p_1 = p_2 = \dots = p_n = 1/n$.

Proof. Take $q_1 = q_2 = \dots = q_n = 1/n$ in Gibbs' inequality. \square

Let $\mathcal{A} = \{\mu_1, \dots, \mu_m\}$, $|\mathcal{B}| = a$ ($m, n \geq 2$). The random variable X takes values μ_1, \dots, μ_m with probabilities p_1, \dots, p_m .

Definition. If $c : \mathcal{A} \rightarrow \mathcal{B}^*$ is a code, we say it is *optimal* if it has the smallest possible expected word length. i.e. $\mathbb{E}S := \sum_{i=1}^m p_i l_i$ is minimal amongst all decipherable codes.

Theorem 2.3 (Shannon's Noiseless Coding Theorem). *The expected word length $\mathbb{E}S$ of an optimal code satisfies*

$$\frac{H(X)}{\log a} \leq \mathbb{E}S < \frac{H(X)}{\log a} + 1.$$

Remark. The lower bound is actually true for any decipherable code.

Proof. We first get the lower bound. Let $c : \mathcal{A} \rightarrow \mathcal{B}^*$ be decipherable with word lengths l_1, \dots, l_m . Let $q_i = \frac{a^{-l_i}}{D}$ where $D = \sum_{i=1}^m a^{-l_i}$. Note $\sum_{i=1}^m q_i = 1$. By Gibbs' inequality

$$\begin{aligned} H(X) &\leq -\sum_{i=1}^m p_i \log q_i \\ &= -\sum_{i=1}^m p_i (-l_i \log a - \log D) \\ &= \left(\sum_{i=1}^m p_i l_i \right) \log a + \log D. \end{aligned}$$

By McMillan, $D \leq 1$ so $\log D \leq 0$. Hence

$$H(X) \leq \left(\sum_{i=1}^m p_i l_i \right) \log a \implies \frac{H(X)}{\log a} \leq \mathbb{E}S.$$

And we have equality iff $p_i = a^{-l_i}$ for some integers l_1, \dots, l_m . Note we have only used decipherability so far.

Now we get the upper bound. Take $l_i = \lceil -\log_a p_i \rceil$. Then

$$-\log_a p_i \leq l_i < -\log_a p_i + 1.$$

Hence $\log_a p_i \geq -l_i$, so $p_i \geq a^{-l_i}$. Therefore $\sum_{i=1}^m a^{-l_i} \leq \sum_{i=1}^m p_i = 1$. By Kraft's inequality, there exists a prefix-free code c with word lengths l_1, \dots, l_m . c has expected word length

$$\mathbb{E}S = \sum_{i=1}^m p_i l_i < \sum_{i=1}^m p_i (-\log_a p_i + 1) = \frac{H(X)}{\log a} + 1.$$

□

Example 2.4 (Shannon-Fano Coding). We mimic the above proof: given p_1, \dots, p_m , set $l_i = \lceil -\log_a p_i \rceil$. Construct a prefix-free code with word lengths l_i by choosing codewords in order of increasing length, ensuring any new codeword has no previous codeword as a prefix (Kraft's inequality ensures we can do this).

Example 2.5. Take $a = 2, m = 5$.

i	p_i	$\lceil -\log_2 p_i \rceil$	code
1	0.4	2	00
2	0.2	3	010
3	0.2	3	011
4	0.1	4	1000
5	0.1	4	1001

Then $\mathbb{E}S = \sum_{i=1}^m p_i l_i = 2.8$, $H = H/\log a = 2.12$. [See also Carne p13.]

3 Huffman Coding

How to construct an optimal code? Take $\mathcal{A} = \{\mu_1, \dots, \mu_m\}$, $p_i = \mathbb{P}(X = \mu_i)$. For simplicity take $|\mathcal{B}| = a = 2$. Without loss of generality $p_1 \geq p_2 \geq \dots \geq p_m$. Huffman gave an inductive definition of codes that we can prove are optimal. If $m = 2$, we take codewords 0, 1. If $m > 2$, first take the Huffman code for messages $\mu_1, \dots, \mu_{m-2}, \nu$ with probabilities $p_1, \dots, p_{m-2}, p_{m-1} + p_m$. Then append 0 (respectively 1) to the codeword for ν to give a codeword for μ_{m-1} (respectively μ_m).

Notes.

- Huffman codes are prefix-free;
- Huffman codes are not unique: choice is needed if some of the p_i are equal.

Example 3.1. Revisit Example 2.5. We have

i	p_i	$c^{(1)}$	$p_i^{(2)}$	$c^{(2)}$	$p_i^{(3)}$	$c^{(3)}$	$p_i^{(4)}$	$c^{(4)}$
1	0.4	1	0.4	1	0.4	1	0.6	0
2	0.2	01	0.2	01	0.4	00	0.4	1
3	0.2	000	0.2	000	0.2	01		
4	0.1	0010	0.2	001				
5	0.1	0011						

Theorem 3.1. *Huffman codes are optimal (Huffman, 1952).*

Proof. We show by induction on m that Huffman codes of size $m = |\mathcal{A}|$ are optimal.

$m = 2$: codewords are 0, 1 - clearly optimal.

$m > 2$: let c_m be a Huffman code for X_m , which takes values μ_1, \dots, μ_m with probabilities $p_1 \geq p_2 \geq \dots \geq p_m$; each c_m is constructed from Huffman code c_{m-1} for X_{m-1} which takes values $\mu_1, \dots, \mu_{m-2}, \nu$ with probabilities $p_1, \dots, p_{m-2}, p_{m-1} + p_m$. Then the expected word length is

$$\mathbb{E}S_m = \mathbb{E}S_{m-1} + p_{m-1} + p_m. \quad (*)$$

Let c'_m be an optimal code for X_m . Wlog c'_m is still prefix-free. Wlog the last two codewords of c'_m have maximal length and differ only in the final position (see next lemma). Say

$$c'_m(\mu_{m-1}) = y0, \quad c'_m(\mu_m) = y1 \text{ for some } y \in \{0, 1\}^*.$$

Let c'_{m-1} be some prefix-free code for X_{m-1} , given by

$$c'_{m-1}(\mu_i) = \begin{cases} c'_m(\mu_i) & 1 \leq i \leq m-2 \\ c'_{m-1}(\nu) = y & \end{cases}.$$

Then the expected word length satisfies

$$\mathbb{E}S'_m = \mathbb{E}S'_{m-1} + p_{m-1} + p_m. \quad (**)$$

By the inductive hypothesis, c_{m-1} is optimal, so $\mathbb{E}S_{m-1} \leq \mathbb{E}S'_{m-1}$. By (*) and (**) this implies $\mathbb{E}S_m \leq \mathbb{E}S'_m$. \square

Lemma 3.2. *Suppose letters μ_1, \dots, μ_m in \mathcal{A} are sent with probabilities p_1, p_2, \dots, p_m . Let c be an optimal (prefix-free) code with word lengths l_1, \dots, l_m . Then*

- (i) *If $p_i > p + j$, then $l_i \leq l_j$;*
- (ii) *Amongst all codewords of maximal length there exist two that differ only in the final digit.*

Proof. (i) is obvious. For (ii), could otherwise just delete the final digit of the codeword of maximal length (since prefix-free). \square

Remark. Note not all optimal codes are Huffman (look at the case $m = 4$).

Our main result says that if we have a prefix-free optimal code with word lengths l_1, \dots, l_m and associated probabilities p_1, \dots, p_m , then there is a Huffman code with these word lengths.

4 Joint Entropy

If X, Y are random variables with values in \mathcal{A} and \mathcal{B} respectively, then (X, Y) is a random variable with values in $\mathcal{A} \times \mathcal{B}$, and the *entropy* $H(X, Y)$ is called the joint entropy, given by

$$H(X, Y) = - \sum_{x \in \mathcal{A}} \sum_{y \in \mathcal{B}} \mathbb{P}(X = x, Y = y) \log \mathbb{P}(X = x, Y = y).$$

This generalises to any finite number of random variables.

Lemma 4.1. *Let X, Y be random variables taking values in \mathcal{A} and \mathcal{B} respectively. Then*

$$H(X, Y) \leq H(X) + H(Y),$$

with equality if and only if X and Y are independent.

Proof. Write $\mathcal{A} = \{x_1, \dots, x_m\}$, $\mathcal{B} = \{y_1, \dots, y_n\}$. Let

$$p_{ij} = \mathbb{P}(X = x_i, Y = y_j), \quad p_i = \mathbb{P}(X = x_i), \quad q_j = \mathbb{P}(Y = y_j).$$

Apply Gibbs' inequality to the probability distributions $\{p_{ij}\}$ and $\{p_i q_j\}$ to obtain

$$\begin{aligned} - \sum_{i,j} p_{ij} \log p_{ij} &\leq - \sum_{i,j} p_{ij} \log(p_i q_j) \\ &= - \sum_i \left(\sum_j p_{ij} \right) \log p_i - \sum_j \left(\sum_i p_{ij} \right) \log q_j \\ &= - \sum_i p_i \log p_i - \sum_j q_j \log q_j \\ &= H(X) + H(Y). \end{aligned}$$

With equality if and only if $p_{ij} = p_i q_j$ for all i, j .

□

Error-correcting codes

5 Noisy channels and Hamming's code

Definition. A *binary* $[m, n]$ -code is a subset C of $\{0, 1\}^n$ of size $m = |C|$. n is the length of the code and the elements of C are called codewords.

We use an $[n, m]$ -code to send one of m messages through a BSC (binary symmetric channel) making n uses of the channel. Clearly $1 \leq m \leq 2^n$, so $0 \leq \frac{1}{n} \log m \leq 1$.

Definition. For any $x, y \in \{0, 1\}^n$ the *Hamming distance* is

$$d(x, y) = |\{i : 1 \leq i \leq n, x_i \neq y_i\}|.$$

Definition.

- (i) The *ideal observer* decoding rule decodes $x \in \{0, 1\}^n$ as $c \in C$ maximising $\mathbb{P}(c \text{ sent} | x \text{ received})$.
- (ii) The *maximum likelihood* decoding rule decodes $x \in \{0, 1\}^n$ as $c \in C$ maximising $\mathbb{P}(x \text{ received} | c \text{ sent})$.
- (iii) The *minimum distance* decoding rule decodes $x \in \{0, 1\}^n$ as $c \in C$ minimising $d(x, c)$.

Lemma 5.1.

- (a) If all the messages are equally likely, then (i) and (ii) above are equivalent.
- (b) If $p < 1/2$ (error probability) then (ii) and (iii) are equivalent.

Remark. If $p = 1/2$ the code is called *useless*. If $p = 0$ the code is called *lossless*.

Proof.

- (a) We have

$$\mathbb{P}(c \text{ sent} | x \text{ received}) = \frac{\mathbb{P}(c \text{ sent}, x \text{ received})}{\mathbb{P}(x \text{ received})} = \frac{\mathbb{P}(c \text{ sent})}{\mathbb{P}(x \text{ received})} \mathbb{P}(x \text{ received} | c \text{ sent}).$$

So by hypothesis, $\mathbb{P}(c \text{ sent})$ is independent of $c \in C$. So for fixed x , maximising $\mathbb{P}(c \text{ sent} | x \text{ received})$ is the same as maximising $\mathbb{P}(x \text{ received} | c \text{ sent})$.

- (b) Let $r = d(x, c)$. Then $\mathbb{P}(x \text{ received} | c \text{ sent}) = p^r (1-p)^{n-r} = (1-p)^n \left(\frac{p}{1-p}\right)^r$. Since $p < 1/2$, $\frac{p}{1-p} < 1$. So maximising $\mathbb{P}(x \text{ received} | c \text{ sent})$ is the same as minimising r .

□

We choose to use minimum distance decoding from now on.

Example 5.1. Suppose 000, 111 are sent with probabilities $\alpha = 9/10$, $\beta = 1/10$ respectively through a BSC with error probability $p = 1/4$. Suppose 110 is received. Then

$$\mathbb{P}(000 \text{ sent} | 110 \text{ received}) = \frac{\alpha p^2(1-p)}{\alpha p^2(1-p) + (1-\alpha)p(1-p)^2} = \frac{3}{4},$$

$$\text{similarly } \mathbb{P}(111 \text{ sent} | 110 \text{ received}) = \frac{1}{4}.$$

So the ideal observer decodes as 000. But the maximum likelihood/minimum distance rules decode as 111.

Remarks.

- Minimum distance decoding may be expensive in terms of time and storage if $|C|$ is large.
- Need to specify a convention in case there is no unique maximiser (e.g make a random choice, or request the message is sent again).

We aim to detect, or even correct errors.

Definition. A code C is

- *d-error* detecting if changing up to d digits in each codeword can never produce another codeword. In other words, each codeword is of Hamming distance greater than d from every other codeword.
- *e-error* correcting if knowing that $x \in \{0,1\}^n$ differs from a codeword in at most e places we can deduce the codeword.

Examples.

- A *repetition code* of length n has codewords $\underbrace{00 \dots 0}_{n \text{ times}}, \underbrace{11 \dots 1}_{n \text{ times}}$. This is a $[n, 2]$ -code. It is $(n-1)$ -error detecting and $\lfloor \frac{n-1}{2} \rfloor$ -error correcting. But the information rate is only $1/n$.
- A *simple parity check code* or *paper tape code*: identify $\{0,1\}$ with \mathbb{F}_2 and let $C = \{(x_1, \dots, x_n) \in \{0,1\}^n : \sum_{i=1}^n x_i = 0\}$. This is a $[n, 2^{n-1}]$ -code, 1-error detecting but cannot correct errors. The information rate is $\frac{n-1}{n}$.
- Hamming's original code (1950): a 1-error correcting binary $[7, 16]$ -code. Take $C \subseteq \mathbb{F}_2^7$ where

$$C = \{c \in \mathbb{F}_2^7 : c_1 + c_3 + c_5 + c_7 = 0, c_2 + c_3 + c_6 + c_7 = 0, c_4 + c_5 + c_6 + c_7 = 0\}.$$

The bits c_3, c_5, c_6, c_7 are arbitrary and c_1, c_2, c_4 are forced (called the check digits) so $|C| = 2^4$. To decode: suppose we receive $x \in \mathbb{F}_2^7$. We form the *syndrome*: $z = z_x = (z_1, z_2, z_4) \in \mathbb{F}_2^3$ where

$$z_1 = x_1 + x_3 + x_5 + x_7$$

$$z_2 = x_2 + x_3 + x_6 + x_7$$

$$z_4 = x_4 + x_5 + x_6 + x_7.$$

If $x \in C$, then $z_x = (0, 0, 0)$. If $d(x, c) = 1$ for some $c \in C$, then place where x and c differ is given by $z_1 + 2z_2 + 4z_4$ (not mod 2). Check: if $x = c + e_i$ where e_i has all 0's except a 1 in the i th position, then $z_x = z_{e_i}$, so check for each $1 \leq i \leq 7$.

Lemma 5.2. *The Hamming distance is a metric on \mathbb{F}_2^n .*

Proof. Trivial. □

Definition. The *minimum distance of a code* is the minimum of $d(c_1, c_2)$ for all codewords c_1, c_2 with $c_1 \neq c_2$.

Lemma 5.3. *Let C be a code with minimum distance $d > 0$. Then*

- (i) *C is $(d - 1)$ -error detecting, but cannot detect all sets of d errors.*
- (ii) *C is $\lfloor \frac{d-1}{2} \rfloor$ -error correcting, but cannot correct all sets of $\lfloor \frac{d-1}{2} \rfloor + 1$ errors.*

Proof.

- (i) If $x \in \mathbb{F}_2^n$ and $c \in C$ are such that $0 < d(x, c) \leq d - 1$, then we know that $x \notin C$ so this is $(d - 1)$ -error detecting. However there must exist $c_1, c_2 \in C$ such that $d(c_1, c_2) = d$, so we cannot say if there's an error if c_1 is 'corrupted' to c_2 in d errors.
- (ii) Take $e = \lfloor \frac{d-1}{2} \rfloor$. If $x \in \mathbb{F}_2^n$ and $c_1 \in C$ are such that $d(x, c_1) \leq e$ then for any $c_1 \neq c_2 \in C$ we have $d(x, c_2) \geq d(c_1, c_2) - d(c_1, x) \geq d - e > e$. So C is e -error correcting. Now take $c_1, c_2 \in C$ with $d(c_1, c_2) = d$. Then take $x \in \mathbb{F}_2^n$ such that x differs from c_1 is precisely $e + 1$ places where c_1 and c_2 differ. Then $d(c_1, x) = e + 1$ and $d(x, c_2) = d - (e + 1) \leq e + 1$. So C cannot be $(e + 1)$ -error correcting. □

Definition. A $[n, m]$ -code with minimum distance d is called a $[n, m, d]$ -code.

Notes.

- $m \leq 2^n$ with equality if and only if $C = \mathbb{F}_2^n$ (trivial code)
- $d \leq n$, with equality in case of the repetition code.

Example 5.2.

- (i) Repetition code of length n is a $[n, 2, n]$ -code, $(n - 1)$ -error detecting and $\lfloor \frac{n-1}{2} \rfloor$ -error correcting.
- (ii) Simple parity check code is a $[n, 2^{n-1}, 2]$ -code, 1-error detecting and 0-error correcting.
- (iii) Hamming's original code is 1-error correcting, implying $d \geq 3$. Also 0000000, 1110000 are distance 3 apart, so $d = 3$. So this is a $[7, 16, 3]$ -code and is 2-error detecting.

6 Covering estimates

Take $x \in \mathbb{F}_2^n$, $r \geq 0$. Then $\overline{B}(x, r) = \{y \in \mathbb{F}_2^n : d(x, y) \leq r\}$ is the *closed Hamming ball*. Denote $V(n, r) = |\overline{B}(x, r)| = \sum_{i=0}^r \binom{n}{i}$, the *volume*.

Lemma 6.1 (Hamming's bound). *An e -error correcting code C of length n has*

$$|C| \leq \frac{2^n}{V(n, e)}.$$

Proof. C is e -error correcting so $\{B(c, e)\}_{c \in C}$ are pairwise disjoint balls, so $\sum_{c \in C} |B(c, e)| = |C|V(n, e) \leq |\mathbb{F}_2^n| = 2^n$. \square

Lemma 6.2. *A code C of length n that can correct e errors is perfect if $|C| = \frac{2^n}{V(n, e)}$. Equivalently, for all $x \in \mathbb{F}_2^n$ there exists a unique $c \in C$ such that $d(x, c) \leq e$. In this case, any $e + 1$ errors will make you decode incorrectly.*

Example 6.1.

(a) Hamming $[7, 16, 3]$ -code is 1-error correcting and

$$\frac{2^n}{V(n, e)} = \frac{2^7}{V(7, 1)} = \frac{2^7}{1 + 7} = 2^4 = |C|.$$

(b) Binary repetition code of length n (for n odd) is perfect.

Remark. If $\frac{2^n}{V(n, e)} \notin \mathbb{Z}$ then there does not exist a perfect e -error correcting code of length n . Converse is also false ($n = 90, e = 2$ on Example Sheet 2).

Definition. Define $A(n, d) = \max\{m : \exists [n, m, d]\text{-code}\}$.

The $A(n, d)$ are unknown in general. But we have some special cases:

Examples.

- $A(n, 1) = 2^n$ (trivial code)
- $A(n, n) = 2$ (repetition code)
- $A(n, 2) = 2^{n-1}$ (simple parity check code)

Lemma 6.3. $A(n, d + 1) \leq A(n, d)$.

Proof. Let $m = A(n, d + 1)$ and take a $[n, m, d + 1]$ -code C . Let $c_1, c_2 \in C$ have $d(c_1, c_2) = d + 1$. Let c'_1 differ from c_1 in a single place where c_1 and c_2 differ. Hence $d(c'_1, c_2) = d$. If $c \in C \setminus \{c_1\}$, then $d(c, c_1) \leq d(c, c'_1) + d(c'_1, c_1)$ so $d + 1 \leq d(c, c'_1) + 1$. Hence $d(c, c'_1) \geq d$. So replacing c_1 with c'_1 , we get an $[n, m, d]$ -code. \square

Corollary 6.4. $A(n, d) = \max\{m : \exists [n, m, d']\text{-code for some } d' \geq d\}$.

Theorem 6.5.

$$\frac{2^n}{V(n, d - 1)} \underbrace{\leq}_{\text{GSV bound}} A(n, d) \underbrace{\leq}_{\text{Hamming bound}} \frac{2^n}{V(n, \lfloor \frac{d-1}{2} \rfloor)}.$$

Proof. We have already proved the Hamming bound. So let $m = A(n, d)$. Let C be a $[n, m, d]$ -code. Then there does not exist $d(x, c) \geq d$ for all $c \in C$ (otherwise could replace C with $C \cup \{x\}$, contradicting maximality of m). Hence

$$\mathbb{F}_2^n \subseteq \bigcup_{c \in C} \overline{B}(c, d-1) \implies 2^n \leq \sum_{c \in C} |\overline{B}(c, d-1)| = mV(n, d-1).$$

□

Example 6.2. $n = 10, d = 3$, have $V(n, 1) = 11, V(n, 2) = 56$. The above theorem gives $19 \leq \frac{2^{10}}{56} \leq A(10, 3) \leq \frac{2^{10}}{11} \leq 93$. It was known that $72 \leq A(10, 3) \leq 93$, but the exact value of $A(10, 3)$ was only found in 1999.

Asymptotics of $V(n, r)$

We study $\frac{\log A(n, \lfloor n\delta \rfloor)}{n}$ as $n \rightarrow \infty$ to see how large the information rate can be for a given error rate.

Proposition 6.6. Let $\delta \in (0, 1/2)$. Then

$$(i) \log V(n, \lfloor n\delta \rfloor) \leq nH(\delta);$$

$$(ii) \frac{1}{n} \log A(n, \lfloor n\delta \rfloor) \geq 1 - H(\delta).$$

Where $H(\delta) = -\delta \log \delta - (1 - \delta) \log(1 - \delta)$.

Proof. First we show (i) \Rightarrow (ii): by the GSV bound,

$$A(n, \lfloor n\delta \rfloor) \geq \frac{2^n}{V(n, \lfloor n\delta \rfloor - 1)} \geq \frac{2^n}{V(n, \lfloor n\delta \rfloor)}$$

and so

$$\frac{\log A(n, \lfloor n\delta \rfloor)}{n} \geq 1 - \frac{\log V(n, \lfloor n\delta \rfloor)}{n} \geq 1 - H(\delta).$$

Now we prove (i): $H(\delta)$ is increasing for $\delta < 1/2$, so wlog we may assume $n\delta \in \mathbb{Z}$. Now

$$\begin{aligned} 1 &= (\delta + (1 - \delta))^n = \sum_{i=0}^n \binom{n}{i} \delta^i (1 - \delta)^{n-i} \geq \sum_{i=0}^{n\delta} \binom{n}{i} \delta^i (1 - \delta)^{n-i} \\ &= (1 - \delta)^n \sum_{i=0}^{n\delta} \binom{n}{i} \left(\frac{\delta}{1 - \delta} \right)^i \\ &\geq (1 - \delta)^n \sum_{i=0}^{n\delta} \binom{n}{i} \left(\frac{\delta}{1 - \delta} \right)^{n\delta} \\ &= \delta^{n\delta} (1 - \delta)^{n(1-\delta)} V(n, n\delta). \end{aligned}$$

Now taking logs:

$$0 \geq n(\delta \log \delta + (1 - \delta) \log(1 - \delta)) + \log V(n, n\delta).$$

□

The constant $H(\delta)$ in the above bound best possible:

Lemma 6.7. *We have*

$$\lim_{n \rightarrow \infty} \frac{\log V(n, \lfloor n\delta \rfloor)}{n} = H(\delta).$$

Proof. Exercise. □

7 Constructing new codes from old

We're given C , a $[n, m, d]$ -code. Can check the details in the following:

Examples.

1. The *parity check extension* C^+ is

$$\left\{ \left(c_1, \dots, c_n, \sum_{i=1}^n c_i \right) : (c_1, \dots, c_n) \in C \right\}$$

Is a $[n+1, m, d']$ -code with $d \leq d' \leq d+1$, depending on whether d is odd or even.

2. Fix $1 \leq i \leq n$. Deleting the i th digit from each codeword gives the *punctured code* C^-

$$\{(c_1, \dots, c_{i-1}, c_{i+1}, \dots, c_n) : (c_1, \dots, c_n) \in C\}.$$

If $d \geq 2$, then it is a $[n-1, m, d']$ -code with $d-1 \leq d' \leq d$.

3. Fix $1 \leq i \leq n$, and $\alpha \in \mathbb{F}_2$. The *shortened code* C' is

$$\{(c_1, \dots, c_{i-1}, c_{i+1}, \dots, c_n) : (c_1, \dots, c_{i-1}, \alpha, c_{i+1}, \dots, c_n) \in C\}.$$

It has parameters $[n, m', d']$ with $d' \geq d$ and $m' \geq \frac{m}{2}$ for a suitable choice of α .

Shannon's Theorems

8 AEP and Shannon's first coding theorem

Definition. A *source* is a sequence of random variables X_1, X_2, \dots taking values in some alphabet \mathcal{A} . A source is *Bernoulli (memoryless)* if X_1, X_2, \dots are iid: write (X, X_n) . A source X_1, X_2, \dots is *reliably encodable at rate r* if there exists a sequence of subsets $(A_n)_{n \geq 1}$ with $A_n \subseteq \mathcal{A}^n$ such that:

1. $\lim_{n \rightarrow \infty} \frac{\log |A_n|}{n} = r$;
2. $\lim_{n \rightarrow \infty} \mathbb{P}((X_1, \dots, X_n) \in A_n) = 1$.

The *information rate* H of a source is the infimum of all reliable encoding rates.
Exercise: $0 \leq H \leq \log |\mathcal{A}|$ with both bounds attainable.

Shannon's first coding theorem computes the information rate of certain sources, including Bernoulli sources.

Reminders from IA Probability:

We have a probability space $(\Omega, \mathcal{F}, \mathbb{P})$. A discrete random variable X is a function $X : \Omega \rightarrow \mathcal{A}$. The probability mass function $p_x : \mathcal{A} \rightarrow [0, 1]$ is defined by $x \mapsto \mathbb{P}(X = x)$. Can consider $p \circ X = p(X) : \Omega \rightarrow [0, 1]$, a random variable taking values in $[0, 1]$.

Given a source X_1, X_2, \dots of random variables with values in \mathcal{A} , the probability mass function of $X^{(n)} = (X_1, \dots, X_n)$ is $p_{X^{(n)}}$ given by $(x_1, \dots, x_n) \mapsto \mathbb{P}((X_1, \dots, X_n) = (x_1, \dots, x_n))$. Since $p_{X^{(n)}} : \mathcal{A}^n \rightarrow [0, 1]$ and $X^{(n)} : \Omega \rightarrow \mathcal{A}^n$, you can form $p(X^{(n)}) = p_{X^{(n)}} \circ X^{(n)} : \Omega \rightarrow [0, 1]$.

Example 8.1. Let $\mathcal{A} = \{A, B, C\}$. Suppose

$$X^{(2)} = \begin{cases} AB & \text{with probability 0.3} \\ AC & \text{with probability 0.1} \\ BC & \text{with probability 0.1} \\ BA & \text{with probability 0.2} \\ CA & \text{with probability 0.25} \\ CB & \text{with probability 0.05} \end{cases}.$$

Then

$$p(X^{(2)}) = \begin{cases} 0.3 & \text{with probability 0.3} \\ 0.1 & \text{with probability 0.2} \\ 0.2 & \text{with probability 0.2} \\ 0.25 & \text{with probability 0.25} \\ 0.05 & \text{with probability 0.05} \end{cases}.$$

So some points are “lumped together”.

Given a source X_1, X_2, \dots *converges in probability* to a random variable L (possibly constant) if for all $\varepsilon > 0$, $\mathbb{P}(|X_n - L| > \varepsilon) \rightarrow 0$ as $n \rightarrow \infty$. We write $X_n \xrightarrow{\mathbb{P}} L$.

The *Weak Law of Large Numbers* (WLLN) says that if $(X; X_n)$ are iid real-valued random variables with finite expectation $\mathbb{E}X$, we have

$$\frac{1}{n} \sum_{i=1}^n X_i \xrightarrow{\mathbb{P}} \mathbb{E}X.$$

Example 8.2. If X_1, X_2, \dots are iid Bernoulli, then $p(X_1), p(X_2), \dots$ are iid random variables and $p(X_1, \dots, X_n) = p(X_1) \dots p(X_n)$. Note

$$-\frac{1}{n} \log p(X_1, \dots, X_n) = -\frac{1}{n} \sum_{i=1}^n \log p(X_i) \xrightarrow{\mathbb{P}} -\mathbb{E}(-\log p(X_1)) = H(X_1) \text{ as } n \rightarrow \infty.$$

Lemma 8.1. *The information rate of a Bernoulli source X_1, X_2, \dots is at most the expected word length of an optimal code $c : \mathcal{A} \rightarrow \{0, 1\}^*$ for X_1 .*

Proof. Let l_1, l_2, \dots be the lengths of codewords when we encode X_1, X_2, \dots using c . Let $\varepsilon > 0$. Set $A_n = \{x \in \mathcal{A}^n : c^*(x) \text{ has length at less than } n(\mathbb{E}l_1 + \varepsilon)\}$. Then

$$\begin{aligned} \mathbb{P}((X_1, \dots, X_n) \in A_n) &= \mathbb{P}\left(\sum_{i=1}^n l_i < n(\mathbb{E}l_1 + \varepsilon)\right) \\ &= \mathbb{P}\left(\left|\frac{1}{n} \sum_{i=1}^n l_i - \mathbb{E}l_1\right| < \varepsilon\right) \\ &\rightarrow 1 \text{ as } n \rightarrow \infty. \end{aligned}$$

Now, c is decipherable so c^* is injective. Hence $|A_n| \leq 2^{n(\mathbb{E}l_1 + \varepsilon)}$. Making A_n larger if necessary, $|A_n| = \lfloor e^{n(\mathbb{E}l_1 + \varepsilon)} \rfloor$ so

$$\frac{\log(A_n)}{n} \rightarrow \mathbb{E}l_1 + \varepsilon.$$

Hence X_1, X_2, \dots is reliably encodable at rate $r = \mathbb{E}l_1 + \varepsilon$ for all $\varepsilon > 0$. Hence the information rate is at most $\mathbb{E}l_1$. \square

Corollary 8.2. *A Bernouilli source has information rate less than $H(X_1) + 1$.*

Proof. Combine the above with the Noiseless Coding Theorem. \square

We encode X_1, X_2, \dots in blocks

$$\underbrace{X_1, \dots, X_N}_{Y_1}, \underbrace{X_{n+1}, \dots, X_{2N}, \dots}_{Y_2}, \dots$$

so Y_1, Y_2, \dots take values in \mathcal{A}^N . Exercise: show that if X_1, X_2, \dots has information rate H then Y_1, Y_2, \dots has information rate NH .

Proposition 8.3. *The information rate H of a Bernouilli source X_1, X_2, \dots is at most $H(X_1)$.*

Proof. Apply the previous corollary to Y_1, Y_2, \dots and obtain

$$NH < H(Y_1) + 1 = H(X_1, \dots, X_N) + 1 = \sum_{i=1}^N H(X_i) + 1 = NH(X_1, \dots, X_n) + 1.$$

Hence $H < H(X_1) + \frac{1}{N}$. Since N is arbitrary, $H \leq H(X_1)$. \square

Definition. A source X_1, X_2, \dots satisfies the *Asymptotic Equipartition Property* (AEP) for some constant $H \geq 0$ if

$$-\frac{1}{n} \log p(X_1, X_2, \dots) \xrightarrow{\mathbb{P}} H \text{ as } n \rightarrow \infty.$$

Example 8.3. Tossing a biased coin, $\mathbb{P}(H) = p$. Let $(X; X_n)$ be the results of independent coin tosses. After a large number N of tosses, expect on average pN heads and $(1-p)N$ tails. The probability of any particular sequence of pN heads and $(1-p)N$ tails is $p^{pN}(1-p)^{(1-p)N} = 2^{N(p \log p + (1-p) \log (1-p))} = 2^{-NH(X)}$. Not every sequence of tosses will be like this, but there is only a small probability of “atypical” sequences. With high probability we get a “typical” sequence and its probability will be close to $2^{-NH(X)}$.

Lemma 8.4. *The AEP for a source X_1, X_2, \dots is equivalent to the following property*

$\forall \varepsilon > 0 \exists n_0(\varepsilon)$ such that $\forall n \geq n_0(\varepsilon) \exists$ a “typical set” $T_n \subseteq \mathcal{A}^n$ such that

(i) $\mathbb{P}((X_1, \dots, X_n) \in T_n) > 1 - \varepsilon$;

(ii) $2^{-n(H+\varepsilon)} \leq p(x_1, \dots, x_n) \leq 2^{-n(H-\varepsilon)}$ for all $(x_1, \dots, x_n) \in T_n$.

Proof. Obvious and non-examinable. □

Theorem 8.5 (Shannon's First Coding Theorem (FCT)). *If a source X_1, X_2, \dots satisfies the AEP with constant H , then the source has information rate H .*

Proof. Let $\varepsilon > 0$ and let $T_n \subseteq \mathcal{A}^n$ be typical sets. Then for some $n_0(\varepsilon)$ and all $n \geq n_0(\varepsilon)$

$$p(x_1, \dots, x_n) \geq 2^{-n(H+\varepsilon)} \text{ for all } (x_1, \dots, x_n) \in T_n$$

$$\implies \mathbb{P}(T_n) \geq 2^{-n(H+\varepsilon)} |T_n| \implies 1 \geq 2^{-n(H+\varepsilon)} |T_n| \implies \frac{\log |T_n|}{n} \leq H + \varepsilon.$$

Taking $A_n = T_n$, shows the source is reliably encodable at rate $H + \varepsilon$. Conversely, if $H = 0$, we're done. Otherwise pick $0 < \varepsilon < H/2$ and suppose for contradiction that the source is reliably encodable at rate $H - 2\varepsilon$, say with sets $A_n \subseteq \mathcal{A}^n$. Let $T_n \subseteq \mathcal{A}^n$ be typical sets. Then for all $(x_1, \dots, x_n) \in T_n$

$$p(x_1, \dots, x_n) \leq 2^{-n(H-\varepsilon)}$$

$$\implies \mathbb{P}(A_n \cap T_n) \leq 2^{-n(H-\varepsilon)} |A_n|$$

$$\implies \frac{\log \mathbb{P}(A_n \cap T_n)}{n} \leq H - \varepsilon + \frac{\log |A_n|}{n} \rightarrow -(H - \varepsilon) + H - 2\varepsilon = -\varepsilon.$$

Hence $\log \mathbb{P}(A_n \cap T_n) \rightarrow -\infty$ and $\mathbb{P}(A_n \cap T_n) \rightarrow 0$. But $\mathbb{P}(T_n) \leq \mathbb{P}(A_n \cap T_n) + \mathbb{P}(\mathcal{A}^n \setminus A_n) \rightarrow 0$, a contradiction to $\mathbb{P}(T_n) \rightarrow 1$. Thus the information rate is exactly H . \square

Corollary 8.6. *A Bernoulli source X_1, X_2, \dots has information rate $H(X_1)$.*

Proof. We've already seen that $-\frac{1}{n} \log p(X_1, \dots, X_n) \xrightarrow{\mathbb{P}} H(X_1)$, so done by Shannon's First Coding Theorem. \square

Remarks.

- The AEP is useful for noiseless coding. We can
 - encode the typical sequences using a block code;
 - encode the atypical sequences arbitrarily.
- Many sources, which are not necessarily Bernoulli satisfy the AEP. Under suitable hypotheses the sequence $\frac{1}{n} H(X_1, \dots, X_n)$ is decreasing and the AEP is satisfied.

9 Capacity & Shannon's second coding theorem

Recall:

Definition. We model n uses of a channel by the n th extension, with input alphabet \mathcal{A}^n and output alphabet \mathcal{B}^n . A code C of length n is a function $\mathcal{M} \rightarrow \mathcal{A}^n$ where \mathcal{M} is the set of possible messages. Implicitly we also have a decoding rule $\mathcal{B}^n \rightarrow \mathcal{M}$. The size of C is $m = |\mathcal{M}|$. The information rate is $\rho(C) = \frac{1}{n} \log_2 m$. The error rate is $\hat{e}(C) = \max_{x \in \mathcal{M}} \mathbb{P}(\text{error} | x \text{ sent})$.

Definition. A channel can *transmit reliably at rate R* if there exists $(C_n)_{n=1}^\infty$ with each C_n a code of length n such that

$$\lim_{n \rightarrow \infty} \rho(C_n) = R \text{ \& } \lim_{n \rightarrow \infty} \hat{e}(C_n) = 0.$$

The *capacity* is the supremum of all reliable transmission rates.

Suppose we are given a source where

- it has information rate r bits per second;
- it emits symbols at s symbols per second.

Suppose we are also given a channel where

- it has capacity R bits per transmission;
- it transmits symbols at S transmissions per second.

Usually, information theorists take $S = s = 1$. If $rs \leq RS$ then you can encode and transmit reliably, and if $rs > RS$ you cannot.

We'll compute the capacity of a BSC with error probability p .

Proposition 9.1. *A binary symmetric channel with error probability $p < 1/4$ has non-zero capacity.*

Proof. We use the GSV bound. Pick $\delta \in (2p, 1/2)$. We claim reliable transmission at rate $R = 1 - H(\delta) > 0$. Let C_n be a code of length n , and suppose it has minimum distance $\lfloor n\delta \rfloor$ of maximal size. Then (by Proposition 6.6(ii))

$$|C_n| = A(n, \lfloor n\delta \rfloor) \geq 2^{n(1-H(\delta))}.$$

Replacing C_n by a subcode, we can assume $|C_n| = \lfloor 2^{nR} \rfloor$ and still minimum distance $\geq \lfloor n\delta \rfloor$. Using minimum distance decoding

$$\begin{aligned} \hat{e}(C_n) &\leq \mathbb{P} \left(\text{in } n \text{ uses, BSC makes } \geq \left\lfloor \frac{\lfloor n\delta \rfloor - 1}{2} \right\rfloor \text{ errors} \right) \\ &\leq \mathbb{P} \left(\text{in } n \text{ uses, BSC makes } \geq \left\lfloor \frac{n\delta - 1}{2} \right\rfloor \text{ errors} \right). \end{aligned}$$

Pick $\varepsilon > 0$ with $p + \varepsilon < \frac{\delta}{2}$. For n sufficiently large, $\frac{n\delta - 1}{2} = n \left(\frac{\delta}{2} - \frac{1}{2n} \right) > n(p + \varepsilon)$. Hence $\hat{e}(C_n) \leq \mathbb{P}(\text{BSC makes } \geq n(p + \varepsilon) \text{ errors}) \rightarrow 0$, using the next lemma. \square

Lemma 9.2. Let $\varepsilon > 0$. A BSC with error probability p is used to transmit n digits. Then

$$\lim_{n \rightarrow \infty} \mathbb{P}(\text{BSC makes } \geq n(p + \varepsilon) \text{ errors}) = 0.$$

Proof. Consider random variables U_i taking value 1 if the i th digit is mistransmitted, and value 0 otherwise. Then $\mathbb{E}U_i = p$, so $\mathbb{P}(\text{BSC makes } \geq n(p + \varepsilon) \text{ errors}) \leq \mathbb{P}(|\frac{1}{n} \sum_{i=1}^n U_i - p| \geq \varepsilon) \rightarrow 0$. \square

Definition (Conditional Entropy). Let X, Y be random variables taking values in alphabets \mathcal{A}, \mathcal{B} respectively. Then we define the *conditional entropy*

$$H(X|Y = y) = - \sum_{x \in \mathcal{A}} \mathbb{P}(X = x|Y = y) \log \mathbb{P}(X = x|Y = y)$$

$$H(X|Y) = \sum_{y \in \mathcal{B}} \mathbb{P}(Y = y) H(X|Y = y).$$

Clearly $H(X|Y) \geq 0$.

Lemma 9.3. We have

$$H(X, Y) = H(X|Y) + H(Y).$$

Proof.

$$\begin{aligned} H(X|Y) &= - \sum_{y \in \mathcal{B}} \sum_{x \in \mathcal{A}} \mathbb{P}(X = x|Y = y) \mathbb{P}(Y = y) \log \mathbb{P}(X = x|Y = y) \\ &= - \sum_{y \in \mathcal{B}} \sum_{x \in \mathcal{A}} \mathbb{P}(X = x, Y = y) \log \left(\frac{\mathbb{P}(X = x, Y = y)}{\mathbb{P}(Y = y)} \right) \\ &= - \sum_{y \in \mathcal{B}} \sum_{x \in \mathcal{A}} \mathbb{P}(X = x, Y = y) \log \mathbb{P}(X = x, Y = y) \\ &\quad + \sum_{y \in \mathcal{B}} \sum_{x \in \mathcal{A}} \mathbb{P}(X = x, Y = y) \log \mathbb{P}(Y = y) \\ &= H(X, Y) - \sum_{y \in \mathcal{B}} \mathbb{P}(Y = y) \log \mathbb{P}(Y = y) \\ &= H(X, Y) - H(Y). \end{aligned}$$

\square

Example 9.1. We roll fair die. Let X be the value defined by the roll, let Y be equal to 0 if X is even and 1 if X is odd. Then $H(X, Y) = H(X) = \log 6$, $H(Y) = \log 2 = 1$, $H(X|Y) = \log 3 = H(X, Y) - H(Y)$. Similarly $H(Y|X) = 0 = H(X, Y) - H(X)$.

Corollary 9.4. $H(X|Y) \leq H(X)$.

Proof. Combine the previous with Lemma 4.1. \square

Now replace X, Y with random vectors $X^{(r)} = (X_1, \dots, X_r), Y^{(s)} = (Y_1, \dots, Y_s)$. Similarly can define $H(X_1, \dots, X_r | Y_1, \dots, Y_s) = H(X^{(r)} | Y^{(s)})$.

Remark. $H(X, Y | Z)$ is the entropy of (X, Y) given Z , not the entropy of X and $Y | Z$.

Lemma 9.5. *Let X, Y, Z be random variables. Then $H(X, Y) \leq H(X | Y, Z) + H(Z)$.*

Proof. We expand $H(X, Y, Z)$ in two different ways.

$$H(X, Y, Z) = H(Z | X, Y) + H(X | Y) + H(Y),$$

$$H(X, Y, Z) = H(X | Y, Z) + H(Z | Y) + H(Y).$$

Since $H(Z | X, Y) \geq 0$, we have

$$\begin{aligned} H(X | Y) &\leq H(X | Y, Z) + H(Z | Y) \\ &\leq H(X | Y, Z) + H(Z). \end{aligned}$$

□

Proposition 9.6 (Fano's inequality). *Let X, Y be random variables taking values in \mathcal{A} , $|\mathcal{A}| = m$. Let $p = \mathbb{P}(X \neq Y)$. Then*

$$H(X | Y) \leq H(p) + p \log(m - 1).$$

Proof. Define $Z = 0$ if $X = Y$ and $Z = 1$ if $X \neq Y$. Then $\mathbb{P}(Z = 0) = 1 - p$, $\mathbb{P}(Z = 1) = p$. So $H(Z) = H(p)$. By the previous lemma,

$$H(X | Y) \leq H(p) + H(X | Y, Z). \quad (*)$$

Since $Z = 0$ implies $X = Y$, $H(X | Y = y, Z = 0) = 0$. Also there are $m - 1$ remaining possibilities for X , so $H(X | Y = y, Z = 1) \leq \log(m - 1)$. Therefore

$$\begin{aligned} H(X | Y) &= \sum_{y, z} \mathbb{P}(Y = y, Z = z) H(X | Y = y, Z = z) \\ &\leq \sum_{y \in \mathcal{A}} \mathbb{P}(Y = y, Z = 1) \log(m - 1) \\ &= \mathbb{P}(Z = 1) \log(m - 1) \\ &= p \log(m - 1). \end{aligned}$$

So by $(*)$, $H(X | Y) \leq H(p) + p \log(m - 1)$. □

Remark. $H(p)$ represents the information needed to decide whether or not there is an error, and $p \log(m - 1)$ represents the information needed to resolve the error assuming the worst possible case.

Definition. Let X, Y be random variables. Then the *mutual information* is $I(X; Y) := H(X) - H(X | Y)$. By Corollary 9.4, $I(X; Y) \geq 0$ with equality if and only if X, Y are independent. Clearly $I(X; Y) = I(Y; X)$.

Suppose we're given a discrete memoryless channel (DMC) with input alphabet \mathcal{A} , $|\mathcal{A}| = m$ and output alphabet \mathcal{B} . Let X be a random variable taking values in \mathcal{A} and used as input to the channel. Let Y be a random variable output, depending on X and the channel matrix.

Definition. The (*information*) *capacity* is $\max_X I(X; Y)$.

Remarks.

- The maximum is over all probability distributions (p_1, \dots, p_m) for X on \mathcal{A} .
- The maximum is attained since $(p_1, \dots, p_m) \mapsto I((p_1, \dots, p_m); Y)$ is continuous on the compact set

$$\{(p_1, \dots, p_m) \in [0, 1]^m : p_1 + \dots + p_m = 1\}.$$

- The information capacity depends only on the channel matrix.

Theorem 9.7 (Shannon's Second Coding Theorem). *For a DMC, the operational capacity is equal to the information capacity.*

Remark. We'll show one inequality in general and the other for the BSC only.

Assuming Shannon's Second Coding Theorem, let us compute the capacity of certain channels.

Example 9.2. BSC, error probability p . Input X : $\mathbb{P}(X = 0) = \alpha$, $\mathbb{P}(X = 1) = 1 - \alpha$. Output Y : $\mathbb{P}(Y = 0) = \alpha(1 - p) + (1 - \alpha)p$, $\mathbb{P}(Y = 1) = (1 - \alpha)(1 - p) + \alpha p$. Then C is

$$\begin{aligned} \max_{\alpha} I(X; Y) &= \max_{\alpha} (H(Y) - H(Y|X)) \\ &= \max_{\alpha} (H(\alpha(1 - p) + (1 - \alpha)p) - H(p)) \\ &= 1 - H(p). \end{aligned}$$

Where the maximum is attained for $\alpha = 1/2$. Hence $C = 1 + p \log p + (1 - p) \log(1 - p)$.

Remark. We can choose to calculate either $H(Y) - H(Y|X)$ or $H(X) - H(X|Y)$ depending on which is easier.

Example 9.3. BEC, erasure probability p . Input X : $\mathbb{P}(X = 0) = \alpha$, $\mathbb{P}(X = 1) = 1 - \alpha$. Output Y : $\mathbb{P}(Y = 0) = \alpha(1 - p)$, $\mathbb{P}(Y = *) = p$, $\mathbb{P}(Y = 1) = (1 - \alpha)(1 - p)$. Can calculate $H(X|Y = 0) = 0$, $H(X|Y = 1) = 0$, $H(X|Y = *) = H(\alpha)$, which gives $H(X|Y) = pH(\alpha)$. So

$$\begin{aligned} C = \max_{\alpha} I(X; Y) &= \max_{\alpha} (H(X) - H(X|Y)) = \max_{\alpha} (H(\alpha) - pH(\alpha)) \\ &= (1 - p) \max_{\alpha} H(\alpha) \\ &= 1 - p. \end{aligned}$$

Where the max is attained for $\alpha = 1/2$.

Now model using a channel n times as the n th extension, i.e replace alphabets \mathcal{A}, \mathcal{B} with $\mathcal{A}^n, \mathcal{B}^n$,

$$\mathbb{P}(y_1, \dots, y_n \text{ received} | x_1, \dots, x_n \text{ sent}) = \prod_{i=1}^n \mathbb{P}(y_i | x_i).$$

Lemma 9.8. The n th extension of a DMC with information capacity C has information capacity nC .

Proof. Take random variable input X_1, \dots, X_n producing output Y_1, \dots, Y_n . Since the channel is memoryless,

$$H(Y_1, \dots, Y_n | X_1, \dots, X_n) = \sum_{i=1}^n H(Y_i | X_1, \dots, X_n) = \sum_{i=1}^n H(Y_i | X_i).$$

Hence

$$\begin{aligned}
 I(X_1, \dots, X_n; Y_1, \dots, Y_n) &= H(Y_1, \dots, Y_n) - H(Y_1, \dots, Y_n | X_1, \dots, X_n) \\
 &= H(Y_1, \dots, Y_n) - \sum_{i=1}^n H(Y_i | X_i) \\
 &\leq \sum_{i=1}^n (H(Y_i) - H(Y_i | X_i)) \\
 &= \sum_{i=1}^n I(X_i; Y_i) \leq nC.
 \end{aligned}$$

To finish, we find a distribution for X_1, \dots, X_n giving equality. Equality is attained by taking X_1, \dots, X_n independent, each of the same distribution such that $I(X_i; Y_i) = C$. Indeed, if X_1, \dots, X_n are independent, so are Y_1, \dots, Y_n . \square

Proposition 9.9. *For a DMC the (operational) capacity is at most the information capacity. Let C be the information capacity. Suppose reliable transmission is possible at some rate $R > C$, i.e there exist a sequence of codes $(C_n)_{n \geq 1}$ with C_n having length n and size $\lceil 2^{nR} \rceil$ for all n such that*

$$\lim_{n \rightarrow \infty} \rho(C_n) = R \text{ and } \lim_{n \rightarrow \infty} \hat{e}(C_n) = 0.$$

(Recall $\hat{e}(C_n) = \max_{c \in C_n} \mathbb{P}(\text{error} | c \text{ sent})$.)

Proof. Define the average error rate $e(C_n) = \frac{1}{|C_n|} \sum_{c \in C_n} \mathbb{P}(\text{error} | c \text{ sent})$. Note $e(C_n) \leq \hat{e}(C_n)$ and so $e(C_n) \rightarrow 0$ as $n \rightarrow \infty$. Take input random variable X equidistributed over C_n (i.e takes all values with equal probability). Let Y be the random variable output when X is transmitted and decoded. So $e(C_n) = \mathbb{P}(X \neq Y)$, define $p := e(C_n)$. Now we have

$$H(X) = \log |C_n| = \log \lceil 2^{nR} \rceil \geq nR - 1 \text{ for sufficiently large } n.$$

And

$$\begin{aligned}
 H(X|Y) &\leq H(p) + p \log(|C_n| - 1) && \text{(Fano's inequality)} \\
 &\leq 1 + pnR.
 \end{aligned}$$

Recall $I(X; Y) = H(X) - H(X|Y)$. By the previous lemma $nC \geq I(X; Y)$ so

$$nC \geq nR - 1 - (1 + pnR)$$

$$\implies pnR \geq n(R - C) - 2 \implies p \geq \frac{n(R - C) - 2}{nR} \rightarrow \frac{R - C}{R} \neq 0.$$

Since $R > C$, which contradicts $p \rightarrow 0$. Hence we conclude that we cannot transmit reliably at any rate exceeding C . \square

To complete the proof of Shannon's SCT for a BSC with error probability p , we need to show that the operational capacity is at least $1 - H(p)$ (i.e the information capacity).

Proposition 9.10. *Consider BSC with error probability p . Suppose $R < 1 - H(p)$. Then there is a sequence of codes $(C_n)_{n \geq 1}$ with C_n of length n and size $\lceil 2^{nR} \rceil$ for all n such that*

$$\lim_{n \rightarrow \infty} \rho(C_n) = R \text{ and } \lim_{n \rightarrow \infty} e(C_n) = 0.$$

Remark. Note that the above proposition deals with the average error rate e , not the error rate \hat{e} .

Proof. We'll use the method of 'random coding'. Wlog $p < 1/2$. Take $\varepsilon > 0$ such that

$$p + \varepsilon < 1/2 \text{ and } R < 1 - H(p + \varepsilon).$$

(This is possible by continuity of H). We use minimum distance decoding (in case of a tie, make arbitrary choice). Let $m = \lceil 2^{nR} \rceil$. Pick $C = \{c_1, \dots, c_m\}$ at random from $\mathcal{C} = \{[n, m]\text{-codes}\}$, a set of size $\binom{2^n}{m}$. Choose $1 \leq i \leq m$ at random (i.e each with probability $1/m$). We sent c_i through the channel and get output Y . Then

$$\begin{aligned} \mathbb{P}(Y \text{ not decoded as } c_i) &= \text{average value of } e(C) \text{ as } C \text{ runs through } \mathcal{C} \\ &= \frac{1}{|\mathcal{C}|} \sum_{C \in \mathcal{C}} e(C). \end{aligned}$$

We can pick C_n , an $[n, m]$ -code with $e(C_n)$ at most this average. So it is enough to show

$$\mathbb{P}(Y \text{ is not decoded as } c_i) \rightarrow 0 \text{ as } n \rightarrow \infty.$$

Let $r = \lfloor n(p + \varepsilon) \rfloor$. Then if $B(Y, r) \cap C = \{c_i\}$, Y is correctly decoded as c_i . So

$$\mathbb{P}(Y \text{ is not decoded as } c_i) \leq \mathbb{P}(c_i \notin B(Y, r)) + \mathbb{P}(B(Y, r) \cap C \supsetneq \{c_i\}).$$

Consider the two cases separately, according as $d(c_i, Y) > r$ (case (i)) or $d(c_i, Y) \leq r$ (case (ii)). For (i) we have

$$\begin{aligned} \mathbb{P}(d(c_i, Y) > r) &= \mathbb{P}(\text{BSC makes } > r \text{ errors}) \\ &= \mathbb{P}(\text{BSC makes } > n(p + \varepsilon) \text{ errors}) \\ &\rightarrow 0 \text{ as } n \rightarrow \infty \text{ by a previous theorem.} \end{aligned}$$

For (ii), if $j \neq i$, $\mathbb{P}(c_j \in B(Y, r) | c_i \in B(Y, r)) = \frac{V(n, r) - 1}{2^n - 1} \leq \frac{V(n, r)}{2^n}$ (since

$V(n, r) \leq 2^n$. Hence

$$\begin{aligned}
 \mathbb{P}(B(Y, r) \cap C \supsetneq \{c_i\}) &\leq \sum_{j \neq i} \mathbb{P}(c_j \in B(Y, r) \text{ and } c_i \in B(Y, r)) \\
 &\leq \sum_{j \neq i} \mathbb{P}(c_j \in B(Y, r) | c_i \in B(Y, r)) \\
 &\leq (m-1) \frac{V(n, r)}{2^n} \\
 &\leq \frac{mV(n, r)}{2^n} \\
 &\leq 2^{nR} 2^{nH(p+\varepsilon)2^{-n}} \\
 &= 2^{n(R-(1-H(p+\varepsilon)))} \xrightarrow{n \rightarrow \infty} 0
 \end{aligned}$$

since $R < 1 - H(p + \varepsilon)$. \square

Proposition 9.11. *We can replace e by \hat{e} in the previous result.*

Proof. Pick R' with $R < R' < 1 - H(p)$. The previous result constructs a sequence $(C'_n)_{n \geq 1}$ of codes, C'_n of length n and size $\lfloor 2^{nR'} \rfloor$ for all n , and $e(C'_n) \rightarrow 0$ as $n \rightarrow \infty$.

For each n , order the codewords of C'_n by $\mathbb{P}(\text{error} | c \text{ sent})$ and delete the worst half, to get a code C_n with $\hat{e}(C_n) \leq 2e(C'_n)$. So $\hat{e}(C_n) \rightarrow 0$ as $n \rightarrow \infty$. Since C_n has length n and size $\lfloor 2^{nR'} \rfloor / 2 = \text{floor } 2^{nR'-1}$. But $2^{nR'-1} = 2^{n(R'-1/n)} \geq 2^{nR}$ for all n sufficiently large. So can replace C'_n by a code of smaller size $\lfloor 2^{nR} \rfloor$ and still get $\hat{e}(C_n) \rightarrow 0$ and $\rho(C_n) \rightarrow R$. \square

Remarks.

1. A BSC with error probability p has operational capacity $1 - H(p)$. Indeed the above proposition says we can transmit reliably at any rate $R < 1 - H(p)$, whereas 9.9 says the capacity is at most $1 - H(p)$.
2. Shannon's SCT shows that good codes exist, but the proof does not say how to construct them.

10 The Kelly criterion

Let $1 > p > 0$, $u > 0$, $1 > w \geq 0$. Suppose a coin is tossed n times in succession, $\mathbb{P}(H) = p$. If I pay k ahead of a particular throw, then get back ku if the throw is heads, or get back 0 if throw is tails.

Say we have initial bankroll $X_0 = 1$. If we have bankroll of X_n after n th toss, bet wX_n , retaining $(1 - w)X_n$. Then bankroll X_{n+1} after $(n + 1)$ th throw is

$$X_{n+1} = \begin{cases} X_n(wu + (1 - w)) & \text{if } n\text{th toss is heads} \\ X_n(1 - w) & \text{if } n\text{th toss is tails} \end{cases}$$

Let $Y_{n+1} = X_{n+1}/X_n$. Then Y_1, Y_2, \dots , is a sequence of iid random variables. Hence $\log Y_1, \log Y_2, \dots$ is a sequence of iid random variables. Note $\log X_n = \sum_{i=1}^n \log Y_i$.

Lemma 10.1. Let $\mathbb{E} \log Y_1 = \mu$, $\text{Var}(\log Y_1) = \sigma^2$. Then if $a > 0$

$$(i) \quad \mathbb{P} \left(\left| \frac{1}{n} \sum_{i=1}^n \log Y_i - \mu \right| \geq a \right) \leq \frac{\sigma^2}{na^2}$$

$$(ii) \quad \mathbb{P} \left(\left| \frac{\log X_n}{n} - \mu \right| \geq a \right) \leq \frac{\sigma^2}{na^2}$$

$$(iii) \quad \text{Given } \varepsilon > 0 \text{ and } \delta > 0 \text{ there exists } N \in \mathbb{N} \text{ such that } \mathbb{P} \left(\left| \frac{\log X_n}{n} - \mu \right| \geq \delta \right) \leq \varepsilon \text{ for all } n \geq N.$$

Proof. Obvious by Chebyshev. \square

Lemma 10.2. Consider a single toss of a coin, $\mathbb{P}(H) = p < 1$. Suppose a bet of k on H has payout ratio $u > 0$. Suppose we have initial bankroll 1, and we bet w on H , retaining $1 - w$ ($0 \leq w < 1$). If $Y =$ fortune after toss then $\mathbb{E} \log Y = p \log(1 + (u - 1)w) + (1 - p) \log(1 - w)$. The value of $\mathbb{E} \log Y$ is maximised by taking $w = 0$ if $up \leq 1$, and otherwise taking $w = \frac{up-1}{u-1}$.

Proof. Not given (see Körner's book). \square

Remark. Put $q = 1 - p$. If $up > 1$, at the optimum w have

$$\mathbb{E} \log Y = p \log p + q \log q + \log u - q \log(u - 1)$$

and $-(p \log p + q \log q)$ is the entropy of a simple probabilistic system.

The suggestion that those making a long sequence of bets should maximise the expectation of the log is called Kelly's criterion.

Exercise: show that if you bet less than the optimum, your bankroll will tend to increase, but more slowly; and if you bet more than some critical proportion, your bankroll will tend to decrease.

11 Algebraic Coding Theory

11.1 Linear codes

Definition. A code $C \subseteq \mathbb{F}_2^n$ is *linear* if

- $0 \in C$;
- whenever $x, y \in C$, we have $x + y \in C$.

i.e C is a \mathbb{F}_2 -vector subspace of \mathbb{F}_2^n .

Definition. The *rank* of a linear code C is its dimension as a \mathbb{F}_2 -vector space. A linear code of length n , rank k is called a (n, k) -code. If the minimum distance of the code is d , it is a (n, k, d) -code.

Let v_1, \dots, v_k be a basis of C . Then $C = \{\sum_{i=1}^k \lambda_i v_i : \lambda_1, \dots, \lambda_k \in \mathbb{F}_2\}$. So $|C| = 2^k$. So an (n, k) -code is a $[n, 2^k]$ -code. The information rate is k/n .

Definition. The *weight* of $x \in \mathbb{F}_2^n$ is $w(x) = d(x, 0)$.

Lemma 11.1. *The minimum distance of a linear code is the minimum weight of a non-zero code word.*

Proof. Obvious. □

Definition. For $x, y \in \mathbb{F}_2^n$, let $x \cdot y = \sum_{i=1}^n x_i y_i \in \mathbb{F}_2$. This operation is symmetric and bilinear, but not positive definite.

Definition. Take some $P \subseteq \mathbb{F}_2^n$. The *parity check code* defined by P is

$$C = \{x \in \mathbb{F}_2^n : p \cdot x = 0, \forall p \in P\}.$$

Examples.

- $P = \{(1, \dots, 1)\}$ gives the simple parity check code;
- $P = \{(1, 0, 1, 0, 1, 0, 1), (0, 1, 1, 0, 0, 1, 1), (0, 0, 0, 1, 1, 1, 1)\}$ gives Hamming's $[7, 4, 3]$ -code;
- Show that C^+ and C^- are linear if C is linear. When is C' linear?

Lemma 11.2. *Every parity check code is linear.*

Proof. Obvious. □

Definition. Take $C \subseteq \mathbb{F}_2^n$ a linear code. The *dual code*

$$C^\perp = \{x \in \mathbb{F}_2^n : x \cdot y = 0, \forall y \in C\}.$$

This is a parity-check code, so is linear. Note $C \cap C^\perp$ is not necessarily $\{0\}$.

Lemma 11.3. $\dim(C) + \dim(C^\perp) = n$.

Proof. See later. □

Corollary 11.4. *Let C be linear. Then $(C^\perp)^\perp = C$. In particular, C is a parity check code.*

Proof. Clearly $C \subseteq (C^\perp)^\perp$. Equality follows from the previous lemma. □

Definition. Let C be a (n, k) -code.

- (i) A *generator matrix* G for C is a $k \times n$ matrix with rows a basis of C ;
- (ii) A *parity check matrix* H for C is a generator matrix for C^\perp . It is a $(n - k) \times n$ matrix.

The codewords of C can be viewed either as:

- Linear combinations of rows of G ;
- Linear dependence relations between the columns of H , i.e $C = \{x \in \mathbb{F}_2^n : Hx^T = 0\}$.